HARNESSING TUNNELS FOR DIRTY-SLATE NETWORK SOLUTIONS

A Dissertation

Presented to the Faculty of the Graduate School

of Cornell University

in Partial Fulfillment of the Requirements for the Degree of

Doctor of Philosophy

by

Hitesh Ballani

August 2009

HARNESSING TUNNELS FOR DIRTY-SLATE NETWORK SOLUTIONS

Hitesh Ballani, Ph.D.

Cornell University 2009

The tremendous success of the Internet has been both a boon and bane for networking research. On one hand, Internet growth has led to a plethora of problems and has prompted work towards next-generation network architectures. While very important, the success of the Internet has also meant that such *clean-slate* proposals are difficult to deploy. Thus, it is imperative that we find practically deployable *dirty-slate* solutions. In this thesis, we explore the possibility of tackling network problems in the existing framework through the use of tunnels. Tunneling involves encapsulating protocols in each other and we argue that this can serve as an enabler to the use of existing protocols in novel ways. We have found that, in many cases, such an approach can be used to address the root cause of a problem afflicting the network without necessitating protocol changes. Further, the increasing adoption of tunnels in mainstream networks augurs well for the deployability of such tunnels-based solutions.

In this thesis, we focus on two important network problems and present tunnel-driven, dirty-slate solutions to address them. The first problem is routing scalability and includes the growing size of the Internet routing table. We note that routing table size is problematic since every router is required to maintain the entire table. Consequently, we propose *ViAggre* (*Virtual Aggregation*), a scalability technique that uses tunnels to ensure that individual routers only maintain a fraction of the global routing table. ViAggre is a "configuration-only" approach to shrinking the routing table on routers – it does not require

any changes to router software and routing protocols and can be deployed independently and autonomously by any ISP. We present the design, evaluation, implementation and deployment of ViAggre to show that it can offer substantial reduction in routing table size with negligible overhead.

The second part of the thesis delves into IP Anycast. The route-to-closest-server abstraction offered by IP Anycast makes it an attractive primitive for service discovery. Further, the growth of P2P, overlay and multimedia applications presents new uses for IP Anycast. Unfortunately, IP Anycast suffers from serious limitations – it is difficult to deploy, scales poorly and lacks important features like load balancing. As a result, its use has been limited to a few critical infrastructure services like DNS root servers. Further, despite such deployments, the performance of IP Anycast and its interaction with IP routing practices is not well understood.

While these are valid concerns, we also believe that IP Anycast has compelling advantages. Motivated by these, we first conduct a *detailed study of IP Anycast* that equips us with the knowledge of how to maximize its potential. Building upon this, we present *PIAS* (*Proxy IP Anycast Service*), an anycast architecture that uses tunnels and proxies to decouple the anycast service from Internet routing. This allows PIAS to overcome IP Anycast's limitations while largely maintaining its strengths. We present simulations, measurement results, implementation and wide-area deployment details and describe how PIAS supports two important P2P and overlay applications.

BIOGRAPHICAL SKETCH

The author of this thesis hails from New Delhi, India where he spent his formative years trying to get ready the face the real world. Almost as a rite of passage, he went to an engineering school in Roorkee, a sleepy little town in Northern India. Feeling that he was not equipped to brave the rigors of the big bad world, he chose to extend his cloistered life by enrolling in Cornell's Ph.D. program. At Cornell, he spent six wonderful years relishing the myriad joys of Ithaca, a sleepy little town in Upstate New York. He also dabbled in some research on the side. While the completion of this document suggests that the author has worn out his welcome at Cornell, he still does not feel up to the challenge posed by what is "out there" and is in the midst of planning his next (possibly sleepy little town) venture.

This document is dedicated to Ma, Papa, Neha and Erin.

ACKNOWLEDGEMENTS

Paul Francis has been the perfect advisor for me. Despite the popular perception about his laissez-faire attitude towards student development, my graduate life suggests that he always had a plan for me. He nurtured me patiently while I learned the ropes of research. Then, he encouraged me to think independently, but was available if I got stuck (which I did). And when I was finally able to do research on my own, he gave me the space to do so. All this while, he has been very patient regarding all my follies.

While Paul has had a profound influence on the way I perceive networking research and think about what lies ahead, his biggest gift to me is a sense of equanimity about research and life in general. As a researcher, it is important to be able to maintain a healthy dose of skepticism, but not to cross over to cynicism. I believe that Paul has struck a good balance in this regard and I can only hope to do the same. The past six years have been very fulfilling and truly amazing, and I owe Paul a lot for that.

Andrew C. Myers and Martin T. Wells were kind enough to agree to be members of my thesis committee. I have often relied on Andrew for advice on matters outside Paul's expertise. He painstakingly read this entire thesis and his comments have contributed significantly to the quality of my writing. I have been fortunate to have had two excellent mentors, Sylvia Ratnasamy and Pablo Rodriguez, over the course of my internships. Sylvia's unbridled enthusiasm has influenced my outlook on research and I am grateful for that.

My graduate work, some of which is discussed in this thesis, includes contributions from many people. Jia Wang and Nikolaos Laoutaris were very helpful collaborators. Tuan Cao did most of the heavy lifting on the ViAggre implementation; he was a pleasure to work with. I am proud of the fact that I de-

ployed an IP Anycast testbed, and have been maintaining it for the past few years. This was made possible through the gracious help and support of the following: Eric Cronise, Dan Eckstrom and Larry Parmelee at Cornell University, James Gurganus, Phil Buonadonna, Timothy Roscoe and Sylvia Ratnasamy at Intel-Research, Kaoru Yoshida, Akira Kato and Yuji Sekiya at WIDE, and JJ Jamison and Tony Talerico at Cisco.

A lot of credit goes to people who have borne my antics and humored me during my stay at Cornell. This includes the Systems Lab crowd: Mahesh Balakrishnan, Tuan Cao, Oliver Kennedy, Jed Liu, Tudor Marian, Rohan Murty, Amar Phanishayee, Alan Shieh, Dan Williams and Xinyang Zhang. Outside the lab, Lucian Leahu, Jonathan Petrie, Radu Popovici, Vidhyashankar Venkataraman and Jonathan Winter helped me keep my mind off work. A special word of thanks to Jed, Mahesh and Vidhya for all the spots at Teagle. Finally, my life at Cornell was made much easier due to the physical proximity to Saikat Guha. His desk was next to mine, and I cannot recall the number of times he saved my skin. Kudos to him for being an amazingly smart and helpful colleague.

Graduate school would not have been possible without the support of my parents. They have worked very hard and sacrificed too much to make sure that I was able to pursue all my dalliances. I can never thank them enough. Neha and Nana kept me on my toes with their persistent yet well-meaning queries about the progress of my thesis. Hazel-cat provided much needed entertainment during the writing of this thesis. Her walks across my keyboard contributed immensely to the thesis; though I hope most of her contributions have been edited out. And last but not the least, Erin, with her love, affection, amazing culinary skills and enlightened taste in music, has added oodles of joy to my existence. Thank you all.

TABLE OF CONTENTS

LIST OF FIGURES

LIST OF TABLES

CHAPTER 1

INTRODUCTION

The explosive growth in the size of the Internet over the past couple of decades has meant that many of the assumptions that the Internet design is based on no longer hold true. This has led to a plethora of problems and has made it imperative that we rethink such assumptions and the concomitant design decisions. This, in turn, has driven work towards next-generation network architectures that are designed to cope with today's needs and challenges. Such "*blue-sky*" or "*clean-slate*" proposals have dominated the networking research arena and have focused on all aspects of network design, including but not limited to scalability [43,124,138], performance [39,74,130], security [11,27,58], management [16,56] and user control [135,136].

Clean-slate solutions tackle problems afflicting the network by addressing the underlying root cause through novel protocols and/or architectures. Examples of these root causes include improper or missing goals (for instance, network management problems resulting from manageability not being a first-class design goal [16]), invalid assumptions (for instance, security problems resulting from the Internet's outdated security model [58]) and improper coupling between entities (for instance, difficulty of traffic engineering resulting from the tight coupling between routing and traffic engineering mechanisms [46] and control plane problems resulting from intertwined decision-logic and distributed systems issues [134]).

However, the tremendous success of the Internet has also been a bane for Internet research. It is difficult, if not impossible, to expect wholesale change in Internet infrastructure. This does not reduce the relevance of architectural

research. We are convinced of the importance of such clean-slate efforts since they, apart from distilling the root cause of network problems, give us a sense of the ultimate goal for the future Internet. However, we do believe that such solutions are unlikely to ever be deployed and hence, it is equally important to find solutions that are economically viable. Keeping in line with the terminology above, we refer to such solutions as "*dirty-slate*".

While it is well accepted that it is difficult to address the Internet's problems without changing the protocols involved, we argue that in many instances, simply by focusing on a subset of the given problem space, it is possible to devise a solution that does not require architectural change. In other words, it is possible to use existing protocols in novel ways to address some of the problems. The resulting incremental solutions have a couple of important benefits. First, they offer a better alignment of cost versus benefits and hence, have a good chance of real-world adoption. Second, if the subset of problems solved happen to be the most pressing of the lot, such solutions buy the time needed for architectural proposals to mature and be deployed.

In this thesis, we explore the possibility of tackling network problems in the existing framework through dirty-slate solutions. To this effect, we recognize the potential of *tunnels* to drive dirty-slate solutions. Tunneling involves encapsulating a protocol within another protocol. Traditionally, tunnels have been used to carry packets over an incompatible underlying network. For instance, tunneling IPv6 packets over IPv4 to connect IPv6 end-sites across the IPv4 Internet [26]. Extending this, encapsulation of protocols provides a fundamental tool that can allow existing protocols to be used in new ways. For instance, tunneling of a protocol could be used to do away with an underlying assumption

that no longer holds. Similarly, tunneling could be used to separate intertwined goals that can then be addressed separately. More precisely, we argue that tunnels can be used to tackle many network problems by using the decoupling they provide to address the underlying root cause.

Recent years have seen the emergence of the use of tunnels in infrastructure networks for security (IPSec, VPNs) and performance (MPLS). Tunnels are also used in overlay networks to provide features not supported by the underlying infrastructure. Examples of such use of tunnels include Akamai [139], ESM [30], RON [8] and OverQoS [125]. Further, the increasing use of tunnels has meant that routers today ship with line-cards that can tunnel and detunnel packets at line-rates [50]. This adoption of tunneling technologies in mainstream networks means that network solutions utilizing tunnels can easily and practically be deployed on existing networks which, in turn, makes them dirty-slate solutions.

In this thesis, we focus on two specific problems afflicting the Internet: Routing Scalability and IP Anycast Scalability. For each problem, we propose a tunnel-based solution that does not require *any* change to existing network hardware and software.

- In the case of routing scalability, we use tunnels to do away with the implicit requirement that every Internet router maintain the entire global routing table and propose a technique wherein individual routers only maintain a fraction of the routing table. This shrinks the routing table size on Internet routers.
- For IP Anycast, tunnels allow separation of the anycast functionality from the underlying routing infrastructure which, in turn, is the root cause of the anycast scalability concerns.

1.1 Routing Scalability

The Internet default-free zone (DFZ) routing table has been growing rapidly for the past few years [68]. Looking ahead, there are concerns that as the IPv4 address space runs out, hierarchical aggregation of network prefixes will further deteriorate resulting in a substantial acceleration in the growth of the routing table [97]. A growing IPv6 deployment would worsen the situation even more [92].

The increase in the size of the DFZ routing table has several harmful implications for inter-domain routing, discussed in detail by Narten et al. [97].[1] At a technical level, increasing routing table size may drive high-end router design into various engineering limits. For instance, while memory and processing speeds might just scale with a growing routing system, power and heat dissipation capabilities may not [94]. A large routing table also causes routers to take longer to boot and exposes the core of the Internet to the dynamics of edge networks, thereby afflicting routing convergence. Further, routers need to forward packets at higher and higher rates while being able to access the routing table. Thus, on the business side, a rapidly growing routing table increases the cost of forwarding packets and reduces the cost-effectiveness of networks [87]. Routing table growth also makes provisioning of networks harder since it is difficult to estimate the usable lifetime of routers, not to mention the cost of the actual upgrades. Instead of upgrading their routers, a few ISPs have resorted to filtering out some small prefixes (mostly /24s) which implies that parts of the Internet may not have reachability to each other [66]. A recent private conversation with

[1] Hereon, we follow the terminology used by Rekhter et al. [107] and use the term "routing table" to refer to the Forwarding Information Base or FIB, commonly also known as the forwarding table. The Routing Information Base is explicitly referred to as the RIB. Both FIB and RIB are defined in detail later in the thesis.

a major Internet ISP revealed that in order to avoid router memory upgrades, the ISP is using a trick that reduces memory requirements but breaks BGP loop-detection and hence, would wreak havoc if adopted by other ISPs too. These anecdotes suggest that ISPs are willing to undergo some pain to avoid the cost of router upgrades.

Such concerns regarding FIB size growth, along with problems arising from a large RIB and the concomitant convergence issues, were part of the reasons that led a recent Internet Architecture Board workshop to conclude that scaling the routing system is one of the most critical challenges of near-term Internet design [94]. The severity of these problems has also prompted a slew of routing proposals [37,38,43,51,61,92,100,138]. All these proposals require changes in the routing and addressing architecture of the Internet. This, we believe, is the nature of the beast since some of the fundamental Internet design choices limit routing scalability; the overloading of IP addresses with "who" and "where" semantics represents a good example [94]. Hence the need for an architectural overhaul. However, the very fact that they require architectural change has contributed to the non-deployment of these proposals.

As mentioned earlier, we take the position that a major architectural change is unlikely and it may be more pragmatic to approach the problem through a series of incremental, individually cost-effective upgrades. Guided by this and the aforementioned implications of a rapidly growing DFZ FIB, we propose **Virtual Aggregation** or **ViAggre** [17,18], a scalability technique that focuses primarily on shrinking the FIB size on routers. ViAggre is a "configuration-only" solution that *applies to legacy routers*. Further, ViAggre can be *adopted independently and autonomously by any ISP* and hence the bar for its deployment is much lower.

The key idea behind ViAggre is very simple: an ISP adopting ViAggre essentially divides the responsibility of maintaining the global routing table amongst its routers such that individual routers only maintain a part of the routing table. ViAggre uses tunnels to ensure that packets can flow through the ISP's network in spite of the fact that routers only hold partial routing information. Thus, tunnels allow ViAggre to work around the (implicit) requirement that all routers need to maintain the complete global routing table.

1.2 IP Anycast Scalability

IP Anycast [102] is an addressing mode in which the same IP address is assigned to multiple hosts. Together, these hosts form an IP Anycast *group* and each host is referred to as an anycast *server*. Packets from a client destined to the group address are automatically routed to the anycast server closest to the client, where "closest" is in terms of the metrics used by the underlying routing protocol. Since Internet routing does not differentiate between multiple routes to multiple hosts (as in IP Anycast) and multiple routes to the same host (as in multihoming), IP Anycast is completely backward compatible requiring no changes to (IPv4 or IPv6) routers and routing protocols.

Ever since it was proposed in 1993, IP Anycast has been viewed as a powerful IP packet addressing and delivery mode. Because IP anycast typically routes packets to the nearest of a group of hosts, it has been seen as a way to obtain efficient, transparent and robust ***service discovery***. In cases where the service itself is a connectionless ***query/reply service***, IP Anycast supports the complete service, not just discovery of the service. The best working example of the latter

is the use of IP Anycast to replicate root DNS servers [1,63] without modifying DNS clients. Other proposed uses include ***host auto-configuration*** [102] and using anycast to reach a ***routing substrate***, such as rendezvous points for a multicast tree [73,81] or a IPv6 to IPv4 (6to4) transition device [67].

In spite of its benefits, there has been very little IP Anycast deployment to date, especially on a global scale. The only global scale use of IP Anycast in a production environment that we are aware of is the anycasting of DNS root servers and AS112 servers [140].[2]

We believe there are two main contributors to this limited deployment. First, despite its use in critical infrastructure services, IP Anycast and its interaction with IP routing practices is not well understood. For example, in the context of the anycasted DNS root servers, the impact of anycasting of the root servers on clients that should, in theory, access the closest server has not been analyzed in any detail. Similarly, there has been no exploration of whether root server operators can control the load on individual servers by manipulating their routing advertisements, nor of the behavior of IP Anycast under server failure. Moreover, the use of IP Anycast in different settings may rely on different assumptions about the underlying service. For example, the use of IP Anycast in Content Distribution Networks (CDNs) would require that client packets are routed to a proximal CDN server and that the impact of a server failure on clients is shortlived (*i.e.*, clients are quickly routed to a different server). To gauge the effectiveness of IP Anycast in existing deployments as also the feasibility of future usage scenarios, it is imperative to evaluate the performance of IP Anycast.

Second, IP Anycast has serious limitations. Foremost among these is IP Any-

[2] AS112 servers are anycasted servers that answer PTR queries for the RFC 1918 private addresses.

cast's poor scalability. As with IP multicast, routes for IP Anycast groups cannot be aggregated—the routing infrastructure must support one route per IP Anycast group. It is also very hard to deploy IP Anycast globally. The network administrator must obtain an address block of adequate size (i.e. a /24 or bigger), and arrange to advertise it into the BGP substrate of its upstream ISPs. Finally, the use of IP routing as the host selection mechanism means that it is not clear whether important selection metrics such as server load can be used. It is important to note that while IPv6 has defined anycast as part of its addressing architecture [64], it is also afflicted by the same set of problems.

By contrast, ***application layer anycast*** provides a one-to-any service by mapping a higher-level name, such as a DNS name, into one of a group of hosts, and then informing the client of the selected host's IP address, for instance through DNS or some redirect mechanism. This approach is much easier to deploy globally, and is in some ways superior in functionality to IP Anycast. For example, the fine grained control over the load across group members and the ability to incorporate other selection criteria makes DNS-based anycast the method of choice for CDNs today.

In spite of these valid concerns, we believe that IP Anycast has compelling advantages, and its appeal increases as overlay and P2P applications increase. First, IP Anycast operates at a low level. This makes it potentially useable by, and transparent to, any application that runs over IP. It also makes IP Anycast the only form of anycast suitable for low-level protocols, such as DNS. Second, it automatically discovers nearby resources, eliminating the need for complex proximity discovery mechanisms [4]. Finally, packets are delivered directly to the target destination without the need for a redirect (frequently required by

application-layer anycast approaches). This saves at least one packet round trip, which can be important for short lived exchanges. It is these advantages that have led to increased use of IP Anycast within the operational community, both for providing useful services (DNS root servers), and increasingly for protecting services from unwanted packets (AS112 and DDoS sinkholes [57]). Further, they have forced a re-look at the feasibility of IP Anycast based CDNs [7].

Motivated by its potential, in the second part of this thesis we study IP Anycast and how to make it more practical. Specifically, we make two main contributions: First, we present a detailed study of inter-domain IP Anycast as measured from a large number of vantage points [19]. To this effect, we focus on four properties of native IP Anycast deployments – *failover*, *load distribution*, *proximity* and *affinity*.[3] Our study uses a two-pronged approach:

1. Using a variant of known latency estimation techniques, we measure the performance of current commercially operational IP Anycast deployments from a large number (>20,000) of vantage points.
2. We deploy our own small-scale anycast service that allows us to perform controlled tests under different deployment and failure scenarios.

To the best of our knowledge, our study represents the first large-scale evaluation of existing anycast services and the first evaluation of the behavior of IP Anycast under failure.

We find that – (1) IP Anycast, if deployed in an ad hoc manner, does not offer good latency-based proximity, (2) IP Anycast, if deployed in an ad hoc

[3] Affinity measures the extent to which consecutive anycast packets from a client are delivered to the same anycast server. This and the other properties are defined in detail later.

manner, does not provide fast failover to clients, (3) IP Anycast typically offers good affinity to all clients with the exception of those that explicitly load balance traffic across multiple providers, (4) IP Anycast, by itself, is not effective in balancing client load across multiple sites. We thus propose and evaluate practical means by which anycast deployments can achieve good proximity, fast failover and control over the distribution of client load. Overall, our results suggest that an IP Anycast service, *if deployed carefully*, can offer good proximity, load balance, and failover behavior.

The aforementioned study equips us with the knowledge of how to maximize the potential of IP Anycast deployments. Building upon this, we note that most of the inherent limitations of IP Anycast arise from the tight coupling of the anycast functionality to the routing infrastructure. Guided by this observation, the second main contribution of our anycast work is the detailed proposal of a deployment architecture for an IP Anycast service that overcomes the limitations of today's "native" IP Anycast while adding new features, some typically associated with application-level anycast, and some completely new. This architecture, called **PIAS (Proxy IP Anycast Service)** [13,14], uses tunnels to decouple the anycast functionality offered to its clients from the anycast functionality provided by the Internet's routing infrastructure, i.e. "native" IP Anycast.

PIAS is composed as an overlay, and utilizes but does not impact the IP routing infrastructure. More specifically, PIAS comprises of an overlay network of proxies that advertise IP Anycast addresses on behalf of the group members and tunnels anycast packets to those members. The fact that PIAS is an IP Anycast service means that *clients* use the service completely transparently—that is, with their existing IP stacks and applications. Further, the use of IP Anycast also

entails that PIAS does not require any changes to routing infrastructure and can be (and is) deployed on the Internet today.

1.3 Outline

The rest of this thesis is organized as follows –

Chapter 2 presents background information about the use of tunnels in the Internet.

Chapter 3 presents ViAggre. Section 3.2 details the ViAggre design while section 3.3 discusses a mathematical framework capturing the trade-offs introduced by ViAggre. Section 3.4 presents evaluation results, section 3.5 details the ViAggre deployment, sections 3.6, 3.7 discuss ViAggre concerns and related work and we summarize the ViAggre proposal in section 3.8.

Chapter 4 presents a IP Anycast measurement study. Section 4.2 reviews related measurement studies, section 4.3 details the IP Anycast deployments we measure while Section 4.4 describes our measurement methodology. We describe our proximity measurements in Section 4.5, failover measurements in Section 4.6, affinity measurements in Section 4.7 and load distribution measurements in Section 4.8. Finally, we discuss related issues in Section 4.9, and summarize the study in Section 4.10.

Chapter 5 presents the PIAS architecture. Section 5.2 identifies the features of an ideal anycast service. Section 5.3 spells out the system design together with the goals satisfied by each design feature. Section 5.4 presents simulations and measurements meant to evaluate various features of the PIAS design. Sec-

tion 5.5 discusses related work. Section 5.6 describes a few applications made possible by PIAS, and we summarize the PIAS proposal in Section 5.8.

Chapter 6 presents concluding remarks.

CHAPTER 2

ALL ABOUT TUNNELS

A prime contributor to the success of the Internet in the face of its massive growth has been the principle of "layering". Internet Protocols are layered, with each layer responsible for certain services while having a fixed interface to layers above and below it. Consequently, as packets are transferred from a higher-layer protocol to a lower one, they are encapsulated in a header corresponding to the lower-layer protocol. Similarly, when packets traverse from a lower-layer to a higher-layer protocol, the packets are decapsulated, i.e. the relevant headers are stripped off. Such layering of protocols leads to protocol modularity and helps with network scalability. For instance, Subnet or layer-2 scalability is helped by the presence of IP or layer-3 protocols.

An extension of such layer-based encapsulation is the possibility of encapsulating peer protocols, i.e. protocols that operate at the same layer, in each other. This is known as Mutual Encapsulation and was first introduced by Cohen and Taft [32] to support the coexistence of the Pup protocols and IP protocols. More generally, mutual encapsulation was proposed as a means for diverse network technologies to interoperate, an important goal given the relative abundance of (competing) network technologies at that point of time.

A specific instance of mutual encapsulation is *tunneling*, which involves the encapsulation of a network protocol inside another network protocol.[1],[2] The protocol being encapsulated is known as the *payload protocol* while the encapsulating protocol is the *delivery protocol* or the *tunnel protocol*. For instance, encap-

[1]Network protocol refers to a layer-3 protocol providing end-to-end connectivity.

[2]Over the years, the term "tunnel" has been generalized to describe any irregular layering of protocols.

sulating IPv4 packets inside an IPv4 header was one of the earliest instances of tunnels [133]. The primary motivation for the use of such tunnels was to bypass routing failures and avoid broken gateways and routing domains [133]. However, the standardization and adoption of dynamic routing protocols like OSPF and BGP has meant that tunnels are no longer needed for this purpose.

Over the years, there has been a significant proliferation in tunneling technologies. This includes IP-IP [116], PPTP [62], L2TP [128], mobile-IP [41], IPSec [76], IPv6-IPv4 [26], IPmcast-IP [101], EtherIP [65], etc. While most of these tunneling technologies are geared towards a specific use scenario involving a specific payload protocol, GRE (Generic Routing Encapsulation) [44] is the only tunneling protocol standardized outside a specific context. As the name suggests, it was designed to satisfy several tunnel requirements. GRE can encapsulate different kinds of network protocols and thus, can be used to create virtual point-to-point links across the Internet. Example GRE usage scenarios include its use with PPTP to create Virtual Private Networks (VPNs), to provide routing functionality in IPsec-based VPNs and in mobility protocols. Finally, one of the most popular tunneling protocols in use today is MPLS [110].[3] MPLS allows creation of tunnels, known as MPLS Label Switched Paths (LSPs), between wide-area nodes in a scalable fashion. MPLS tunnels are agnostic to the protocol being encapsulated and offer high forwarding performance. This, in turn, has led to extensive MPLS deployment in Internet ISPs where its use ranges from building a BGP-less core to MPLS-based VPNs.

The diverse set of tunneling technologies mentioned above serve in a lot of different settings. However, their use can be broadly classified into three main

[3]Strictly speaking, MPLS is a layer-2 protocol and hence, IP-inside-MPLS is just normal encapsulation. However, MPLS Label Switched Paths (LSPs) are considered as tunnels because MPLS is often seen as a network protocol.

categories [103]:

1. Feature support: Tunnels can be used to create a virtual network that provides features or serves goals not satisfied by the underlying network. Examples of such goals include security properties, isolation properties, performance guarantees, etc. For instance, tunneling technologies such as IPsec create a point-to-point virtual link that provides security guarantees across the underlying (insecure) network. Similarly, PPTP and L2TP tunnels allow for a virtual PPP link to be extended across the Internet and are mainly used for the creation of VPNs.

2. Protocol Evolution: Packets can be tunneled to make them traverse an incompatible network and thus, tunneling presents an ideal vehicle for incremental deployment of new network protocols. For instance, IPv6 packets are tunneled over IPv4 to connect IPv6 end-sites across the IPv4 Internet [26]. Similarly, multicast IPv4 packets are tunneled over IPv4 to reach the mbone network, thus allowing for an incrementally growing IP Multicast deployment. On the research side, VINI [23] aims to create a virtual network infrastructure for experimental research and is a good example of such use of tunnels.

3. Interface Preservation: Protocols can be tunneled so as to retain backwards compatibility with existing systems. This generally involves a lower-layer protocol serving as the payload and a higher-layer protocol serving as the delivery protocol. For instance, EtherIP involves tunneling Ethernet over IP and is used to provide layer-2 connectivity between geographically distributed end-sites.

We note that a theme common across the three categories of tunnel usage mentioned above is the creation of virtual networks over the existing network through the application of tunnels. Such a virtual network can then be used to provide new features, support new protocols or even retain backwards compatibility. In this thesis, we extend the application of tunnels and hypothesize that it is possible to design virtual networks to alleviate specific problems afflicting the Internet. This involves using the decoupling abilities of tunnels to address the root cause of the network problem. In the rest of this thesis, we illustrate this by using tunnels to tackle two very different yet important problems facing the Internet.

CHAPTER 3

VIRTUAL AGGREGATION (VIAGGRE)

3.1 Background and Contributions

Internet routers participate in routing protocols to establish routes to network destinations. The domain-based structure of the Internet has led to a two-tiered routing architecture wherein *intra-domain routing protocols* establish routes within a domain while inter-domain routing protocols establish routes between domains. Examples of intra-domain routing protocols include OSPF, RIP, IS-IS while examples of inter-domain routing protocols include BGP, IDRP. In the Internet, BGP (BGPv4 for IPv4 routes and BGPv6 for IPv6 routes) serves as the de-facto inter-domain routing protocol and allows routers in a domain to determine routes to publicly-reachable destinations in other domains. Such destinations are represented by their *network prefix*, i.e. the block of addresses assigned to them. The set of publicly-reachable prefixes on the Internet is referred to as the global routing table.

The routing information learned by a router through its participation in routing protocols is stored in what is effectively a database of routes and is referred to as the *Routing Information Base* (RIB). As shown in figure 3.1, the RIB is part of the router's control plane and is optimized for efficient updating by routing protocols. Note that the RIB can and often does contain multiple routes to the same destination. The router then uses a decision algorithm to select the best path to each destination that is installed on the router's forwarding plane (data plane). This set of routes is referred to as the routing table or the forwarding table or the Forwarding Information Base (FIB). Since the FIB is used to forward

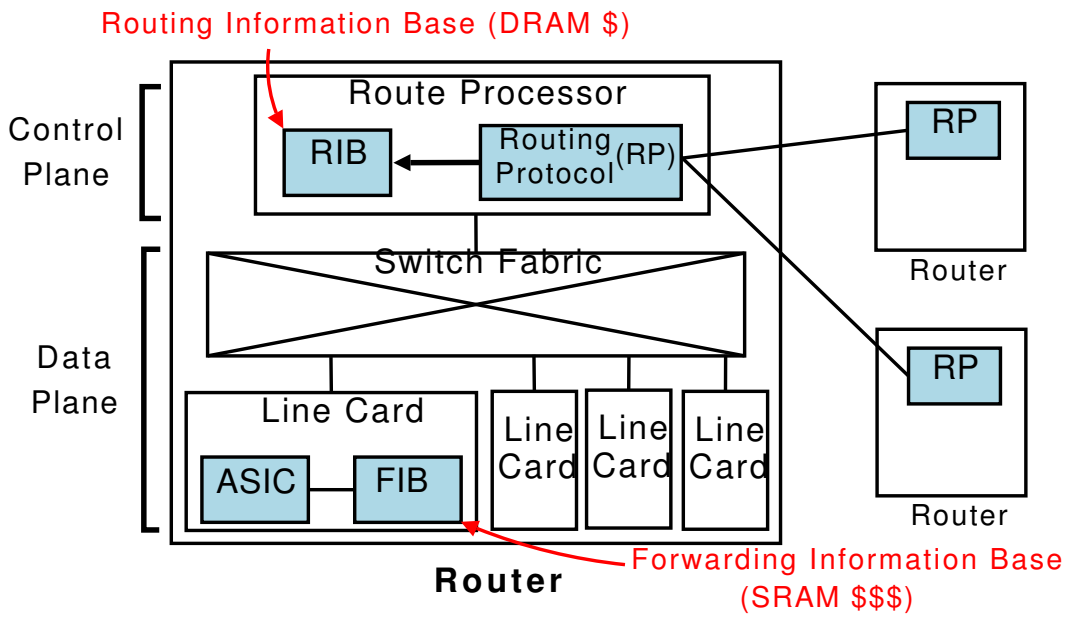


Figure 3.1: Router Innards: A router exchanging routing information with two neighboring routers.

data packets, it needs to be accessible at lines rates and thus, resides on fast (and expensive) memory.

Internet growth has caused more and more prefixes to be advertised into the Internet, resulting in a larger routing table. This entails a larger FIB and RIB for Internet routers. From a historical perspective, the scalability of the Internet's routing system in the face of such growth has relied on topological hierarchy which, in turn, requires that the addressing of domains in the Internet be in line with the actual physical topology. Such alignment of addressing and topology leads to a routing hierarchy wherein destination addresses or network prefixes can be aggregated as they propagate up the hierarchy.

However, Internet growth has also meant that Internet addressing is no longer aligned with the actual physical topology. For instance, "site multihoming" wherein an end-site connects to multiple upstream providers, leads to an

address-topology mismatch and makes it impossible for the site's providers to aggregate the site's prefixes. Other factors that can cause such a mismatch between addressing and topology include traffic engineering by ISPs, address fragmentation and bad operational practices, including operator laziness. Further, studies have shown that this mismatch between addressing and topology is the root cause for the rapid growth in the Internet's routing table [94,138].

Most past proposals to improve routing scalability recognize this mismatch and propose architectural changes to ensure that topology follows addressing or vice versa. However, the need for change has proven to be a significant obstacle to deployment. Instead of focusing on the address-topology mismatch and reducing the size of the global routing table, we argue that an alternative way to approach the problem is to reduce the amount of routing state that individual routers are required to maintain. Today, every router in the default-free zone (DFZ) of the Internet maintains the entire global routing table. In this thesis, we engineer a routing design that obviates this (implicit) requirement.

To this effect, we propose *Virtual Aggregation* (ViAggre), a scalability technique that allows an ISP to modify its internal routing such that individual routers in the ISP's network only maintain a part of the global routing table. ViAggre is a "configuration-only" approach to shrinking the routing table on routers. Consequently, ViAggre does not require any changes to router software and routing protocols and can be deployed independently and autonomously by any ISP. Hence, the ViAggre part of this thesis makes the following contributions:

- We discuss two deployment options through which an ISP can adopt ViAggre. The first one uses *FIB suppression* to shrink the FIB of all the ISP's

routers while the second uses *route filtering* to shrink both the FIB and RIB on all *data-path* routers.

- We analyze the application of ViAggre to an actual tier-1 ISP and several inferred (Rocketfuel [118]) ISP topologies. We find that ViAggre can reduce FIB size by more than an order of magnitude with negligible stretch on the ISP's traffic and very little increase in load across the ISP's routers. Based on predictions of future routing table growth, we estimate that ViAggre can be used to extend the life of already outdated routers by more than 10 years.
- We propose utilizing the notion of prefix popularity to reduce the impact of ViAggre on the ISP's traffic and use a two-month study of a tier-1 ISP's traffic to show the feasibility of such an approach.
- As a proof-of-concept, we configure test topologies comprising of Cisco routers (on WAIL [21]) according to the ViAggre proposal. We use the deployment to benchmark the control plane processing overhead that ViAggre entails. One of the presented designs actually reduces the amount of processing done by routers and preliminary results show that it can reduce convergence time too. The other design has high overhead due to implementation issues and needs more experimentation.
- ViAggre involves the ISP reconfiguring its routers which can be a deterrent to adoption. We quantify this configuration overhead. We also implement a configuration tool that, given the ISPs existing configuration files, can automatically generate the configuration files needed for ViAggre deployment. We discuss the use of this tool on our testbed.

Overall, the incremental version of ViAggre presented in this chapter can be

seen as little more than a simple and structured hack that assimilates ideas from existing work including, but not limited to, VPN tunnels and CRIO [138]. We believe that its very simplicity makes ViAggre an attractive short-term solution that provides ISPs with an alternative to upgrading routers in order to cope with routing table growth till more fundamental, long-term architectural changes can be agreed upon and deployed in the Internet. However, the basic ViAggre idea can also be applied in a clean-slate fashion to address routing concerns beyond FIB growth. While we defer the design and the implications of such a non-incremental ViAggre architecture for future work, the notion that ViAggre can serve both as an immediate alleviative and as the basis for a next-generation routing architecture seems interesting and worth exploring.

3.2 ViAggre design

ViAggre applies to both the IPv4 and IPv6 routing tables. However, in the rest of this chapter, we focus on the reduction of the IPv4 routing table which is referred to as the global routing table.

ViAggre allows individual ISPs in the Internet's DFZ to do away with the need for their routers to maintain routes for all prefixes in the global routing table. An ISP adopting ViAggre divides the global address space into a set of *virtual prefixes* such that the virtual prefixes are larger than any aggregatable (real) prefix in use today. So, for instance, an ISP could divide the IPv4 address space into 128 parts with a /7 virtual prefix representing each part (0.0.0.0/7 to 254.0.0.0/7). Note that such a naïve allocation would yield an uneven distribution of real prefixes across the virtual prefixes. However, the virtual prefixes

need not be of the same length and hence, the ISP can choose them such that they contain a comparable number of real prefixes.

The virtual prefixes are not topologically valid aggregates, i.e. there is not a single point in the Internet topology that can hierarchically aggregate the encompassed prefixes. ViAggre makes the virtual prefixes aggregatable by organizing *virtual networks*, one for each virtual prefix. In other words, a virtual topology is configured that causes the virtual prefixes to be aggregatable, thus allowing for routing hierarchy that shrinks the routing table. To create such a virtual network, some of the ISP's routers are assigned to be within the virtual network. These routers maintain routes for all prefixes in the virtual prefix corresponding to the virtual network and hence, are said to be *aggregation points* for the virtual prefix. A router can be an aggregation point for multiple virtual prefixes and is required to only maintain routes for prefixes in the virtual prefixes it is aggregating.

Given this, a packet entering the ISP's network is routed to a close-by aggregation point for the virtual prefix encompassing the actual destination prefix. This aggregation point has a route for the destination prefix and forwards the packet out of the ISP's network in a tunnel. In figure 3.3 (figure details explained later), router C is an aggregation point for the virtual prefix encompassing the destination prefix and B $\rightarrow$ C $\rightarrow$ D is one such path through the ISP's network.

3.2.1 Design Goals

The discussion above describes ViAggre at a conceptual level. While the design space for organizing an ISP's network into virtual networks has several

dimensions, this thesis aims for deployability and hence is guided by two major design goals:

1. *No changes to router software and routing protocols*: The ISP should not need to deploy new data-plane or control plane mechanisms.
2. *Transparent to external networks*: An ISP's decision to adopt the ViAggre proposal should not impact its interaction with its neighbors (customers, peers and providers).

These goals, in turn, limit what can be achieved through the ViAggre designs presented here. As explained earlier, routers today have a Routing Information Base (RIB) generated by the routing protocols and a Forwarding Information Base (FIB) that is used for forwarding the packets. Consequently, the FIB is optimized for looking up destination addresses and is maintained on fast(er) memory, generally on the line cards themselves [97]. All things being equal, it would be nice to shrink both the RIB and the FIB for all ISP devices, as well as make other improvements such as shorter convergence time.

While the basic ViAggre idea can be used to achieve these benefits (section 3.6), we have not been able to reconcile them with the aforementioned design goals. Instead, our work is based on the hypothesis that given the performance and monetary implications of the FIB size for routers, an immediately deployable solution that reduces FIB size is useful. Actually, one of the presented designs also shrinks the RIB on routers; only components that are off the data path (i.e., route reflectors) need to maintain the full RIB. Further, this design is shown to help with route convergence time too.

3.2.2 Design-I: FIB Suppression

This section details one way an ISP can deploy virtual prefix based routing while satisfying the goals specified in the previous section. The discussion below applies to IPv4 (and BGPv4) although the techniques detailed here work equally well for IPv6. The key concept behind this design is to operate the ISP's internal distribution of BGP routes untouched and in particular, to populate the RIB on routers with the full routing table but to suppress most prefixes from being loaded in the FIB of routers. A standard feature on routers today is *FIB Suppression*, which can be used to prevent routes for individual prefixes in the RIB from being loaded into the FIB. We have verified support for FIB suppression as part of our ViAggre deployment on Cisco 7300 and 12000 routers. Documentation for Juniper [153] and Foundry [152] routers specify this feature too. We use this as described below.

The ISP does not modify its routing setup — the ISP's routers participate in an intra-domain routing protocol that establishes internal routes through which the routers can reach each other, while BGP is used for inter-domain routing just as today. For each virtual prefix, the ISP designates some number of routers to serve as aggregation points for the prefix and hence, to form a virtual network. Each router is configured to only load prefixes belonging to the virtual prefixes it is aggregating into its FIB, while suppressing all other prefixes.

Given such assignment of aggregation points to routers, the ISP needs to ensure that packets to any prefix can flow through the network in spite of the fact that only a few routers have a route to the prefix. This goal is achieved as follows:

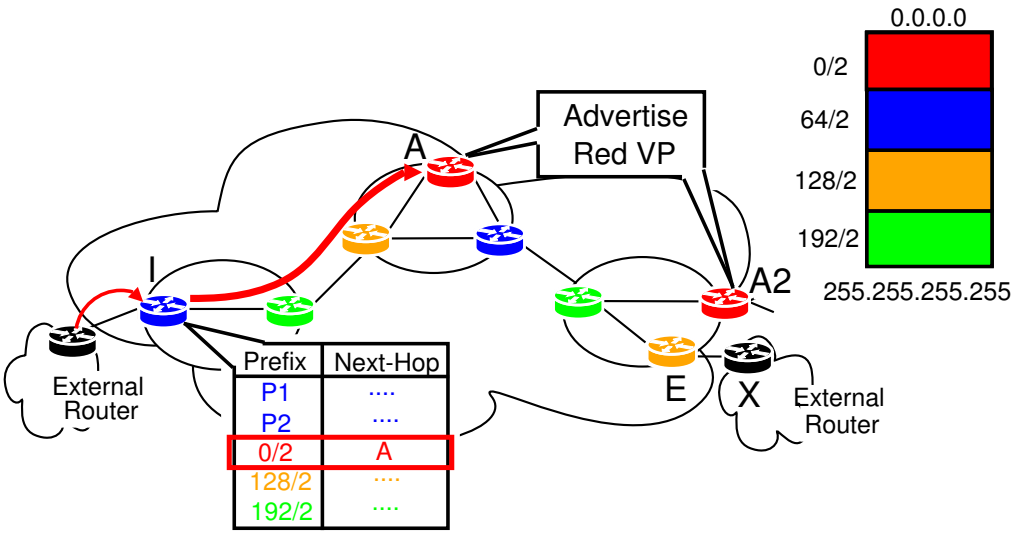


Figure 3.2: A ViAggre ISP with four virtual prefixes (0/2, 64/2, 128/2, 192/2). The virtual prefixes are color-coded with each router serving as an aggregation point for the corresponding color. The red routers are aggregation points for the 0/2 virtual prefix and advertise it into the ISP's internal routing.

– *Connecting Virtual Networks.* Aggregation points for a virtual prefix originate a route to the virtual prefix that is distributed throughout the ISP's network but not outside. Specifically, an aggregation point advertises the virtual prefix to its iBGP peers. For instance, in figure 3.2, the two red routers aggregate the 0.0.0.0/2 virtual prefix (denoted by the red part of the IPv4 address space and labeled as 0/2) and hence, advertise 0/2 to their iBGP peers. A router that is not an aggregation point for the virtual prefix would choose the route advertised by the aggregation point closest to it and hence, forward packets destined to any prefix in the virtual prefix to this aggregation point.[1] In figure 3.2, router I forwards packets destined to a prefix in 0/2 to red aggregation point A.

[1]All other attributes for the routes to a virtual prefix are the same and hence, the decision is based on the IGP metric to the aggregation points. Hence, "closest" means closest in terms of IGP metric.

– *Sending packets to external routers.* When a router receives a packet destined to a prefix in a virtual prefix it is aggregating, it can look up its FIB to determine the route for the packet. However, such a packet cannot be forwarded in the normal hop-by-hop fashion since a router that is not an aggregation point for the virtual prefix in question might forward the packet back to the aggregation point, resulting in a loop. Hence, the packet must be tunneled from the aggregation point to the external router that was selected as the BGP NEXT_HOP. While the ISP can probably choose from many tunneling technologies, we use MPLS Label Switched Paths (LSPs) for such tunnels. This choice was influenced by the fact that MPLS is widely supported in routers, is used by ISPs, and operates at wire speed. Further, protocols like LDP [9] automate the establishment of MPLS tunnels and hence, reduce the configuration overhead.

However, a LSP from the aggregation point to an external router would require cooperation from the neighboring ISP. To avoid this, every edge router of the ISP initiates a LSP for every external router it is connected to. Thus, all the ISP routers need to maintain LSP mappings equal to the number of external routers connected to the ISP, a number much smaller than the routes in the DFZ routing table (we relax this constraint in Section 3.4.2). Note that even though the tunnel endpoint is the external router, the edge router can be configured to strip the MPLS label from the data packets before forwarding them onto the external router. This, in turn, has two implications. First, external routers don't need to be aware of the adoption of ViAggre by the ISP. Second, even the edge router does not need a FIB entry for the destination prefix. Instead, it chooses the external router to forward the packets to based on the MPLS label of the packet. The behavior of the edge router here is similar to the penultimate hop

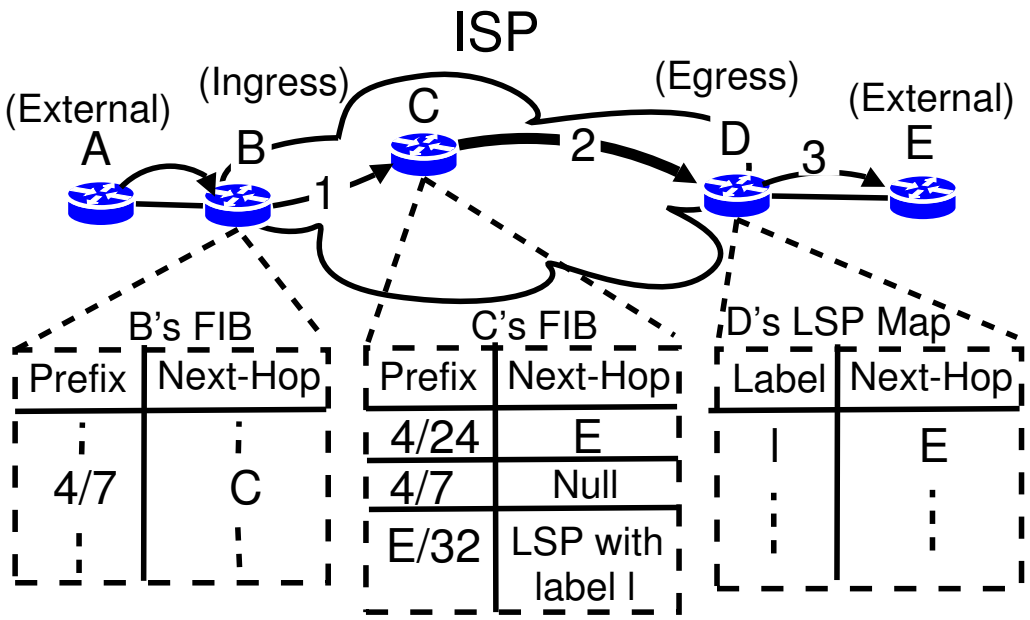


Figure 3.3: Path of packets destined to prefix 4.0.0.0/24 (or, 4/24) between external routers A and E through an ISP with ViAggre. Router C is an aggregation point for virtual prefix 4.0.0.0/7 (or, 4/7).

in a VPN scenario and is achieved through standard configuration.

We now use a concrete example to illustrate the flow of packets through an ISP network that is using ViAggre. Figure 3.3 shows the relevant routers. The ISP is using /7s as virtual prefixes and router C is an aggregation point for one such virtual prefix 4.0.0.0/7. Edge router D initiates a LSP to external router E with label l and hence, the ISP's routers can get to E through MPLS tunneling. The figure shows the path of a packet destined to prefix 4.0.0.0/24, which is encompassed by 4.0.0.0/7, through the ISP's network. The path from the ingress router B to the external router E comprises three segments:

1. VP-routed: Ingress router B is not an aggregation point for 4.0.0.0/7 and hence, forwards the packet to aggregation point C.
2. MPLS-LSP: Router C, being an aggregation point for 4.0.0.0/7, has a route

for 4.0.0.0/24 with BGP NEXT_HOP set to E. Further, the path to router E involves tunneling the packet with MPLS label *l*.

3. Map-routed: On receiving the tunneled packet from router C, egress router D looks up its MPLS label map, strips the MPLS header and forwards the packet to external router E.

3.2.3 Design-II: Route Reflectors

The second design offloads the task of maintaining the full RIB to devices that are off the data path. Many ISPs use route reflectors for scalable internal distribution of BGP prefixes, and we require only these route reflectors to maintain the full RIB.

ISP networks are composed of points of presence or PoPs that are connected via a backbone network. Each PoP is a physical location that houses the ISP's equipment. For ease of exposition, we assume that the ISP is already using per-PoP route reflectors that are off the data path, a common deployment model for ISPs using route reflectors.

In the proposed design, the external routers connected to a PoP are made to peer with the PoP's route reflector. This is necessary since the external peer may be advertising the entire DFZ routing table and we don't want all these routes to reside on any given data-plane router. The route reflector also has iBGP peerings with other route reflectors and with the routers in its PoP. Egress filters are used on the route reflector's peerings with the PoP's routers to ensure that a router only gets routes for the prefixes it is aggregating. This shrinks both the RIB and

the FIB on the routers. The data-plane operation and hence, the path of packets through the ISP's network, remains the same as with the previous design.

With this design, a PoP's route reflector peers with all the external routers connected to the PoP. The RIB size on a BGP router depends on the number of peers it has and hence, the RIB for the route reflectors can potentially be very large. If needed, the RIB requirements can be scaled by using multiple route reflectors. Note that the RIB scaling properties here are better than in the status quo. Today, edge routers have no choice but to peer with the directly connected external routers and maintain the resulting RIB. Replicating these routers is prohibitive because of their cost but the same does not apply to off-path route reflectors, which could even be BGP software routers.

3.2.4 Design Comparison

As far as the configuration is concerned, configuring suppression of routes on individual routers in design-I is comparable, at least in terms of complexity, to configuring egress filters on the route reflectors. In both cases, the configuration can be achieved through BGP route-filtering mechanisms (access lists, prefix lists, etc.).

Design-II, apart from shrinking the RIB on the routers, does not require the route suppression feature on routers. Further, as we detail in Section 3.5.2, design-II reduces the ISP's route propagation time while the specific filtering mechanism used in design-I increases it. However, design-II does require the ISP's eBGP peerings to be reconfigured which, while straightforward, violates our goal of not impacting neighboring ISPs.

3.2.5 Network Robustness

ViAggre causes packets to be routed through an aggregation point, which leads to robustness concerns. When an aggregation point for a virtual prefix fails, routers using that aggregation point are rerouted to another aggregation point through existing mechanisms without any explicit configuration by the ISP. In case of design-I, a router has routes to all aggregation points for a given virtual prefix in its RIB and hence, when the aggregation point being used fails, the router installs the second closest aggregation point into its FIB and packets are rerouted almost instantly. With design-II, it is the route reflector that chooses the alternate aggregation point and advertises this to the routers in its PoP. Hence, as long as another aggregation point exists, failover happens automatically and quickly.

3.2.6 Routing popular prefixes natively

The use of aggregation points implies that packets in ViAggre may take paths that are longer than native paths. Apart from the increased path length, the packets may incur queuing delay at the extra hops. In Section 3.3, we discuss how the ISP can assign aggregation points to its routers so as to minimize the extra delay imposed on packets.

The extra hops also result in an increase in load on the ISP's routers and links and a modification in the distribution of traffic across them. We tackle this as follows. Past studies have shown that a large majority of Internet traffic is destined to a very small fraction of prefixes [42,48,108,127]. The fact that routers today have no choice but to maintain the complete DFZ routing table implies

that this observation wasn't very useful for routing configuration. However, with ViAggre, individual routers only need to maintain routes for a fraction of prefixes. The ISP can thus configure its ViAggre setup such that the small fraction of popular prefixes are in the FIB of every router. This ensures that a vast majority of the ISP's traffic is routed natively and follows the shortest path. Only a very small fraction of traffic is routed through the aggregation points and hence, there is limited increase in load across the ISP's routers and links.

For design-I, the use of popular prefixes involves configuring each router with a set of prefixes that should not be suppressed from the FIB. For design-II, each PoP's route reflector is configured to not filter advertisements for popular prefixes from the PoP's routers. Beyond this, the ISP may also choose to install customer prefixes into its routers such that they don't incur any stretch. The rest of the proposal involving virtual prefixes remains the same and ensures that individual routers only maintain routes for a fraction of the unpopular prefixes. In section 3.4.2, we analyze Netflow data from a tier-1 ISP network to show that not only is such an approach feasible, it ensures that the extra router load is negligible.

3.3 Allocating aggregation points

An ISP adopting ViAggre would obviously like to minimize the stretch imposed on its traffic. Ideally, the ISP would deploy an aggregation point for all virtual prefixes in each of its PoPs. This would ensure that for every virtual prefix, a router chooses the aggregation point in the same PoP and hence, the traffic stretch is minimal. However, this may not be possible in practice. This is be-

cause ISPs, including tier-1 ISPs, often have some small PoPs with just a few routers. Therefore, there may not be enough cumulative FIB space in the PoP to hold all the actual prefixes. For instance, an analysis of the Rocketfuel topologies [118] of 10 tier-1 and tier-2 ISPs shows that 6 ISPs have at least one PoP of size 2 while 3 have at least one PoP of size 3. Note that assigning all virtual prefixes to a PoP with two routers would imply that even in the best case scenario, each of the routers would have to maintain half of the Internet routing table. More generally, ISPs may be willing to bear some stretch for substantial reductions in FIB size. To achieve this, the ISP needs to be smart about the way it designates routers to aggregate virtual prefixes. In this section we explore this choice.

3.3.1 Problem Formulation

We first introduce the notation used in the rest of this section. Let T represent the set of prefixes in the Internet routing table, R be the set of ISP's routers and X is the set of external routers directly connected to the ISP. For each $r \in R$, P_r represents the set of popular prefixes for router r. V is the set of virtual prefixes chosen by the ISP and for each $v \in V$, n_v is the number of prefixes in v. We use two matrices, $D = (d_{i,j})$ that gives the distance between routers i and j and $W = (w_{i,j})$ that gives the IGP metric for the IGP-established path between routers i and j. We also define two relations:

– "BelongsTo" relation B: $T \rightarrow V$ such that $B(p)=v$ if prefix p belongs to or is encompassed by virtual prefix v.

– "Egress" relation E: R x $T \rightarrow R$ such that $E(i, p)=j$ if traffic to prefix p from

router i egresses at router j.

The mapping relation A: $R \rightarrow 2^V$ captures how the ISP assigns aggregation points; i.e. $A(r) = \{v_1 \ldots v_n\}$ implies that router r aggregates virtual prefixes $\{v_1 \ldots v_n\}$. Given this assignment, we can determine the aggregation point any router uses for its traffic to each virtual prefix. This is captured by the "Use" relation U: R x $V \rightarrow R$ where $U(i, v) = j$ or router i uses aggregation point j for virtual prefix v if the following conditions are satisfied:

$$\begin{aligned} &1) \quad v \in A(j) \\ &2) \quad w_{i,j} \leq w_{i,k} \quad \forall k \in R, v \in A(k) \end{aligned}$$

Here, condition 1) ensures that router j is an aggregation point for virtual prefix v. Condition 2) captures the operation of BGP with design-I and ensures that a router chooses the aggregation point that is closest in terms of IGP metrics.[2]

Using this notation, we can express the FIB size on routers and the stretch imposed on traffic.

Routing State

In ViAggre, a router needs to maintain routes to the (real) prefixes in the virtual prefixes it is aggregating, routes to all the virtual prefixes themselves and routes to the popular prefixes. Further, the router needs to maintain LSP mappings for LSPs originated by the ISP's edge routers with one entry for each external router connected to the ISP. Hence, the "routing state" for the router r, simply referred to as the FIB size (F_r), is given by:

[2]With design-II, a router chooses the aggregation point closest to the router's route reflector in terms of IGP metrics and so a similar formulation works for the second design too.

$$F_r = \sum_{v \in A(r)} n_v + |V| + |P_r| + |X|$$

The **Worst FIB size** and the **Average FIB size** are defined as follows:

$$\text{Worst FIB size} = \max_{r \in R}(F_r)$$

$$\text{Average FIB size} = \sum_{r \in R} (F_r)/|R|$$

Traffic Stretch

If router i uses router k as an aggregation point for virtual prefix v, packets from router i to a prefix p belonging to v are routed through router k. Hence, the stretch (S) imposed on traffic to prefix p from router i is given by:

$$\begin{aligned} S_{i,p} &= 0, && p \in P_i \\ &= (d_{i,k} + d_{k,j} - d_{i,j}), && p \in (T - P_i),\ v = B(p) \\ & && k = U(i,v)\ \&\ j = E(k,p) \end{aligned}$$

The **Worst Stretch** and **Average Stretch** are defined as follows:

$$\text{Worst Stretch} = \max_{i \in R, p \in T}(S_{i,p})$$

$$\text{Average Stretch} = \sum_{i \in R, p \in T} (S_{i,p})/(|R| * |T|)$$

Problem: ViAggre shrinks the routing table on routers by ensuring that individual routers only maintain routes to a fraction of the prefixes and forward packets to an aggregation point for the rest. Thus, through the use of aggregation points, ViAggre trades off an increase in path length for a reduction in routing state. The ISP can use the assignment of aggregation points as a knob to tune this trade-off. Here we consider the simple goal of minimizing the FIB

Size on the ISP's routers while bounding the stretch. Specifically, the ISP needs to assign aggregation points by determining a mapping A that

$$\begin{aligned} \min \quad & \text{Worst FIB Size} \\ \text{s.t.} \quad & \text{Worst Stretch} \leq C \end{aligned}$$

where C is the specified constraint on Worst Stretch. Note that much more complex formulations are possible. Our focus on worst-case metrics is guided by practical concerns – the Worst FIB Size dictates how the ISP's routers need to be provisioned, while the Worst Stretch characterizes the most unfavorable impact of the use of ViAggre. Specifically, bounding the Worst Stretch allows the ISP to ensure that its existing SLAs are not breached and applications sensitive to increase in latency (example, VOIP) are not adversely affected. A dual version of the assignment problem is to minimize the Worst Stretch while constraining the Worst FIB Size. The solution for this formulation is analogous to the one described in the next section.

3.3.2 A Greedy Solution

The problem of assigning aggregation points is similar to the facility location problem studied in the theory community. As a matter of fact, the MultiCommodity Facility Location (MCL) problem [106] can be mapped to the problem of assigning aggregation points while satisfying the conditions above. Using the MCL terminology, this involves routers representing facilities, virtual prefixes being commodities and each router's traffic to virtual prefixes serving as clients. MCL is NP-hard and hence, so is the assignment of aggregation points. Ravi et al. [106] present a logarithmic approximation algorithm for MCL. Here we

discuss a greedy approximation solution for assignment of aggregation points, similar to the algorithm in [80].

The first solution step is to determine that if router i were to aggregate virtual prefix v, which routers can it serve without violating the stretch constraint. This is the $can\_serve_{i,v}$ set and is defined as follows:

$$\begin{aligned} can\_serve_{i,v} \quad = \quad & \{j \mid j \in R, (\forall p \in T, B(p) = v, E(i,p) \\ & = k, (d_{j,i} + d_{i,k} - d_{j,k}) \leq C)\} \end{aligned}$$

Given this, the key idea behind the solution is that any assignment based on the *can_serve* relation will have Worst Stretch less than C. Hence, our algorithm designates routers to aggregate virtual prefixes in accordance with the *can_serve* relation while greedily trying to minimize the Worst FIB Size. The algorithm, shown below, stops when each router can be served by at least one aggregation point for each virtual prefix.

1: $Worst\_FIB\_Size$=0
2: **for all** r in R **do**
3: **for all** v in V **do**
4: Calculate $can\_serve_{r,v}$
5: **end for**
6: **end for**
7: Sort V in decreasing order of n_v
8: **for all** v in V **do**
9: Sort R in decreasing order of $|can\_serve_{r,v}|$
10: **repeat**
11: **for all** r in R **do**
12: **if** $(F_r + n_v) \leq Worst\_FIB\_Size$ **then**
13: A[r]=A[r] $\cup$ v {Assign v to r}

14: $F_r = F_r + n_v$ {r's FIB size increases}
15: Mark all routers in $can\_serve_{r,v}$ as served
16: **end if**
17: **if** All routers are served for v **then**
18: break {Break from the for loop around R}
19: **end if**
20: **end for**
21: **if** All routers are not served for v **then**
22: {$Worst\_FIB\_Size$ needs to be raised}
23: **for all** r in R **do**
24: **if** v $\notin$ A[r] **then**
25: {r is not an aggregation point for v}
26: A[r]=A[r] $\cup$ v
27: $F_r = F_r + n_v$
28: $Worst\_FIB\_Size = F_r$
29: break
30: **end if**
31: **end for**
32: **end if**
33: **until** All Routers are served for virtual prefix v
34: **end for**

3.4 Evaluation

In this section we evaluate the application of ViAggre to a few Internet ISPs. The main results presented here are:

– Using data from a tier-1 ISP, we show that ViAggre can reduce the FIB size by a factor of more than 10 with negligible stretch on the ISP's traffic and a very small increase in router load. Given predictions of future routing table growth, we find that ViAggre would allow ISPs to extend the life of outdated routers by more than 10 years.

– Based on a two-month-long study of the ISP's traffic, we conclude that prefix popularity can indeed be used to minimize the impact of ViAggre on the ISP's traffic.

– We analyze the application of ViAggre to the Rocketfuel topologies of 10 ISPs, and conservative estimates show that in the worst case, the FIB size on the ISP's routers is reduced to less than 15% of the DFZ routing table.

3.4.1 Metrics of Interest

We defined (**Average** and **Worst**) **FIB Size** and **Stretch** metrics in section 3.3.1. The use of ViAggre also affects the flow of the ISP's traffic and the load on its routers. Here we define the metrics concerning these that we then use for ViAggre evaluation.

Impact on Traffic

Apart from the stretch imposed, another aspect of ViAggre's impact is the amount of traffic affected. To account for this, we define **traffic impacted** as the fraction of the ISP's traffic that uses a different router-level path than the native path. Note that in many cases, a router will use an aggregation point for

the destination virtual prefix in the same PoP and hence, the packets will follow the same PoP-level path as before. Thus, another metric of interest is the **traffic stretched**, the fraction of traffic that is forwarded along a different PoP-level path than before. In effect, this represents the change in the distribution of traffic across the ISP's inter-PoP links and hence, captures how ViAggre interferes with the ISP's inter-PoP traffic engineering.

Impact on Router Load

The extra hops traversed by traffic increases the traffic load on the ISP's routers. We define the **load increase** across a router as the extra traffic it needs to forward due to ViAggre, as a fraction of the traffic it forwards natively.

3.4.2 Tier-1 ISP Study

We analyzed the application of ViAggre to a large tier-1 ISP in the Internet. For our study, we obtained the ISP's router-level topology (to determine router set R) and the routing tables of routers (to determine prefix set T and the Egress E and BelongsTo B relations). We used information about the geographical locations of the routers to determine the distance matrix D such that $d_{i,j}$ is 0 if routers i and j belong to the same PoP (and hence, are in the same city); otherwise $d_{i,j}$ is set to the propagation latency corresponding to the great circle distance between i and j. Further, we did not have information about the ISP's link weights. However, guided by the fact that intra-domain traffic engineering is typically latency-driven [117], we use the Distance matrix D as the weight matrix W. We also obtained the ISP's traffic matrix; however, in order to char-

acterize the impact of vanilla ViAggre, the first part of this section assumes that the ISP does not consider any prefixes as popular.

Deployment decisions

The ISP, in order to adopt ViAggre, needs to decide what virtual prefixes to use and which routers aggregate these virtual prefixes. We describe the approaches we evaluated.

– *Determining set V*. The most straightforward way to select virtual prefixes while satisfying the two conditions specified in section 3.2 is to choose large prefixes (/6s, /7s, etc.) as virtual prefixes. We assume that the ISP uses /7s as its virtual prefixes and refer to this as the "/7 allocation".

However, such selection of virtual prefixes could lead to a skewed distribution of (real) prefixes across them with some virtual prefixes containing a large number of prefixes. For instance, using /7s as virtual prefixes implies that the largest virtual prefix (202.0.0.0/7) contains 22,772 of the prefixes in today's BGP routing table or 8.9% of the routing table. Since at least one ISP router needs to aggregate each virtual prefix, such large virtual prefixes would inhibit the ISP's ability to reduce the Worst FIB size on its routers. However, as we mentioned earlier, the virtual prefixes need not be of the same length and so large virtual prefixes can be split to yield smaller virtual prefixes. To study the effectiveness of this approach, we started with /7s as virtual prefixes and split each of them such that the resulting virtual prefixes were still larger than any prefix in the Internet routing table. This yielded 1024 virtual prefixes with the largest containing 4,551 prefixes or 1.78% of the BGP routing table. We also use this virtual

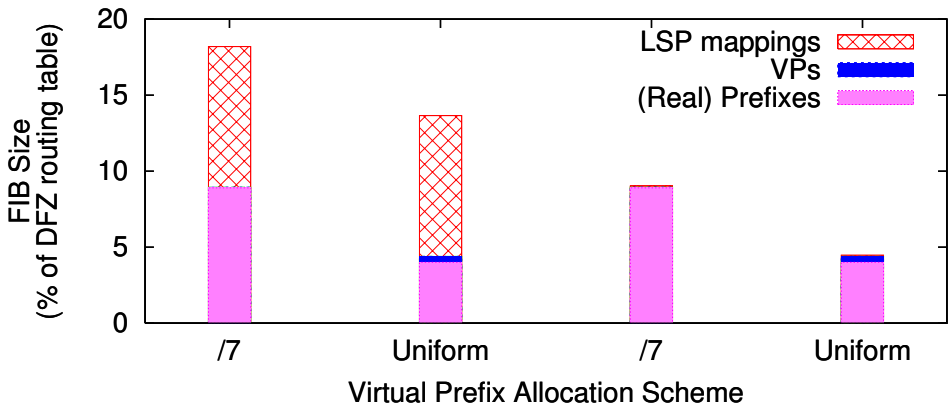


Figure 3.4: FIB composition for the router with the largest FIB, C=4ms and no popular prefixes.

prefix allocation for our evaluation and refer to it as "Uniform Allocation".

– *Determining mapping A*. We implemented the algorithm described in section 3.3.2 and use it to designate routers to aggregate virtual prefixes.

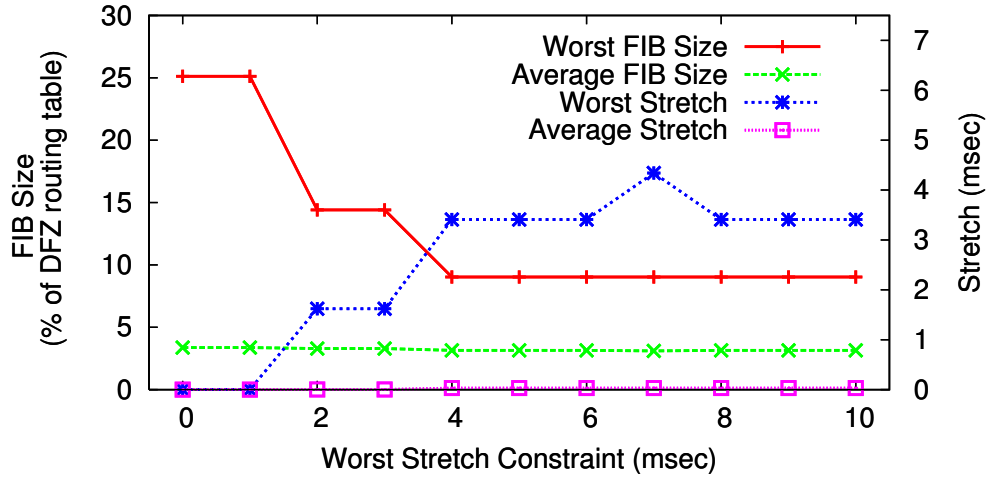
Router FIB

We first look at the size and the composition of the FIB on the ISP's routers with a ViAggre deployment. Specifically, we focus on the router with the largest FIB for a deployment where the worst-case stretch (C) is constrained to 4ms. The first two bars in figure 3.4 show the FIB composition for a /7 and uniform allocation respectively. With a /7 allocation, the router's FIB contains 46,543 entries which represents 18.2% of the routing table today. This includes 22,772 prefixes, 128 virtual prefixes, 23,643 LSP mappings and 0 popular prefixes. As can be seen, in both cases, the LSP mappings for tunnels to the external routers contribute significantly to the FIB. This is because the ISP has a large number of customer routers that it has peerings with.

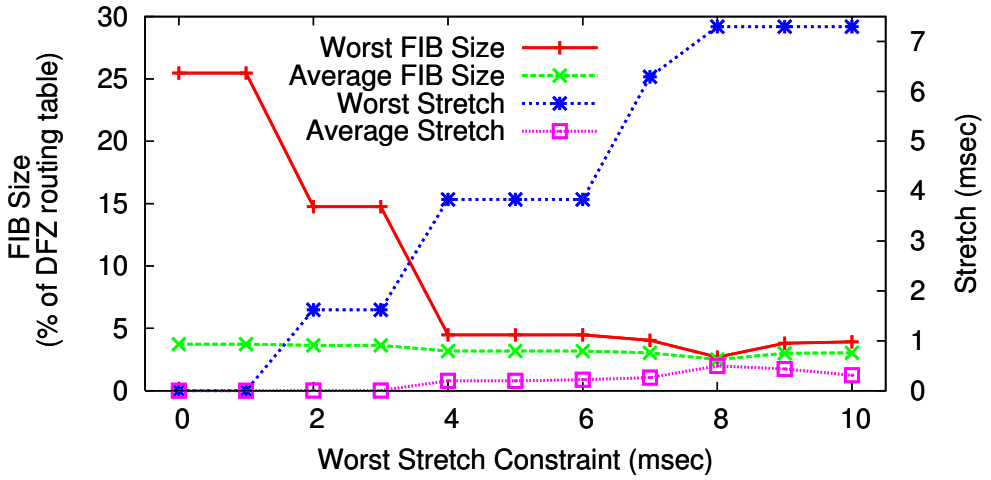
However, we also note that customer ISPs do not advertise the full routing table to their provider. Hence, edge routers of the ISP could maintain routes advertised by customer routers in their FIB, advertise these routes onwards with themselves as the BGP NEXT_HOP and only initiate LSP advertisements for themselves and for peer and provider routers connected to them. With such a scheme, the number of LSP mappings that the ISP's routers need to maintain and the MPLS overhead in general reduces significantly. The latter set of bars in Figure 3.4 shows the FIB composition with such a deployment for the router with the largest FIB. For the /7 allocation, the Worst FIB size is 23,101 entries (9.02% of today's routing table), while for the Uniform allocation, it is 10,226 entries (4.47%). In the rest of this section, we assume this model of deployment.

Stretch Vs. FIB Size

We ran the assignment algorithm with Worst Stretch Constraint (C) ranging from 0 to 10 ms and determined the (Average and Worst) Stretch and FIB Size of the resulting ViAggre deployment. Figure 3.5(a) plots these metrics for the /7 allocation. The Worst FIB size, shown as a fraction of the DFZ routing table size today, expectedly reduces as the constraint on Worst Stretch is relaxed. However, beyond C=4ms, the Worst FIB Size remains constant. This is because the largest virtual prefix with a /7 allocation encompasses 8.9% of the DFZ routing table and the Worst FIB Size cannot be any less than 9.02% (0.12% overhead is due to virtual prefixes and LSP mappings). Figure 3.5(b) plots the same metrics for the Uniform allocation and shows that the FIB can be shrunk even more. The figure also shows that the Average FIB Size and the Average stretch are expectedly small throughout. The anomaly beyond C=8 ms in figure 3.5(b) results



(a) With /7 allocation



(b) With Uniform allocation

Figure 3.5: Variation of FIB Size and Stretch with Worst Stretch constraint and no popular prefixes.

from the fact that our assignment algorithm is an approximation that can yield non-optimal results.

Another way to quantify the benefits of ViAggre is to determine the extension in the life of a router with a specified memory due to the use of ViAggre. As proposed in [69], we used data for the DFZ routing table size from Jan'02 to Dec'07 [68] to fit a quadratic model to routing table growth. Further, it has

Table 3.1: Estimates for router life with ViAggre

		Today	ViAggre			
	Worst Stretch (ms)	–	0	2	4	8
239K FIB	Quad. Fit	Expired	2015	2020	2039	2051
	Expo. Fit	Expired	2018	2022	2031	2035
1M FIB	Quad. Fit	2015	2033	2044	2081	2106
	Expo. Fit	2018	2029	2033	2042	2046

been claimed that the DFZ routing table has seen exponential growth at the rate of 1.3x every two years for the past few years and will continue to do so [94]. We use these models to extrapolate future DFZ routing table size. We consider two router families: Cisco's Cat6500 series with a supervisor 720-3B forwarding engine that can hold up to 239K IPv4 FIB entries and hence, was supposed to be phased out by mid-2007 [36], though some ISPs still continue to use them. We also consider Cisco's current generation of routers with a supervisor 720-3BXL engine that can hold 1M IPv4 FIB entries. For each of these router families, we calculate the year to which they would be able to cope with the growth in the DFZ routing table with the existing setup and with ViAggre. Table 3.1 shows the results for the Uniform Allocation.

For ViAggre, relaxing the worst-case stretch constraints reduces FIB size and hence, extends the router life. The table shows that if the DFZ routing table were to grow at the aforementioned exponential rate, ViAggre can extend the life of the previous generation of routers to 2018 with no stretch at all. We realize that estimates beyond a few years are not very relevant since the ISP would need to upgrade its routers for other reasons such as newer technologies and higher data rates anyway. However, with ViAggre, at least the ISP is not forced to upgrade due to growth in the routing table.

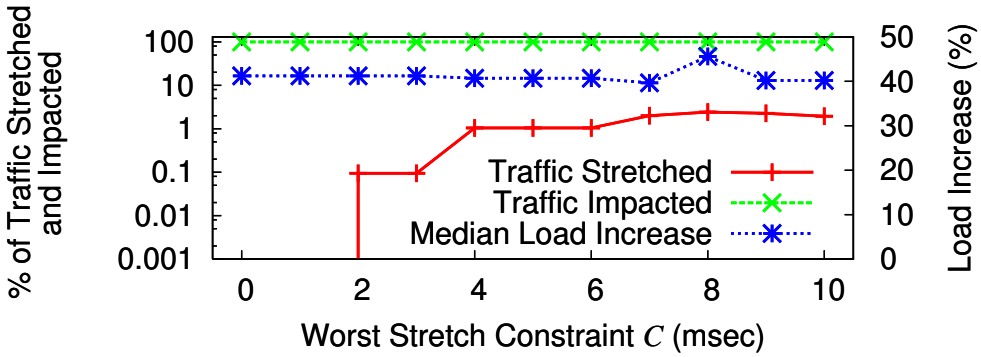


Figure 3.6: Variation of the percentage of traffic stretched/impacted and load increase across routers with Worst Stretch Constraint (Uniform Allocation) and no popular prefixes.

Figure 3.6 plots the impact of ViAggre on the ISP's traffic and router load. The percentage of traffic stretched is small, less than 1% for $C \leq 6$ ms. This shows that almost all the traffic is routed through an aggregation point in the same PoP as the ingress. However, the fact that no prefixes are considered popular implies that almost all the traffic follows a different router-level path as compared to the status quo. This shows up in figure 3.6 since the traffic impacted is ≈100% throughout. This, in turn, results in a median increase in load across the routers by ≈39%. In the next section we discuss how an ISP can use the skewed distribution of traffic to address the load concern while maintaining a small FIB on its routers.

Popular Prefixes

Past studies of ISP traffic patterns from as early as 1999 have observed that a small fraction of Internet prefixes carry a large majority of ISP traffic [42,48,108,127]. We used Netflow records collected across the routers of the same tier-1 ISP as in the last section for a period of two months (20^{th} Nov'07

to 20^{th} Jan'07) to generate per-prefix traffic statistics and observed that this pattern continues to the present day. The line labeled "Day-based, ISP-wide" in figure 3.7 plots the average fraction of the ISP's traffic destined to a given fraction of popular prefixes when the set of popular prefixes is calculated across the ISP on a daily basis. The figure shows that 1.5% of most popular prefixes carry 75.5% of the traffic while 5% of the prefixes carry 90.2% of the traffic.

ViAggre exploits the notion of prefix popularity to reduce its impact on the ISP's traffic. However, the ISP's routers need not consider the same set of prefixes as popular; instead the popular prefixes can be chosen per-PoP or even per-router. We calculated the fraction of traffic carried by popular prefixes, when popularity is calculated separately for each PoP on a daily basis. This is plotted in the figure as "Day-based, per-PoP" and the fractions are even higher. [3]

When using prefix popularity for router configuration, it would be preferable to be able to calculate the popular prefixes over a week, month, or even longer durations. The line labeled "Estimate, per-PoP" in the figure shows the amount of traffic carried to prefixes that are popular on a given day over the period of the next month, averaged over each day in the first month of our study. As can be seen, the estimate based on prefixes popular on any given day carries just a little less traffic as when the prefix popularity is calculated daily. This suggests that prefix popularity is stable enough for ViAggre configuration and the ISP can use the prefixes that are popular on a given day for a month or so. However, we admit that that these results are very preliminary and we need to study ISP traffic patterns over a longer period to substantiate the claims made above.

[3]We did not have Netflow records for individual routers and hence, were unable to generate router-specific popular prefixes.

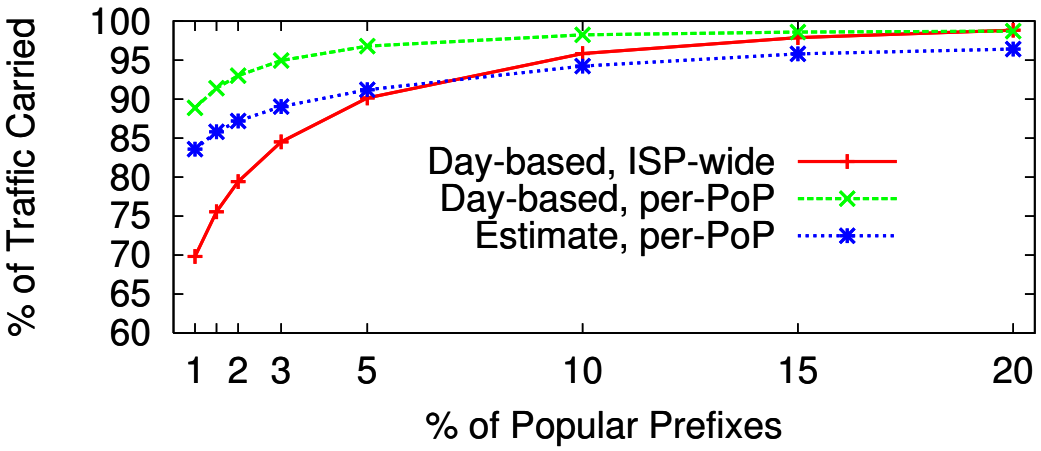
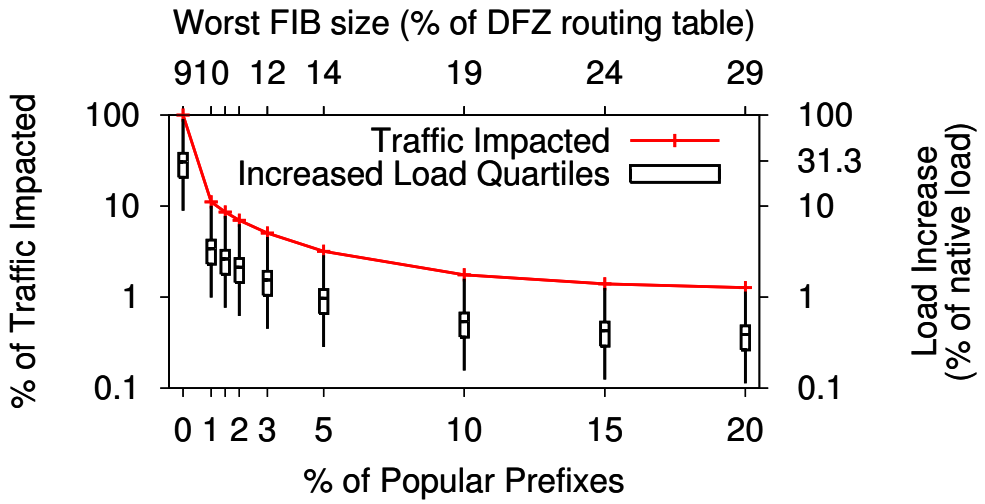


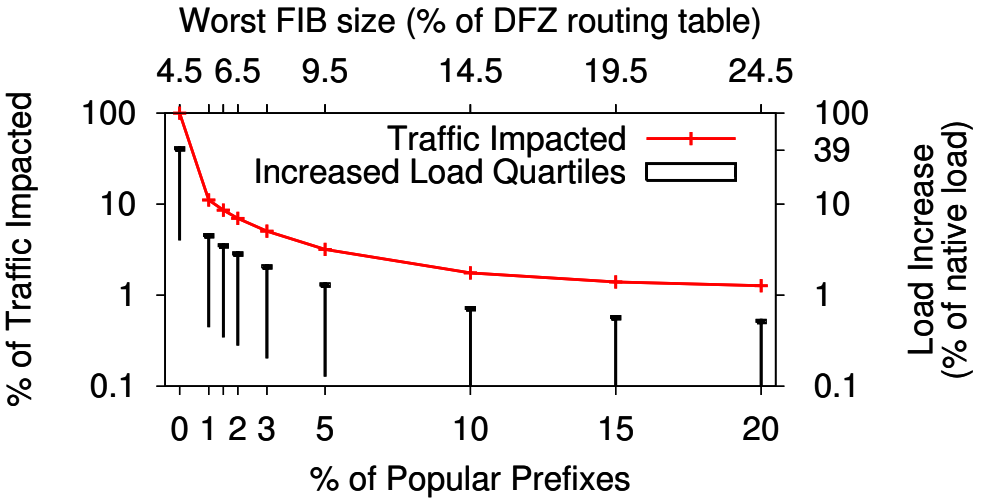
Figure 3.7: Popular prefixes carry a large fraction of the ISP's traffic.

Load Analysis

We now consider the impact of a ViAggre deployment involving popular prefixes, i.e. the ISP populates the FIB on its routers with popular prefixes. Specifically, we focus on a deployment wherein the aggregation points are assigned to constrain Worst Stretch to 4ms, i.e. C = 4ms. Figure 3.8 shows how the traffic impacted and the quartiles for the load increase vary with the percentage of popular prefixes for both allocations. Note that using popular prefixes increases the router FIB size by the number of prefixes considered popular and thus, the upper X-axis in the figure shows the Worst FIB size. The large fraction of traffic carried by popular prefixes implies that both the traffic impacted and the load increase drops sharply even when a small fraction of prefixes is considered popular. For instance, with 2% popular prefixes in case of the uniform allocation (figure 3.8(b)), 7% of the traffic follows a different router-level path than before while the largest load increase is 3.1% of the original router load. With 5% popular prefixes, the largest load increase is 1.38%. Note that the more even



(a) With /7 allocation



(b) With Uniform allocation

Figure 3.8: Variation of Traffic Impacted and Load Increase (0-25-50-75-100 percentile) with percentage of popular prefixes, C=4ms.

distribution of prefixes across virtual prefixes in the uniform allocation results in a more even distribution of the excess traffic load across the ISP's routers – this shows up in the load quartiles being much smaller in figure 3.8(b) as compared to the ones in figure 3.8(a).

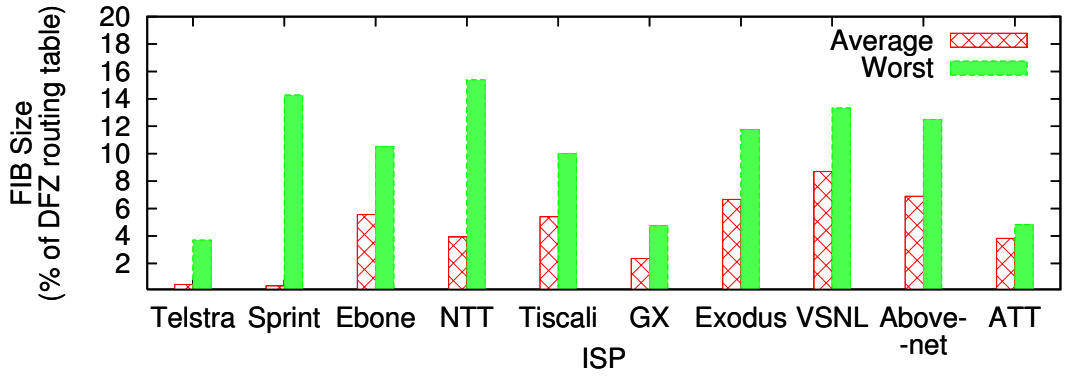


Figure 3.9: FIB size for various ISPs using ViAggre.

3.4.3 Rocketfuel Study

We studied the topologies of 10 ISPs collected as part of the Rocketfuel project [118] to determine the FIB size savings that ViAggre would yield. Note that the fact we don't have traffic matrices for these ISPs implies that we cannot analyze the load increase across their routers. For each ISP, we used the assignment algorithm to determine the worst FIB size resulting from a ViAggre deployment where the worst stretch is limited to 5ms. Figure 3.9 shows that the worst FIB size is always less than 15% of the DFZ routing table. The FIB size is relatively higher for NTT and Sprint because they have a global footprint with a few small PoPs outside their main area of influence. For instance, Sprint has a few small PoPs in the Asia-Pacific region. The constraint on the worst stretch implies that in many cases, the traffic from these PoPs cannot be routed to an aggregation point in another PoP and so these PoPs must have aggregation points for all virtual prefixes. Consequently, the routers in these PoPs end up with a relatively large FIB. However, the Rocketfuel topologies are not complete and are missing routers. Hence, while the results presented here are encouraging, they should be treated as conservative estimates of the savings that ViAggre

would yield for these ISPs.

3.4.4 Discussion

The analysis above shows that ViAggre can significantly reduce FIB size. Most of the ISPs we studied are large tier-1 and tier-2 ISPs. However, smaller tier-2 and tier-3 ISPs are also part of the Internet DFZ. Actually, it is probably more important for such ISPs to be able to operate without needing to upgrade to the latest generation of routers. The fact that these ISPs have small PoPs might suggest that ViAggre would not be very beneficial. However, given their small size, the PoPs of these ISPs are typically geographically close to each other. Hence, it is possible to use the cumulative FIB space across routers of close-by PoPs to shrink the FIB substantially. And the use of popular prefixes ensures that the load increase and the traffic impact is still small. For instance, we analyzed router topology and routing table data from a regional tier-2 ISP (AS2497) and found that a ViAggre deployment with worst stretch less than 5ms can shrink the Worst FIB size to 14.2% of the routing table today.

Further, the fact that such ISPs are not tier-1 ISPs implies they are a customer of at least one other ISP. Hence, in many cases, the ISP could substantially shrink the FIB size on its routers by applying ViAggre to the small number of prefixes advertised by their customers and peers while using default routes for the rest of the prefixes.

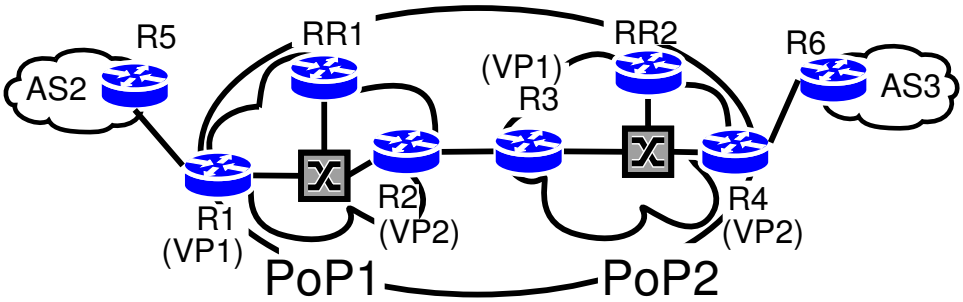


Figure 3.10: WAIL topology used for our deployment. All routers in the figure are Cisco 7300s. RR1 and RR2 are route reflectors and are not on the data path. Routers R1 and R3 aggregate virtual prefix VP1 while routers R2 and R4 aggregate VP2.

3.5 Deployment

To verify the claim that ViAggre is a configuration-only solution, we deployed both ViAggre designs on a small network built on the WAIL testbed [21]. The test network is shown in figure 3.10 and represents an ISP with two PoPs. Each PoP has two Cisco 7301 routers and a route reflector.[4] For the ViAggre deployment, we use two virtual prefixes: 0.0.0.0/1 (VP1) and 128.0.0.0/1 (VP2) with one router in each PoP serving as an aggregation point for each virtual prefix. Routers R1 and R4 have an external router connected to them and exchange routes using an eBGP peering. Specifically, router R5 advertises the entire DFZ routing table and this is, in turn, advertised through the ISP to router R6. We use OSPF for intra-domain routing. Beyond this, we configure the internal distribution of BGP routes according to the following three approaches:

1). **Status Quo.** The routers use a mesh of iBGP peerings to exchange the routes and hence, each router maintains the entire routing table.

[4]These are used only for the design-II deployment. We used both a Cisco 7301 and a Linux PC as a route reflector.

2). **Design-I.** The routers still use a mesh of iBGP peerings to exchange routes. Beyond this, the routers are configured as follows:

– *Virtual Prefixes.* Routers advertise the virtual prefix they are aggregating to their iBGP peers.

– *FIB Suppression.* Each router only loads the routes that it is aggregating into its FIB. For instance, router R1 uses an `access list` to specify that only routes belonging to VP1, the virtual prefix VP2 itself and any popular prefixes are loaded into the FIB. A snippet of this access list is shown below.

Snippet 1: Access List Configuration

```
1: ...
   ! R5's IP address is 198.18.1.200
2: distance 255 198.18.1.200 0.0.0.0 1
   ! Don't mark anything inside 0.0.0.0/1
3: access-list 1 deny 0.0.0.0 128.255.255.255
   ! Don't mark virtual prefix 128.0.0.0/1
4: access-list 1 deny 0.0.0.0 128.0.0.0
   ! Don't mark popular prefix 122.1.1.0/24
5: access-list 1 deny 122.1.1.0 0.0.0.255
   ! ... other popular prefixes follow ...
   ! Mark the rest with admin distance 255
6: access-list 1 permit any
```

Here, the `distance` command on line 2 sets the administrative distance of all prefixes that are accepted by `access list` 1 to "255" and these routes are not loaded by the router into its FIB.

– *LSPs to external routers.* We use MPLS for the tunnels between routers. To this effect, LDP [9] is enabled on the interfaces of all routers and establishes LSPs between the routers. Further, each edge router (R1 and R4) initiates a Downstream Unsolicited tunnel [9] for each external router connected to them to all their IGP neighbors using LDP. This ensures that packets to an external router are forwarded using MPLS to the edge router which strips the MPLS header before forwarding them onwards.

Given this setup and assuming no popular prefixes, routers R1 and R3 store 40.9% of today's routing table (107,943 prefixes that are in VP1) while R2 and R4 store 59.1%.

3). **Design-II.** The routers in a PoP peer with the route reflector of the PoP and the route reflectors peer with each other. External routers R1 and R6 are reconfigured to have eBGP peerings with RR1 and RR2 respectively. The advertisement of virtual prefixes and the MPLS configuration is the same as above. Beyond this, the route reflectors are configured to ensure that they only advertise the prefixes being aggregated by a router to it. For instance, RR1 uses a `prefix list` to ensure that only prefixes belonging to VP1, virtual prefix VP2 itself and popular prefixes are advertised to router R1. The structure of this prefix list is similar to the access list shown above. Finally, route reflectors use a route-map on their eBGP peerings to change the BGP NEXT_HOP of the advertised routes to the edge router that the external peer is connected too. This ensures that the packets don't actually flow through the route reflectors.

3.5.1 Configuration Overhead

A drawback of ViAggre being a "configuration-only" approach is the overhead that the extra configuration entails. The discussion above details the extra configuration that routers need to participate in ViAggre. Based on our deployment, the number of extra configuration lines needed for a router r to be configured according to design-I is given by $(r_{int} + r_{ext} + 2|A(r)| + |P_r| + 6)$ where r_{int} is the number of router interfaces, r_{ext} is the number of external routers r is peering with, $|A(r)|$ is the number of virtual prefixes r is aggregating and $|P_r|$ is the number of popular prefixes in r. Given the size of the routing table today, considering even a small fraction of prefixes as popular would cause the expression to be dominated by $|P_r|$ and can represent a large number of configuration lines.

However, quantifying the extra configuration lines does not paint the complete picture since given a list of popular prefixes, it is trivial to generate an access or prefix list that would allow them. To illustrate this, we developed a configuration tool as part of our deployment effort. The tool is 334 line Python script which takes as input a router's existing configuration file, the list of virtual prefixes, the router's (or representative) Netflow records and the percentage of prefixes to be considered popular. The tool extracts relevant information, such as information about the router's interfaces and peerings, from the configuration file. It also uses the Netflow records to determine the list of prefixes to be considered popular. Based on these extracted details, the script generates a configuration file that allows the router to operate as a ViAggre router. We have been using this tool for experiments with our deployment. Further, we use *clogin* [151] to automatically load the generated ViAggre configuration file onto the router. Thus, we can reconfigure our testbed from status quo operation

to ViAggre operation (design-I and design II) in an automated fashion. While our tool is specific to the router vendor and other technologies in our deployment, its simplicity and our experience with it lends evidence to the argument that ViAggre offers a good trade-off between the configuration overhead and increased routing scalability.

3.5.2 Control Plane Overhead

Section 3.4 evaluated the impact of ViAggre on the ISP's data plane – the impact on the ISP's traffic and the routers forwarding this traffic. Beyond this, ViAggre uses control plane mechanisms to divide the routing table amongst the ISP's routers. Design-I uses `access lists` on routers to suppress routes from being loaded into the FIB and Design-II uses `prefix lists` on route reflectors to prevent the routes from being advertised to their clients, i.e., the routers in the PoP. We quantify the performance overhead imposed by these mechanisms using our deployment. Specifically, we look at the impact of our designs on the propagation of routes through the ISP.

To this effect, we configured the internal distribution of BGP routes in our testbed according to the three approaches described above. External router R5 is configured to advertise a variable number of prefixes through its eBGP peering. We restart this peering on router R5 and measure the time it takes for the routes to be installed into the FIB of the ISP's routers and then advertised onwards; hereon we refer to this as the *installation time*. During this time, we also measure the CPU utilization on the routers. We achieve this by using a clogin script to execute the "*show process cpu*" command on each router every 5 sec-

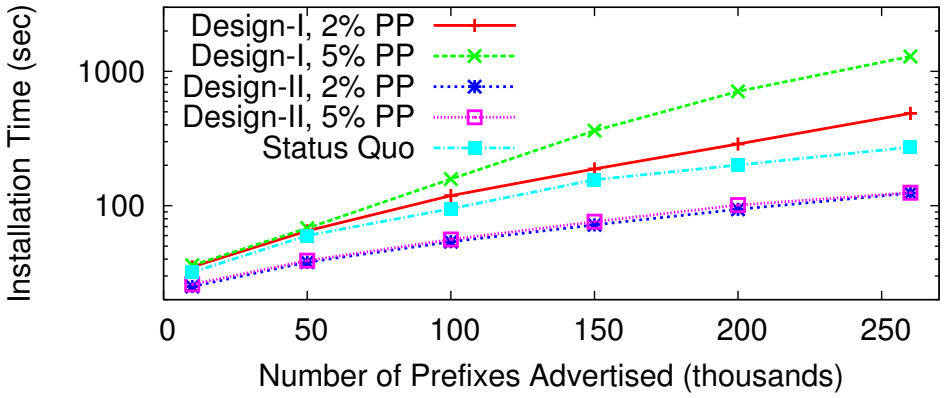
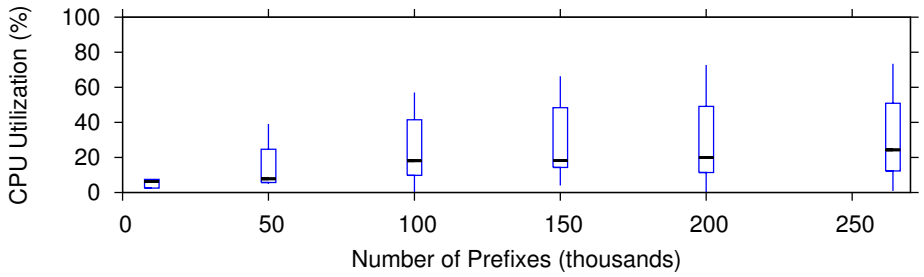


Figure 3.11: Installation time with different approaches and varying fraction of Popular Prefixes (PP).

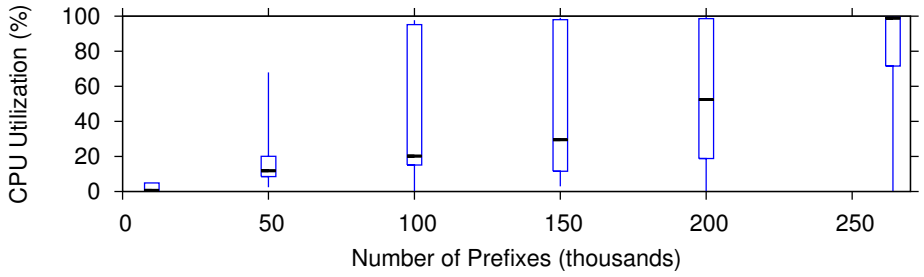
onds. The command gives the average CPU utilization of individual processes on the router over the past 5 seconds and we extract the CPU utilization of the "BGP router" process.

We measured the installation time and the CPU utilization for the three approaches. For status quo and design-I, we focus on the measurements for router R1 while for design-II, we focus on the measurements for route reflector RR1. We also varied the number of popular prefixes. Here we present results with 2% and 5% popular prefixes. Figures 3.11 and 3.12 plot the installation time and the quartiles for the CPU utilization respectively.

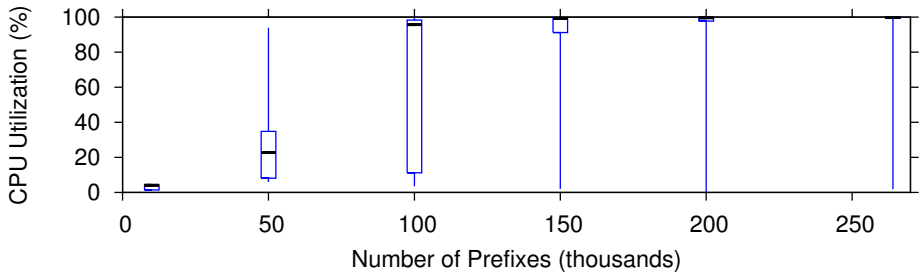
Design-I versus Status Quo. Figure 3.11 shows that the installation time with design-I is much higher than that with status quo. For instance, with status quo, the complete routing table is transferred and installed on router R1 in 273 seconds while with design-I and 2% popular prefixes, it takes 487 seconds. Further, the design-I installation time increases significantly as the number of popular prefixes increases. Finally, figures 3.12(b) and 3.12(c) show that design-I leads to a very high CPU load during the transfer, which increases as more prefixes



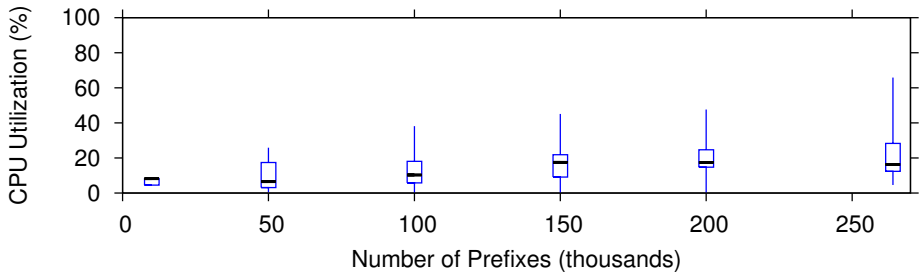
(a) Status Quo, Measured on R1



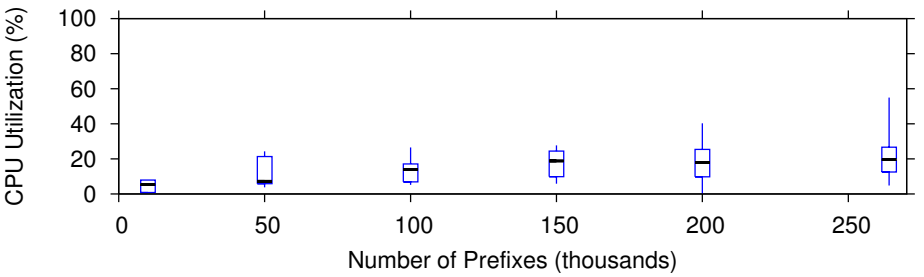
(b) Design-I, 2% PP, Measured on R1



(c) Design-I, 5% PP, Measured on R1



(d) Design-II, 2% PP, Measured on RR1



(e) Design-II, 5% PP, Measured on RR1

Figure 3.12: CPU Utilization quartiles (0-25-50-75-100 percentile) for the three approaches and different fraction of Popular Prefixes (PP).

are considered popular. This results from the fact that access lists with a large number of rules are very inefficient and would obviously be unacceptable for an ISP deploying ViAggre. We are currently exploring ways to achieve FIB suppression without the use of access list.

Design-II versus Status Quo. Figure 3.11 shows that the time to transfer, install and propagate routes with design-II is lesser than status quo. For instance, design-II with 2% popular prefixes leads to an installation time of 124 seconds for the entire routing table as compared to 273 seconds for status quo. Further, the installation time does not change much as the fraction of popular prefixes increases. Figures 3.12(d) and 3.12(e) show that the CPU utilization is low with median utilization being less than 20%. Note that the utilization shown for design-II was measured on route reflector RR1 which has fewer peerings than router R1 in status quo. This explains the fact that the utilization with design-II is less than status quo. While preliminary, this experiment suggests that design-II can also help with route convergence within the ISP.

3.5.3 Failover

As detailed in section 3.2.5, as long as alternate aggregation points exist, traffic in a ViAggre network is automatically rerouted upon failure of the aggregation point being used. We measured this failover time using our testbed. In the interest of space, we very briefly summarize the experiment here. We generated UDP traffic between PCs connected to routers R5 and R6 (figure 3.10) and then crashed the router being used as the aggregation point for the traffic. We measured the time it takes for traffic to be rerouted over 10 runs with each design.

In both cases, the maximum observed failover time was 200 us. This shows that our designs ensure fast failover between aggregation points.

3.6 Discussion

Pros. ViAggre can be *incrementally deployed* by an ISP since it does not require the cooperation of other ISPs and router vendors. The ISP does not need to change the structure of its PoPs or its topology. What's more, an ISP could experiment with ViAggre on a limited scale (a few virtual prefixes or a limited number of PoPs) to gain experience and comfort before expanding its deployment. None of the attributes in the BGP routes advertised by the ISP to its neighbors are changed due to the adoption of ViAggre. Also, the use of ViAggre by the ISP does not restrict its routing policies and route selection. Further, at least for design-II, control plane processing is reduced. Finally, there is *incentive for deployment* since the ISP improves its own capability to deal with routing table growth.

Management Overhead. As detailed in section 3.5.1, ViAggre requires extra configuration on the ISP's routers. Beyond this, the ISP needs to make a number of deployment decisions such as choosing the virtual prefixes to use, deciding where to keep aggregation points for each virtual prefix, and so on. Apart from such one-time or infrequent decisions, ViAggre may also influence very important aspects of the ISP's day-to-day operation such as maintenance, debugging, etc. All this leads to increased complexity and there is a cost associated with the extra management.

In section 3.5.1 we discussed a configuration tool that automates ViAggre

configuration. We are also implementing a planning tool that takes as input high-level constraints specified by the human ISP manager such as constraints on the traffic stretch, router load, router memory used and the robustness of the resulting design. It then uses ILP to solve a multiple-constraint optimization problem to generate VA-specific deployment details such as the assignment of aggregation points. These two tools combined would provide human ISP managers an automated means to adopt ViAggre without needing to delve into ViAggre and configuration-specific details.

It is difficult to speculate about actual costs, so we do not compare the increase in management costs against the cost of upgrading routers, which apart from capital costs requires reconfiguring every customer on every router twice. While we hope that our tools will actually lead to cost savings for a ViAggre network, an ISP might be inclined to adopt ViAggre just because it breaks the dependency of various aspects of its operation on the size of the routing table. These aspects include its upgrade cycle, the per-byte forwarding cost, and the per-byte forwarding power.

Tunneling Overhead. An important concern arising out of the use of ViAggre is the tunneling overhead. However, the extensive use of tunnels (MPLS, GRE-IP, IPSec, VLAN tunneling) in ISP networks has meant that most routers are already equipped with interfaces that have extensive tunneling and detunneling capabilities at line rates [50].

Popular Prefixes. As mentioned earlier, ViAggre represents a trade-off between FIB shrinkage on one hand and increased router load and traffic stretch on the other. The fact that Internet traffic follows a power-law distribution makes this a very beneficial trade-off. This power-law observation has held up in measure-

ment studies from 1999 [42] to 2008 (in this thesis) and hence, Internet traffic has followed this distribution for at least the past nine years in spite of the rise in popularity of P2P and video streaming. We believe that, more likely than not, future Internet traffic will be power-law distributed and hence, ViAggre will represent a good trade-off for ISPs.

Other design points. The ViAggre proposal presented here represents one point in the design space that we focused on for the sake of concreteness. The basic idea of dividing the routing table such that individual routers only need to maintain part of the routes can be achieved using a few alternative approaches.

– *Adding routers.* We have presented a couple of techniques that ensure that only a subset of the routing table is loaded into the FIB. Given this, an ISP could install "slow-fat routers", low-end devices (or maybe even a stack of software routers [55]) in each PoP that are only responsible for routing traffic destined to unpopular prefixes. These devices forward a low volume of traffic, so it would be easier and cheaper to hold the entire routing table. The popular prefixes are loaded into existing routers. This approach can be seen as a variant of route caching and does away with a lot of deployment complexity. In fact, ViAggre may allow us to revisit route caching [79].

– *Router changes.* Routers can be changed to be ViAggre-aware and hence, make virtual prefixes first-class network objects. This would do away with a lot of the configuration complexity that ViAggre entails, ensure that ISPs get vendor support and hence, make it more palatable for ISPs. We, in cooperation with a router vendor, are exploring this option [52].

Routers today tend to have multiple blades with each blade maintaining its

own copy of the entire routing table. Another approach involving vendor support is to split the routing table amongst router blades using ViAggre and hence, achieve FIB shrinkage with less burden on the ISP itself.

– *Clean-slate ViAggre.* The basic concept of virtual networks can be applied in an inter-domain fashion. The idea here is to use cooperation amongst ISPs to induce a routing hierarchy that is more aggregatable and hence, can accrue benefits beyond shrinking the router FIB. This involves virtual networks for individual virtual prefixes spanning domains such that even the RIB on a router only contains the prefixes it is responsible for. This would reduce both the router FIB and RIB and in general, improve routing scalability.

3.7 Related Work

A number of efforts have tried to directly tackle the routing scalability problem through clean-slate designs. One set of approaches try to reduce routing table size by dividing edge networks and ISPs into separate address spaces [38,43,92,100,138]. Our work resembles some aspects of CRIO [138], which uses virtual prefixes and tunneling to decouple network topology from addressing. However, CRIO requires adoption by all provider networks and like other proposals [38,43,92,100], requires a new mapping service to determine tunnel endpoints. APT [70] presents such a mapping service. Similar to CRIO, Verkaik et. al. [129] group prefixes with similar behavior into policy atoms and use these atoms and tunneling of packets to reduce routing table size. Alternatively, it is possible to encode location information into IP addresses [37,51,61] and hence, reduce routing table size. Finally, an interesting set of approaches

that trade-off stretch for routing table size are *Compact Routing* algorithms; the key idea behind such algorithms is the adoption of a more flexible notion of best path. Krioukov et. al. [83] analyze the performance of such an algorithm for Internet-like graphs; see [84] for a survey of the area.

The use of tunnels has long been proposed as a routing scaling mechanism. VPN technologies such as BGP-MPLS VPNs [40] use tunnels to ensure that only PE routers directly connected to the VPN'ed customers need to keep the VPN routes. As a matter of fact, ISPs can and probably do use tunneling protocols such as MPLS and RSVP-TE to engineer a BGP-free core [113]. However, edge routers still need to keep the full FIB. With ViAggre, none of the routers on the data-path need to maintain the full FIB. Forgetful routing [72] selectively discards alternative routes to reduce RIB size.

Router vendors, if willing, can use a number of techniques to reduce the FIB size. This includes FIB compression [113], in which routers don't install redundant more-specific prefixes in the FIB. Another technique is route caching but it seems to have lost favor with both router vendors and ISPs [113].

Another major problem for ISPs with regards to routing scalability is VPN routing tables. While these tables are only kept by the Provider-Equipment (PE) routers directly connected to customers need to keep them, they can be several times larger. In recent work, Kim et. al. [80] use relaying, similar to ViAggre's use of aggregation points, to address the VPN routing scalability problem. The VPN setting involves VPN-specific routing tables and the task of maintaining these can be split amongst PE routers; in our setting there is just the Internet routing table and we use the concept of virtual prefixes to make it divisible. We also have the additional challenge of dealing with networks other than cus-

tomers since these networks might be advertising the full routing table, which is solved by not installing some routes in the FIB (design-I) or through the use filters on route reflectors (design-II).

Over the years, several articles have documented the existing state of inter-domain routing and delineated requirements for the future [35,45,89]; see [45] for other routing related proposals. While scalability might be the most important problem afflicting inter-domain routing, several aspects of routing, such as improving BGP convergence time [124,126], enabling host control over routing [135], and improving routing security [77,122] have also received a lot of attention. Mao et al. [89] discuss other such challenges facing the routing system.

RCP [25] and 4D [56] argue for logical centralization of routing in ISPs to provide scalable internal route distribution and a simplified control plane respectively. We note that ViAggre fits well into these alternative routing models since the centralized control plane can be used to ensure that routers only obtain the relevant part of the routing table while the data-plane design is the same as proposed here. As a matter of fact, the use of route reflectors in design-II is similar in spirit to RCSs in [25] and DEs in [56].

3.8 Summary

In this chapter we presented ViAggre, a technique that can be used by an ISP to substantially shrink the FIB on its routers and hence, extend the lifetime of its installed router base. The ISP may have to upgrade the routers for other reasons but at least it is not driven by DFZ growth over which it has no control.

While it remains to be seen whether the use of automated tools to configure and manage large ViAggre deployments can offset the complexity concerns, we believe that the simplicity of the proposal and its possible short-term impact on routing scalability suggest that is an alternative worth considering.

CHAPTER 4

IP ANYCAST MEASUREMENT STUDY

4.1 Overview

IP Anycast [102] is an IP addressing and delivery mode whereby an IP packet is sent to one of a group of hosts identified by the IP Anycast address, typically the closest one. While IP unicast is one-to-one, and IP Multicast is one-to-many, IP Anycast is one-to-any.

IP Anycast offers an attractive primitive for service discovery – the route-to-closest-server abstraction offers reduced access latency for clients, load balancing across servers, and network-level resilience to DDoS attacks while its implementation at the network layer allows these advantages to be realized with no special configuration at clients or servers and with no dependence on higher-layer services such as the DNS. While this potential has long been recognized [22,93,102], it is mostly in recent years that IP Anycast has gained in importance. This is in large part due to its use in the critical DNS root server deployment – six of the thirteen DNS root servers have been transparently replicated using IP Anycast and this deployment continues to grow [63]. In addition IP Anycast is being used in a variety of other infrastructure services. For example, IP Anycast is used to improve the performance of caching DNS servers [95], for drawing in private address space DNS queries as part of the AS-112 project [140], to discover rendezvous points for multicast groups [81], as a transition mechanism from IPv4 to IPv6 [67], for sinkholing DoS attacks [57] and for redirection in commercial CDNs [141]. Finally, recent research efforts include renewed interest in the use of IP Anycast for CDNs [7] and a proposal

for its use as a vehicle for next-generation architecture deployment [12,105].

However, despite its growing use in critical infrastructure services, IP Anycast and its interaction with IP routing practices is not well understood. A couple of root server operators [20,33] have offered valuable reports on the performance of their anycast deployments. These are probably the first reports on the performance of IP Anycast and represent the best source of data on operational IP Anycast deployments from the point of view of the anycast servers. However, the analysis is preliminary (as the authors themselves state) and details of the study are not published, nor is the data publicly available. Drawing from these, in this chapter we present a detailed study of inter-domain IP Anycast as measured from a large number of vantage points. Specifically, our study seeks to answer the following questions:

1. What kind *failover* properties does a typical IP Anycast deployment offer?
2. What kind of *load distribution* do existing IP Anycast deployments see? Also, can the operator of an anycast deployment control this distribution of client load?
3. Past studies [112] have reported that existing IP Anycast deployments may offer poor latency-based *proximity*; *i.e.*, many clients may not be routed to the server closest in terms of latency. Using a larger number of clients and anycast groups, we aim to confirm and understand the reasons for this poor latency as well as explore possible remedial measures.
4. Past studies [20,24,33,111] have presented conflicting reports regarding the *affinity* offered by IP Anycast and consequently, the ability to run stateful

services of top of anycast.[1] We seek to measure, at scale, the affinity offered by IP Anycast.

To explore these questions, we study four existing IP Anycast deployments including two anycasted DNS root servers. In terms of methodology, our study differs from previous efforts on two fronts:

1. We use a variant of the King [59] measurement technique to observe and evaluate IP Anycast deployments from a very large number (>20,000) of vantage points. To the best of our knowledge, this represents a two order of magnitude increase in the number of vantage points from which IP Anycast deployments have been actively probed for evaluation.
2. We deploy our own small scale IP Anycast service for controlled evaluation of anycast under different deployment and failure scenarios. Performing such experiments would be difficult using commercial IP Anycast deployments such as the DNS root servers.

The main results of this study are as follows:

- We corroborate evidence from past studies indicating that IP Anycast, by itself, does not offer good latency-based *proximity*. For example, for the 13 server J-root deployment, we find that 8903 (≈40%) of the 22,281 measured clients are directed to a root server that is more than 100 ms farther away than the closest server. While the impact of inter-domain routing on end-to-end path length has been well documented [117], we find that inter-

[1] Affinity measures the extent to which consecutive anycast packets from a client are delivered to the same anycast server.

domain routing metrics have an even more severe impact on the selection of paths to anycast destinations.

- We propose and evaluate a practical deployment scheme designed to alleviate the *proximity* concerns surrounding IP Anycast. Specifically, ensuring that an ISP that provides transit to an anycast server has global presence and is (geographically) well covered by such servers improves the latency-based proximity offered by the anycast deployment.
- We find that IP Anycast is affected by delayed routing convergence and hence, clients using anycast services may experience slow *failover*. However, our proposed deployment scheme addresses this by reducing the scope of routing convergence that follows a server failure and hence, can ensure that clients failover at a fast rate. For instance, we find that in case of a server failure in an IP Anycast deployment conforming to our proposal, a vast majority (>95%) of clients can be rerouted to other operational servers in less than 20 seconds.
- Ours is a much larger scale study than past efforts. Through this, we find that the anycasting of an IP prefix does not have any unfavorable interactions with inter-domain routing. Hence, IP Anycast offers very good *affinity* for all but a very small fraction of clients. Using temporal clustering, we show that the poor affinity observed by this small fraction of clients can be attributed to dynamic load balancing mechanisms near them.
- We find that a naive IP Anycast deployment does not lead to an even distribution of *client load* across servers. However, we also propose and evaluate the impact of operators manipulating BGP advertisements at individual anycast servers to control their load. Our results show that such mechanisms can achieve coarse-grained load balancing across anycast servers.

Overall, our measurements show that an IP Anycast service can be deployed so as to provide a robust substrate offering good proximity and fast failover while allowing for coarse-grained control over server load. In the rest of this chapter, we detail our measurement methodology and results. In the next chapter, we use our insights into how to deploy a native IP Anycast service to guide a new Anycast architecture that addresses the limitations of native IP Anycast.

4.2 Related Measurement Studies

An invaluable vantage point for measuring anycast deployments is at the anycast servers themselves. In recent presentations [20,33], the operators of the J and K root servers report on their analysis of client logs collected at their anycast servers. They present the observed distribution of client load and affinity. In terms of load distribution, both studies report a skewed distribution of client load across their respective deployments. With regard to affinity, the J-root operators report instances of clients that exhibit poor affinity and conjecture that anycast may not be suitable for stateful services. By contrast, the K-root operators find that most of their clients experience very high affinity.

Our study builds on these earlier reports. Using *active* measurements from over 20,000 clients, we measure the affinity and load distribution for our own small-scale anycast deployment, explore the reasons behind the observed load and affinity, and evaluate techniques to control server load. In addition to load and affinity, we use active measurements to evaluate the (latency) proximity seen at clients to four different anycast deployments. Finally, using our own deployment, we study the behavior of IP Anycast under server failure. As we

describe in Section 4.6, our desire to use a large number of client vantage points prevents us from measuring the affinity and load to the DNS root server deployments. However, for completeness, we did perform such measurements from a smaller set of clients that we have direct access to (i.e., PlanetLab nodes [31] and approximately 200 publicly-available traceroute servers). Since the results were consistent with the larger-scale measurements over our own deployment, we only present the latter here. The details of the PlanetLab-based study can be found in [15].

We are aware of two recent efforts that measure the performance of IP Anycast using active probing from clients. Sarat et al. [111,112] use PlanetLab nodes as vantage points for evaluating the K-root, F-root and .ORG TLD deployments. They measure proximity and affinity and report poor proximity and moderate-to-poor affinity. Similarly, using PlanetLab and approximately 200 volunteer nodes, Boothe et al. [24] measure the affinity offered by the anycasted DNS root servers and report poor affinity.

4.3 Deployments Measured

An IP Anycast group is associated with an IP address (referred to as the anycast address for the group) and servers join the group by just advertising this address into the routing infrastructure. For an *intra-domain anycast group* with servers restricted to a single administrative domain, this advertisement is into the intra-domain routing protocol for the domain in question. For *inter-domain anycast groups*, each server advertises the anycast address into BGP. [2] Despite the

[2] In practice, each server must advertise a prefix for a block of addresses into BGP. This is the *anycast prefix* for the group and servers are accessible through all addresses in the prefix.

Table 4.1: The three external IP Anycast deployments that we evaluate.

Name	Anycast prefix	AS#	No. of servers
F root server [145]	192.5.5.0/24	3557	27
J root server [148]	192.58.128.0/24	26415	13
AS 112 [140]	192.175.48.0/24	112	20

Table 4.2: The internal IP Anycast deployment comprising of five servers. Each of these advertise the anycast prefix (204.9.168.0/22) through a BGP peering with their host-site onto the upstream provider. Note that IR stands for "Intel-Research".

Server Unicast address	Host-Site	Host-site AS#	Upstream provider
128.84.154.99	Cornell University	26	WCG
12.155.161.153	IR Berkeley	2386	ATT
195.212.206.142	IR Cambridge	65476	ATT-World
12.108.127.148	IR Pittsburgh	2386	ATT
12.17.136.150	IR Seattle	2386	ATT

simplicity of the basic idea, the interaction with BGP, the involvement of multiple administrative domains etc. raise several interesting questions regarding the behavior of inter-domain IP Anycast. We restrict ourselves to studying issues related to inter-domain IP Anycast. Also, while we focus on IPv4 Anycast, our results should apply equally to IPv6 Anycast deployments.

Since clients can access an anycast group simply by sending packets to the IP address associated with the group, IP Anycast has been used for transparent replication of many services including the DNS root servers. For example, the F root server deployment is currently comprised of 37 servers that form an IP Anycast group and each server advertises the F root server anycast prefix. The deployment is also associated with its own AS number (AS#) which serves as the origin AS for the anycast prefix. Thus, clients can access a F root server by sending packets to an address in the prefix. Here we evaluate three such

currently operational deployments - two anycasted DNS root servers and the AS112 anycast deployment.[3] The anycasted AS112 servers are used to draw in reverse DNS queries to and for the link local address space (RFC1918 addresses– 10.0.0.0/8, 172.16.0.0/12 and 192.168.0.0/16). Since we have no control over these deployments, we refer to these as *external deployments*. Table 4.1 gives details for the external deployments including the number of the servers in each deployment at the time of our experiments.[4]

In practice, each "server" in these deployments is a cluster of hosts located behind some form of load balancing device. For example, for the F root server, hosts in each cluster form an intra-domain anycast group and it is the server site's gateway routers that balance incoming traffic between them [2]. However, our focus on inter-domain anycast makes our measurements oblivious to this cluster-based deployment. Thus, for the purpose of our study, each server site can be thought of as a single host.

The production-mode nature of these external deployments makes it difficult, if not impossible, to conduct controlled experiments such as injecting server failure, or manipulating the advertisements at individual servers. For such experiments we deployed our own IP Anycast service that we call the *internal deployment*. For this, we obtained a /22 prefix (204.9.168.0/22) and an AS# (33207) from ARIN and deployed anycast servers at the five sites listed in table 4.2. Each server advertises this prefix and AS# into BGP through their host site and on to their upstream provider and hence they form an IP Anycast group.

[3]The choice of the anycast deployment to measure was limited by the need to know the *unicast* addresses of the individual anycast servers for our experiments.

[4]All three deployments have since grown substantially.

Note that an anycast server's "host-site" refers to the server's immediately upstream AS while the "upstream provider" refers to the closest major ISP (tier-1 or tier-2) that provides transit for the server's traffic. For example, the internal deployment anycast server at Cornell has Cornell (AS# 26) as its host-site and Williams Communication (WCG – AS#7911) as its upstream provider. This difference between the host-site and upstream provider is a quirk of the internal deployment; for most commercial IP Anycast deployments, the host-site is also the upstream provider.

The internal deployment is as yet very small in scale. The biggest hurdle in growing the deployment has been the site-by-site negotiation with upstream providers (ATT, WCG) to clear the advertisement of our anycast prefix. Note that we do not require these upstream providers to actively inject our prefix in the BGP but only propagate the advertisement from the anycast servers onwards. Instead, the approval from the ISPs is only due to the access control ISPs often enforce on the AS numbers and network prefixes they expect to see advertised from their customers. To accelerate this site-by-site deployment, we are currently in the process of deploying anycast servers over NLR [146]. Further details about the internal deployment as well as the requirements for much welcome volunteer sites are available at [10]. While the small size of the internal deployment certainly raises questions regarding the generality of our results, the fact that (in retrospect) most of our results follow intuitive reasoning supports the applicability of our study.

Table 4.3: Geographic distribution of the clients used in our study.

Region	No. of clients	% of Total
North America	12931	54.827
Central America	317	1.344
South America	461	1.954
Europe	5585	23.680
Asia	2402	10.184
S.E. Asia	566	2.400
Oceania	1196	5.071
Africa	187	0.792
Arctic Region	9	0.038
Unknown	204	0.864
Total	23858	100.000

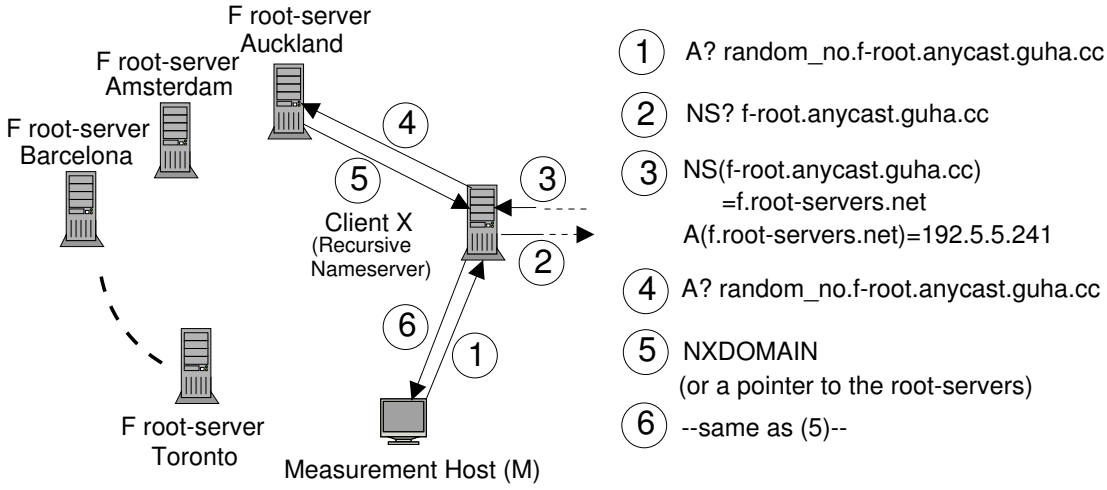


Figure 4.1: Measurement Host M uses DNS queries to direct client X to send packets to the anycast address associated with the F root server deployment. In the scenario depicted here, packets from X are routed to the anycasted F root server located at Auckland, Australia. Note that we use a domain under our control (`anycast.guha.cc`) to trick client X into assuming that 192.5.5.241 is the authoritative nameserver for the domain `f-root.anycast.guha.cc`.

4.4 Methodology

A key observation guiding our measurement methodology is that all three external deployments are DNS services. Hence, the anycast servers that are part of these deployments are all DNS nameservers, and we can probe them using DNS queries. However, since packets from a client to the anycast address of any given anycast deployment are delivered to one of the servers, a large number of clients are needed to ensure that we are able to reach all the anycast servers. To achieve this we used the King measurement technique [59]. King allows for measurement of latency between any two arbitrary hosts on the Internet by utilizing recursive nameservers near the hosts in question.

Using the same basic idea, we used recursive DNS nameservers in the Internet as *clients* in our study. To this effect, we took an address in each routable prefix in a BGP routing table obtained from Route-Views [149] and determined if any of the nameservers authoritative for the `in-addr.arpa` name associated with the address had recursion enabled. For example, for the prefix 128.84.223.0/24, we determined if any of the nameservers authoritative for the name `1.223.84.128.in-addr.arpa.` had recursion enabled. This approach yielded a list of 23,858 unique recursive nameservers. Table 4.3 details the geographic distribution of these nameservers. The nameservers belong to 7,566 different ASs. Hence, they provide us a view of the anycast deployments from 7,566 of the 18,391 routeable ASs on the Internet (based on a BGP routing table obtained from Route-Views). The quantity and the spread of these nameservers makes us confident that our measurements closely reflect the behavior of IP Anycast as seen from hosts in the Internet in general.

Note that using a large number of clients yields a more representative picture for all the metrics we measure here. For example, the scale of our study makes our arguments regarding client load distribution across the anycast deployment significantly more representative than (say) just using PlanetLab for measurements. Moreover, our use of widely-dispersed Internet nameservers as clients avoids the bias that would be introduced were we to use PlanetLab hosts as clients. This bias could severely impact the affinity, proximity and load measurements. For example, one of the J-root servers is connected to the GEANT network and hence, can be accessed through Internet2. Consequently, a large fraction of PlanetLab nodes are routed to this server when they probe the J-root anycast address.

Evaluating an IP Anycast deployment from the perspective of any client (say X) in the list requires that we be able to direct X to send packets to the anycast address of the deployment. As suggested by [59], we leveraged the fact that X is a recursive nameserver and hence, is willing to resolve DNS queries on behalf of other hosts. We can thus "trick" client X into sending DNS queries to an anycast address by making the NS record for a domain point to the address in question and querying X for any record in that domain. We used `anycast.guha.cc`, a domain we own, for this purpose. For example, in case of the F root server deployment with anycast address 192.5.5.241, we created a domain `f-root.anycast.guha.cc` with its NS record pointing to 192.5.5.241. As illustrated in figure 4.1, querying client X for any record in this domain (for example, `random_no.f-root.anycast.guha.cc`) causes X to resolve the NS record for `f-root.anycast.guha.cc` (packets (2) and (3)), and then send a query to the anycast address for the F root server deployment (packet (4)). A minor practical problem with this approach is that a client may be

configured to resend query (4) a number of times on seeing that the response (5) is an error. As suggested by [59], we weeded out clients that may resend queries on receiving an error as follows: we directed each client to the anycast address of the internal anycast deployment and logged the DNS queries at each of the servers of the internal deployment to determine the number of times query (4) is sent.

The experiments presented here use this basic technique for various tasks such as determining the particular anycast server accessed by each client and the latency of doing so. For example, to determine the latency from a client X to the anycast address of the F root server, we:

- Send a recursive DNS query to client X for the NS record for `f-root.anycast.guha.cc`: this primes client X's cache with the fact that the F root server anycast address corresponds to the authoritative nameserver for the domain `f-root.anycast.guha.cc`
- Send an iterative DNS query to client X: since an iterative query is answered by client X based on local information, this provides us with an estimate of the latency for packets {(1),(6)} in figure 4.1.
- Send a recursive DNS query to client X for the A record for `random_no.f-root.anycast.guha.cc`: as shown in figure 4.1, this causes client X to send packets to the F root server anycast address and provides us with an estimate of the latency for packets {(1), (4), (5), (6)}.

This process is repeated eight times and the difference between the minimum measured latency for packets {(1), (4), (5), (6)} and {(1), (6)} is used as an estimate of the round-trip anycast latency from client X to the F root server.

4.5 Proximity

The value of anycast as a server selection primitive is in part due to its ability to find close servers. With IP Anycast, packets destined to an anycast address are routed to the server closest to the client in terms of the metrics used by the underlying routing protocol. For inter-domain IP Anycast, it is the BGP decision process (including routing policies) at the various ASs that governs the anycast server accessed by each client. This implies that anycast packets from clients may not be delivered to servers that are close in terms of *latency* and recent studies [112] have, in fact, indicated that IP Anycast offers poor latency-based *proximity*. In this section we use latency measurements from our $\approx$ 20,000 clients to show that this is indeed the case for existing anycast deployments. However, we also argue that poor latency can be avoided through a planned deployment.

Methodology: To determine the quality of proximity offered by an IP Anycast deployment to a given client, we need to determine the following latencies:

- *Unicast Latency* to all anycast servers - here, unicast latency to an anycast server is the latency from the client to the unicast address of the server. Given that each client is a recursive nameserver, the King approach for determining latencies between two hosts applies here directly.
- *Anycast Latency* for the client or the latency from the client to the anycast address of the deployment. The procedure for determining the anycast latency for a client was described in the previous section.

We define *stretch-factor* as the difference between the anycast latency and the minimum unicast latency for a client. The stretch factor thus represents

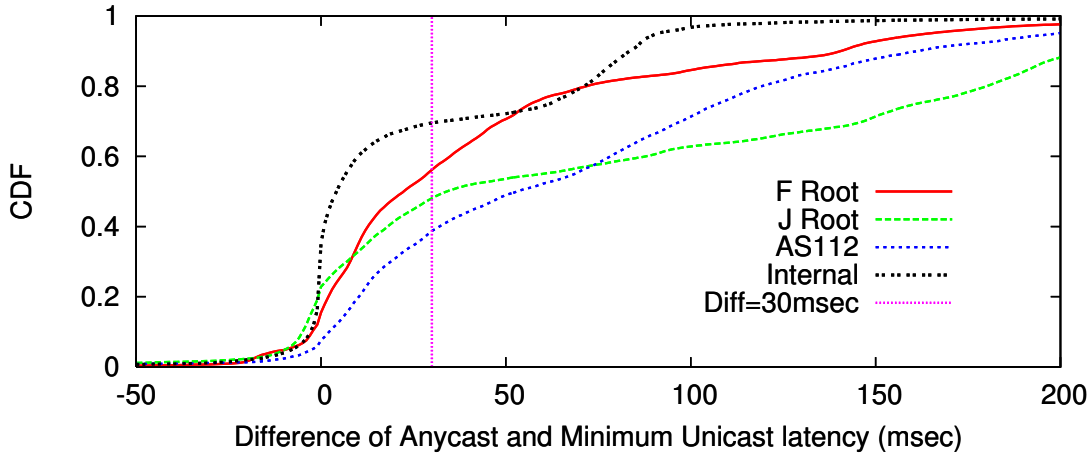


Figure 4.2: CDF for the difference between the anycast and minimum unicast latency for the external and the internal deployments.

the quality of latency-based proximity an anycast deployment offers a client. We determined the stretch factor for each client in our list for the external and internal anycast deployments.

Results: Figure 4.2 shows the CDF for the stretch factor for all the clients. The stretch factor can be negative because anycast packets may be routed to the closest server, yet take a path shorter than the unicast path to the closest server. We see that for all four anycast deployments, a fair fraction of clients are not routed to the server closest to them.[5] For example, the number of clients that are routed to a server that is more than 30 ms farther away from the closest server ranges from 31% (for the internal deployment) to 61% (for the AS112 deployment). Similarly, in case of the J root server deployment, 40% of the clients incur a stretch-factor of more than 100 ms.

[5]The poor selection in case of F root server deployment can be attributed to their use of hierarchical anycast [1,145]. In effect, only 2 of the 27 F root servers advertise the F-root anycast prefix globally. So it is not fair to compare the anycast latency to the minimum unicast latency across all the 27 servers. However, this is not the case for the other deployments we measured.

IP Anycast deployments, by their very nature, are vulnerable to the fact that Internet routing is not based on latency. This is because the anycast servers are geographically distributed and hence, there is greater margin for error in selecting which server to route to. This, in turn, results in the inefficacy of IP Anycast in selecting closeby anycast servers. The issue is exacerbated by the fact that anycast servers tend to have different upstream providers, which is true for both the internal and the external anycast deployments. For example, in case of the internal deployment, the anycast servers have three different upstream providers.

The following anecdotal scenario serves to illustrate the causes of poor proximity. We use the example of a publicly available traceroute-server at UC-Berkeley (`net.berkeley.edu`) acting as a client trying to access the internal anycast deployment. Anycast packets from this client are routed to the server at Cornell (latency=87msec) instead of the close-by server at Berkeley (latency=9msec). Figure 4.3 shows the relevant AS-level connectivity with the relevant POPs of ATT (upstream provider for the server at Berkeley) and WCG (upstream provider for the server at Cornell). We used BGP looking-glass servers to determine that Level3 has at least two paths for the internal deployment's anycast prefix: one through WCG's Santa Clara POP with AS-PATH=[7911, 26, 33207] and the other through ATT's San Francisco POP with AS-PATH=[7018, 2386, 33207]. The Level3 routers in the area choose the first path as the best path and hence, anycast packets from the client are delivered to the anycast server at Cornell. This path is labeled as "Actual Path" in the figure.

This example points to the crux of why Internet routing yields poor proximity for the measured anycast deployments. From the point of view of inter-

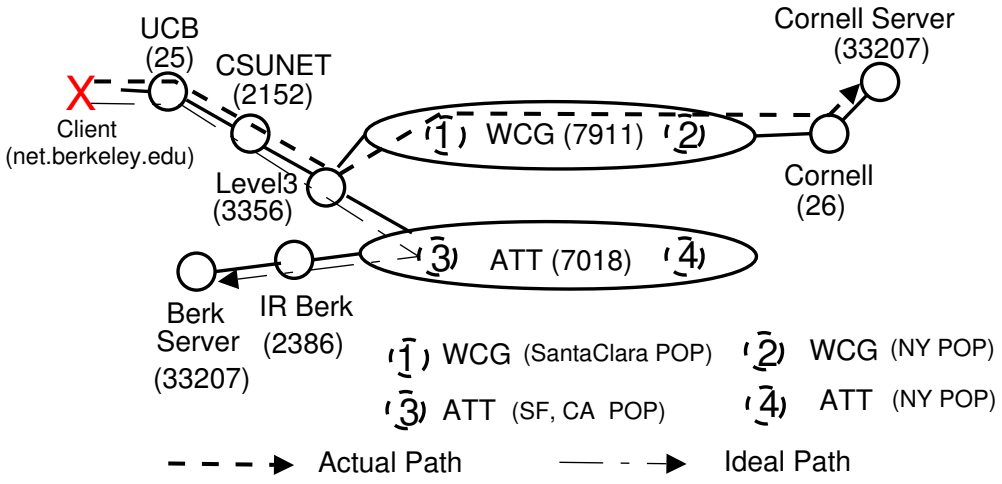


Figure 4.3: The relevant AS-level connectivity (with corresponding AS numbers in parenthesis) between the client (`net.berkeley.edu`) in UC-Berkeley and the internal anycast deployment servers at Cornell and Berkeley. Level3 receives at least two different advertisements for the internal deployment anycast prefix with the following AS-PATHs: [7911, 26, 33207] and [7018, 2386, 33207].

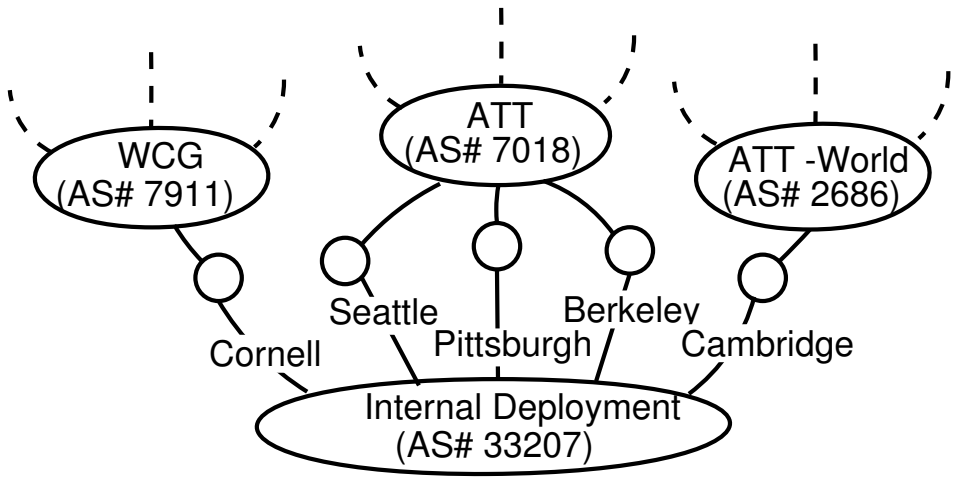


Figure 4.4: AS-level connectivity for the anycasted internal deployment – from the point of inter-domain routing, AS# 33207 is a multi-homed stub AS.

domain routing, an anycasted AS is equivalent to a multi-homed stub AS (see figure 4.4). However, anycasting introduces multi-homing scenarios which differ significantly from normal multi-homing scenarios. In typical multihoming, multiple peerings of the multihomed stub AS are in the same geographical area. As a consequence, selection of paths from clients to the multihomed AS based on ISP policies and AS-PATH length leads to acceptable performance. On the other hand, for an anycasted AS, the multiple peerings are geographically dispersed, but this is not accounted for in the existing inter-domain routing set-up. In the example above, Level3 receives two paths to the anycasted internal deployment of equal AS-PATH length and is not aware of the actual physical difference between the length of these paths. Consequently, there is a good chance that Level3 may choose a route that causes the anycast packets to be routed to a distant anycast server. In other words, current route selection mechanisms have a much higher chance of making an unsuitable choice when selecting paths to an anycasted AS.

Although negative, the importance of this observation cannot be overemphasized. It brings to light that while the routing protocols used to choose paths to unicast destinations work naturally for anycast destinations too, the metrics used for routing decisions can lead to a poor choice of anycast server in terms of latency. While changing routing protocols to differentiate between anycast and unicast prefixes would be one possible approach to address this problem, a more practical approach would be to plan the deployment of the anycast servers to account for inter-domain route selection.

We hypothesize that deploying the anycast servers such that all of them have the same upstream provider and the servers are spread across the provider is

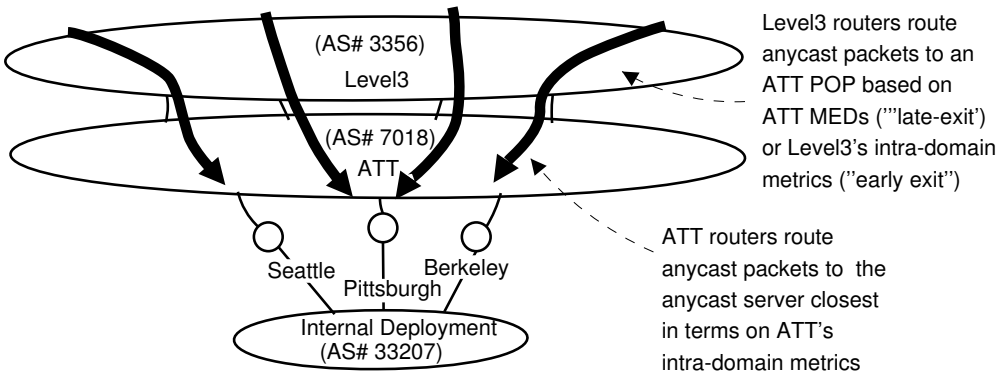


Figure 4.5: Shown here is a deployment with ATT as the common upstream provider – ASs beyond ATT route the anycast packets to ATT's network. Here, Level3's routers route anycast packets to a close-by ATT POP which then routes the packets to the closest anycast server.

one such deployment approach. This approach is based on two key observations. First, an ISP that is an upstream provider for some of the anycast servers routes incoming anycast traffic to the server closest among them; this is a consequence of the fact intra-domain traffic engineering is mostly consistent with latency sensitive routing [117]. For example, in case of the internal deployment, ATT is an upstream provider for three of the five anycast servers. Routers in ATT's network receive routes of equal AS-PATH length and equal preference from each of these anycast servers. Hence, the BGP decision process causes incoming anycast packets at any ATT POP to be routed to the server that is closest in terms of the intra-domain metrics used by ATT. As the measurements presented next show, this server is also closest in terms of latency in a large majority of cases.

Second, such a deployment decouples route selection at the common upstream provider from the selection at ASs beyond it. Due to reasons detailed above, the common upstream provider delivers incoming anycast packets to

the closest server. Any ASs farther away from the anycast server sites than the upstream provider only need to select among the possible routes to the upstream provider. Figure 4.5 illustrates this. Of these ASs, the ones that use an "early-exit" (or hot-potato routing) routing policy would route anycast packets to the closest POP of the common upstream provider. Alternatively, ASs that use a "late-exit" (or cold-potato routing) routing policy would honor the MEDs of the upstream provider and route anycast packets to the upstream provider POP that is most suitable for delivery to the closest of the deployed anycast servers. While it is possible that an AS may choose overly long paths to the upstream provider, the prevalence of the two aforementioned policies amongst ASs leads us to believe that this would not be the norm and our measurements confirm this. Its also important to note that any overhead arising due to the choices made at this step of the selection process also apply to the inter-domain routing in general and are not specific to anycast per se.

Hence, in case of the internal deployment, we would like to ensure that instead of five servers with three different upstream providers, all the servers should have the same provider. As a matter of fact, the subset of the internal deployment comprising of the three servers at Berkeley, Pittsburgh and Seattle conforms to our deployment proposal. These servers have the same upstream provider (ATT) and are geographically spread out. With this three server deployment, all anycast packets would be routed to the ATT network and then delivered by ATT to the (closest) anycast server. For instance, if the sample client (`net.berkeley.edu`) presented above accessed this deployment, it would be routed to the server at Berkeley. This is because, in figure 4.3, Level3 would not receive an advertisement for the anycast prefix from WCG. Instead, Level3's routers would route the client's packets to the ATT POP at San Francisco. The

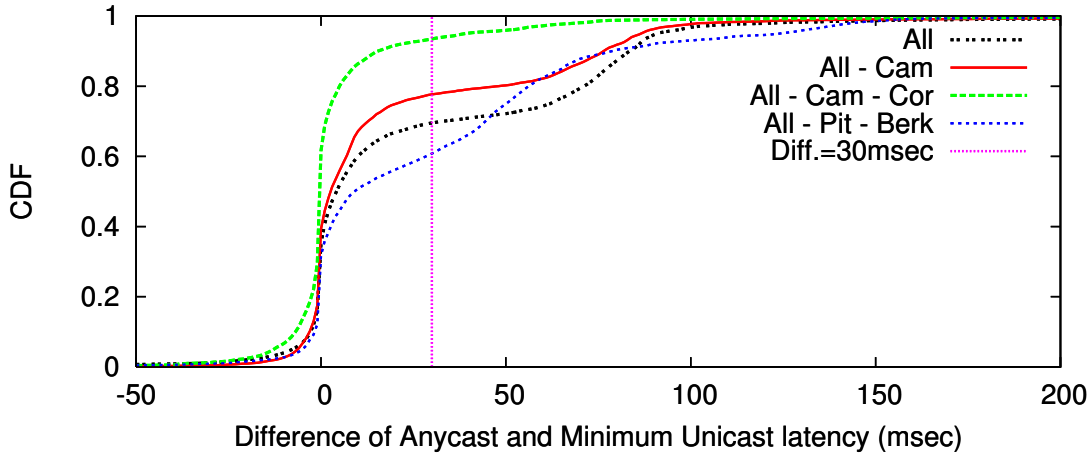


Figure 4.6: CDF for the difference between the anycast and minimum unicast latency for various subsets of the internal deployment. Here, All - x implies measurements with server x switched off.

routers in ATT's San Francisco POP would in turn route these packets to the nearest anycast server. Thus, the anycast packets would be routed to the server at Berkeley and not at Seattle or Pittsburgh. We validated this by stopping the advertisement of the anycast prefix from the Cornell and the Cambridge server to remove them from the anycast deployment and observing that the anycasted packets from the sample client were delivered to the anycast server at Berkeley. This path is labeled "Ideal Path" in Figure 4.3.

To validate this hypothesis, we repeated the stretch factor measurements for our list of clients with only subsets of the internal deployment operational. Figure 4.6 shows the results from these measurements. As can be seen, an anycast deployment comprising of just the three servers at Berkeley, Pittsburgh and Seattle (labeled as "All - Cam - Cor" in the figure) yields good proximity with just 5% of the clients incurring a stretch factor of more than 30msec. Note that it may seem that this improvement in the stretch factor for clients is due to the

fact that the choice of servers has reduced from five to three. To account for this we also measured the proximity offered by a deployment with the three servers at Seattle, Cornell and Cambridge. All these servers have a different upstream provider and as can be seen from the figure (curve labeled as "All - Pit - Berk"), this yields poor proximity too. These results show that an anycast deployment with all servers having the same upstream provider does provide good latency-based proximity to clients.

As a matter of fact, this result can be generalized. It is possible to have anycast deployments with multiple upstream providers for the anycast servers. However, all the upstream providers should have reasonable global geographic spread (in essence, tier-1 ISPs with a global network) and for each upstream provider, there must be a sufficient number of servers to cover the geographical spread of the provider. In such a set-up, whenever any of the upstream providers receives an anycast packet, there is a close-by anycast server that the packet can be routed to. ASs beyond these upstream providers would choose to route the anycast packets to one of the upstream providers and since all the providers are well covered by the anycast servers, this choice does not have a lot of bearing on the anycast path length. Due to reasons described earlier, each AS beyond the upstream providers for the anycast deployment would most likely route the packets to a suitable POP of the upstream provider it chooses. Thus, a deployment according to this model would offer good latency based proximity to clients. The small size of the existing internal deployment does not allow us to validate this claim; however, this is something that we aim to address in future work.

4.6 Failover time

Inter-domain IP Anycast involves each site advertising the anycast prefix into BGP. Consequently, when an anycast server fails, the process by which clients using the failed server are rerouted to other operational servers is tied to BGP convergence. Past studies have shown that failures at multi-homed end sites can lead to a convergence process that may last for minutes [86]. Such a slow failover, if it applies to IP Anycast to, would not bode well for a number of proposed uses of IP Anycast. Hence we decided to measure the failover rate for the internal anycast deployment. To the best of our knowledge, this represents the first attempt to study the failure-mode behavior of IP Anycast.

Methodology: In order to determine the rate at which clients failover when an anycast server fails, we need to be able to determine the specific anycast server that each client is routed to. To this effect, we configured the anycast servers in the internal deployment to act as authoritative nameservers for a domain under our control (`internal.anycast.guha.cc`) and to respond to TXT-type DNS queries for this domain with a location-specific string. For example, a TXT-type DNS query for this domain from a client whose anycast packets are routed to the anycast server at Cornell will receive "Cornell" as the response. This, when combined with the ability to direct clients to send DNS queries to the anycast address of the internal deployment, allows us to determine the anycast server being accessed by all clients in our list. Figure 4.7 illustrates this process for one client.

The internal deployment servers have been configured to respond to TXT-type queries with a TTL value of 0. This implies that clients cannot cache the

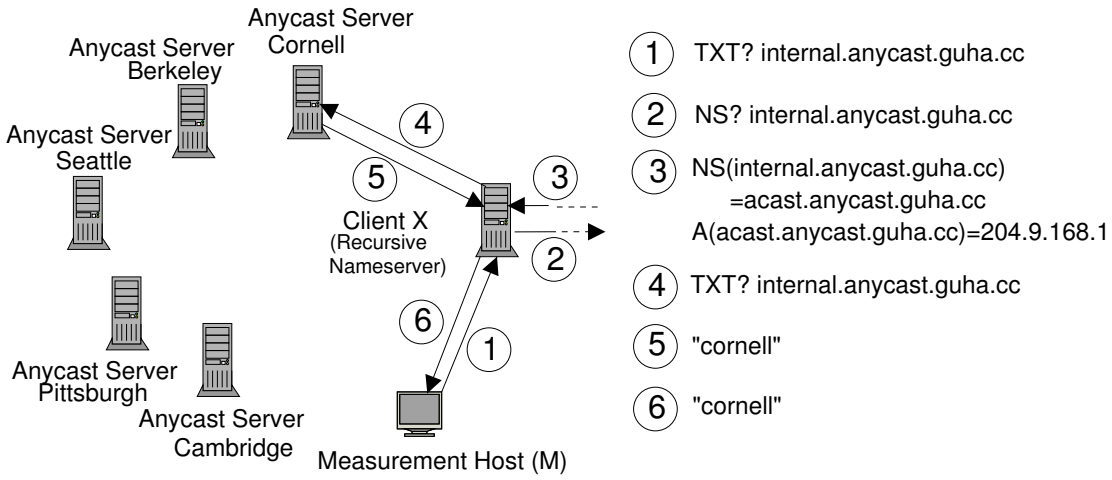


Figure 4.7: TXT-type DNS queries from client X to the internal deployment anycast address (204.9.168.1) are routed to the server at Cornell which responds with a location-specific string ("cornell")

response from the anycast server and hence, need to send packets to the deployment's anycast address each time they are queried. Also, note that since we don't have any way to determine the anycast server accessed by clients for the external deployments, the measurements in this and the following sections are restricted to the internal deployment.[6]

Given this, we determined the impact of the failure of each server in the internal deployment on clients being routed to that server. We induced failures at individual servers by tearing down their BGP peerings leading to a BGP withdrawal being generated for the anycast prefix. Concurrently, we sent the aforementioned TXT queries to the anycast address through the clients that were being routed to the failed server at a rate of once every five seconds for three minutes and at a rate of once per minute for the next fifty-seven minutes lead-

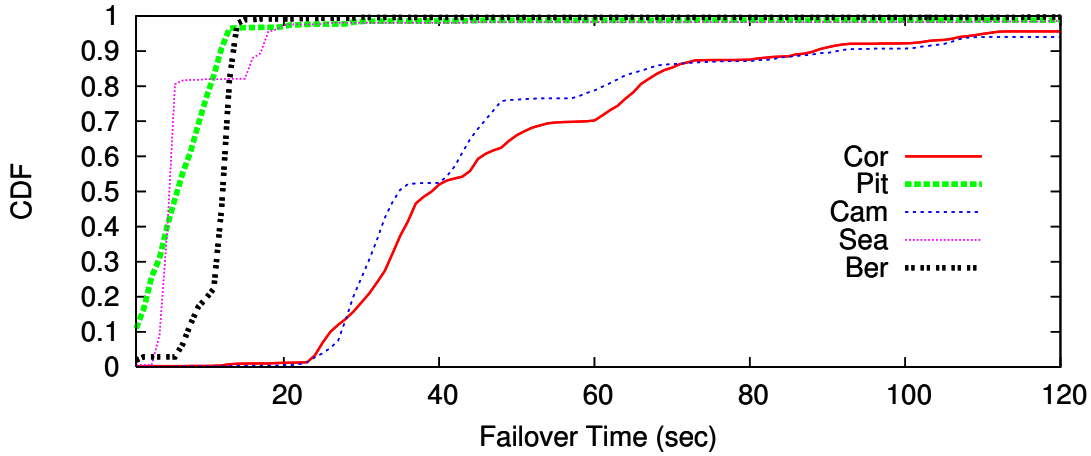
[6]Some of the anycasted DNS root server deployments do allow users to query them for the particular server the user is being routed to [145]. However, these queries have been chosen such that they cannot be generated through recursive nameservers, probably to avoid the possibility of this being used for a DNS-amplification attack on the root servers [147].

ing to a total probe period of one hour. For each such client, we determined the time it takes for the client to failover and "settle" at a different anycast server. This is referred to as the *failover time* for the client. Note that during the convergence process, a client may be temporarily routed to a server different from the server it is finally settles on. For example, a client accessing the Cornell server before it failed may be temporarily routed to the server at Cambridge before being routed to the server at Pittsburgh for good. The time between the failure and the first query that is routed to the Pittsburgh is the client's failover time.

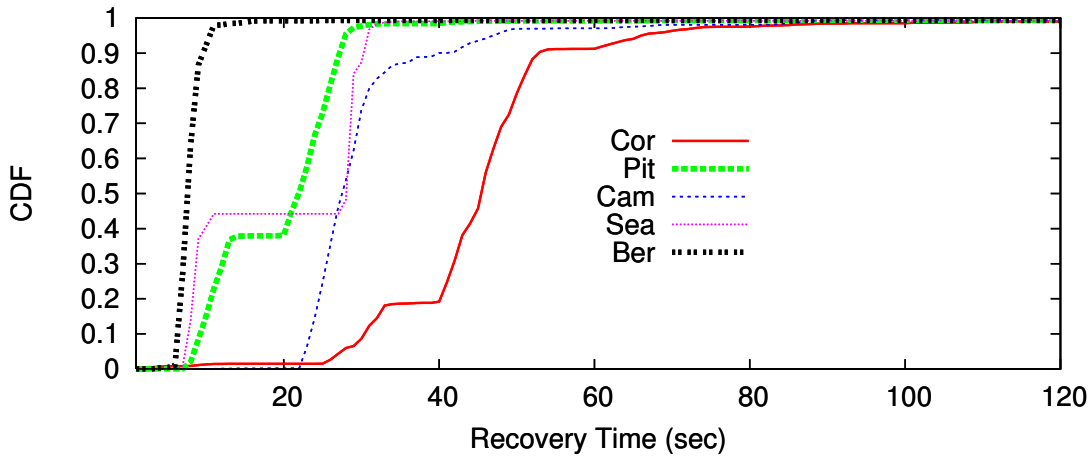
We would also like to clarify that unlike affinity (as discussed in the next section), fast failover is not very relevant with regard to the feasibility of running connection oriented services on top of IP Anycast. In a vast majority of the scenarios, the failure of a server would also break all client connections irrespective of how fast the failover process is. Instead, the failover time characterizes the time after a failure for which clients using the failed server cannot utilize the anycast service.

Similarly, we also restarted the failed servers by re-establishing their BGP peerings and determined the time it takes for clients that were originally using the server in question to be routed back to it again. This time between the re-establishment of the peering and the rerouting of a client to the server is referred to as the *recovery time* for the client. For each server in the internal deployment, we induced failures and restarts and measured the failover and the recovery time for the clients using the server. These experiments were repeated 5 times each – we avoided more runs because the experiments impose a heavy query load on nameservers that we don't own.

Results: Figure 4.8 shows the CDF for the failover and recovery times cor-



(a) Failover Time



(b) Recovery Time

Figure 4.8: The failover and recovery times for the servers in the internal deployment.

responding to each server in the deployment (the time axis has been shortened in the interest of clarity). The servers at Cornell and Cambridge have similar failure mode behavior with a median failover time of ≈35 seconds and a 95^{th} percentile of ≈120 seconds. The peaks around 30 and 60 seconds in these curves can be attributed to the BGP MinRouteAdverTimer which governs the time between updates for the same prefix to the BGP same peer. The value of this timer defaults to 30 seconds on many routers [90]. On the other hand, the servers at Berkeley, Pittsburgh and Seattle have a median failover time ranging from 7 to

12 seconds with a 95^{th} percentile of 14 to 18 seconds. The trends in the recovery times are less obvious and we don't completely understand these results.

In case of the server at Cornell, there is no other anycast server with WCG as an upstream provider. Consequently, the failure of this server causes clients using the server to be rerouted to servers with other upstream providers. This also implies that the BGP updates resulting from the server failure need to be propagated beyond WCG's network. As a result, the failover process involves a number of ASs, is impacted by the various BGP timers and overall, is affected by delayed routing convergence resulting in the large failover time. A similar explanation applies to the failure of the server at Cambridge.

On the other hand, the other three servers at Berkeley, Pittsburgh and Seattle offer much faster failover. This is an outcome of the fact that these servers have the same upstream provider (ATT). Thus, when one of the servers fails, clients accessing that server are routed to the one of the other two servers. To verify this, we determined the fraction of clients that go to other operational servers when a particular server fails. This fraction is plotted in figure 4.9. As can be seen, in case of a failure at Berkeley, Seattle and Pittsburgh, most of the clients are rerouted to one of the other two operational servers from the same group. This implies that when one of these servers fails, the BGP convergence process is restricted mostly to ATT's network resulting in faster failover.

The distribution of clients when a server fails (figure 4.9) can also be used to explain some of the trends in figure 4.8(a). For example, almost all the clients accessing the Berkeley server when it fails are routed to the server at Seattle. As a result, the failover time for the Berkeley server is almost the same for all clients accessing it. On the other hand, a small fraction of clients accessing the

Seattle server when it fails are routed to the server at Cornell.[7] This explains the inflection point in the failover time for the Seattle server showing that a small fraction of the clients take much more time to failover than the rest.

These results suggest that in an IP Anycast deployment with a number of anycast servers per upstream provider (as proposed in the previous section), the failure of an anycast server would cause clients to be routed to one of the other servers with the same upstream provider. Hence, the proposed deployment model is conducive to fast failover with a large majority of clients being rerouted to another server within 20 seconds. Reports that many commercial DNS-based anycast deployments aim for similar sub-minute failover times (by using a TTL values between 10-20 seconds [85,114]) lead us to conclude the failover rate of a planned IP Anycast deployment should suffice for almost all anycast applications.

On a broader note, our study shows that while results from previous studies reporting slow BGP convergence [86] do apply to IP Anycast deployments in general, an IP Anycast deployment can be planned so as to decouple anycast failover from delayed routing convergence. In effect, this addresses the long held belief that IP Anycast is bound to provide very slow failover, for example, the possibility of server failures causing outages of five or more minutes [143]. As a matter of fact, the clustered deployment model (as described in section 4.3) used by most commercial IP Anycast deployments was primarily motivated by the need to decouple host failure from BGP events. While we agree that using clustered hosts at each anycast server site is a necessary part of the IP Anycast

[7]The failure of the Seattle server probably causes ATT to modify the MEDs on the anycast prefix advertisements it propagates to its peers, some of whom then choose to route anycast packets to other servers that don't have ATT as an upstream provider. This explains why some clients are routed to Cornell when the server at Seattle fails.

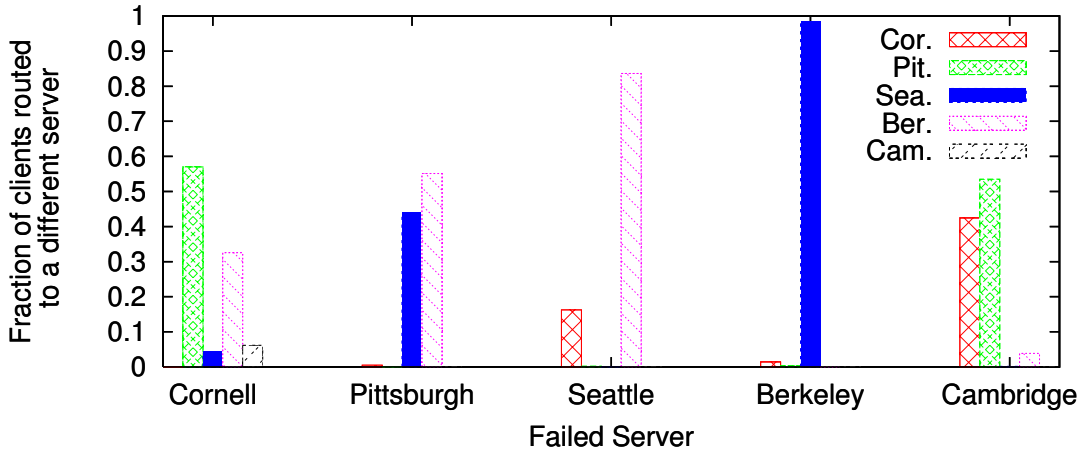


Figure 4.9: As a server fails, clients that were being routed to the server are now routed to other operational servers. The Y-axis shows the fraction of clients that failover to each other operational server when the particular server on the X-axis fails.

deployment picture, an IP Anycast deployment conforming to our deployment proposal ensures that even when an entire server site fails, clients do not have to wait an inordinate amount of time for failover.

4.7 Affinity

The fact that IP Anycast is a network-layer service implies that two consecutive anycast packets from a single client need not be routed to the same server. For example, in case of the internal deployment, it is possible that a client whose anycast packets were being routed to the server at Cornell suddenly gets routed to the server at Cambridge. Such an occurrence, referred to as a *flap*, would break any higher-layer connections (for example, TCP connections) that may exist between the client and the anycast service. Hence, determining the affinity offered by IP Anycast is important for characterizing its impact on stateful

services being run on top.

As mentioned earlier, past studies have painted a contradictory picture of IP Anycast affinity. For instance, operators of the anycasted K root server [33] found that their IP Anycast deployment offers very good affinity. On the other hand, a study based on a few (<200) volunteer and PlanetLab nodes [24,111] and anecdotal evidence from the anycasted J root server [20] support claims to the contrary. For example, Boothe et. al. [24] found the median inter-flap duration to be 1.4 hours for PlanetLab nodes and 3 hours for their volunteer nodes. As a matter of fact, IP Anycast affinity and its suitability for stateful services has been passionately debated on many mailing lists; a summary of some these discussions can be found at [144]. However, none of the aforementioned studies have attempted to delve into the reasons behind the routing flaps (few or many) observed by them. In this section, we present a detailed analysis of the affinity offered by the internal anycast deployment as determined through active probing from our clients.

Methodology: The TXT-record based querying described in section 3.5.3 allows us to determine the particular anycast server that anycast packets from a given client are routed to. Using this, we can periodically query a client in order to capture the anycast flaps experienced by it. For these experiments, we randomly chose 5200 clients from our list of clients and determined the number of flaps they experience by querying them at a rate of once per minute for a period of 17 days.

Results: Figure 4.10 shows a CDF for the number of clients observing the respective number of flaps. The figure shows that ~40% of the clients do not ob-

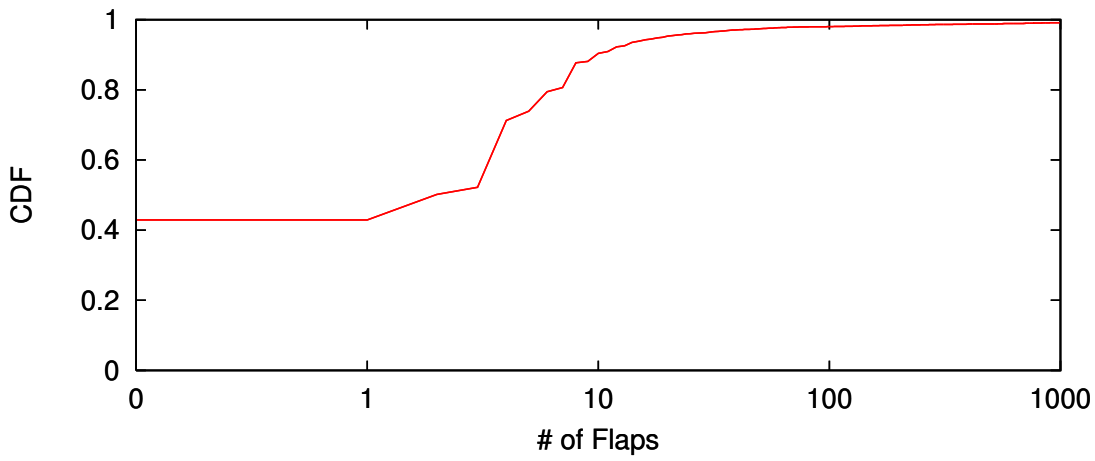


Figure 4.10: Affinity measurements for our anycast deployment. The measurements involve 5277 nameservers as vantage points and span a period of 17 days.

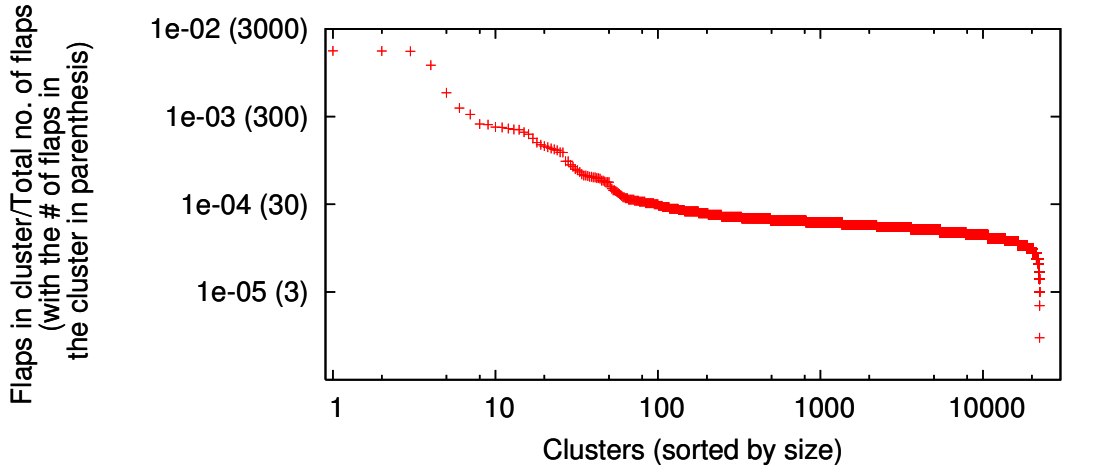


Figure 4.11: Clustered flaps and their contribution towards the total number of flaps – there are a small number of large clusters but a majority of the flaps belong to very small sized clusters.

serve any flap, ~95% of the nodes observe less than a flap per day (less than 17 flaps in total) and ~99% of the clients observe less than 10 flaps per day. Overall, the trace comprises of 290,814 flaps of which 266,967 (~92% of the total) were observed by just 58 (~1%) clients. Hence, apart from this very small fraction of nodes, the internal anycast deployment seems to provide very good affinity.

These anycast flaps can occur due to a variety of events ranging from link, peering or server failures to ASs load balancing the anycast traffic across multiple links. Typically, the impact of a given event depends on its location: events near an anycast server or in a core ISP would cause a number of clients to flap while the impact of an event close to a client site would be restricted to a small number of clients. Given this observation, we proceeded to use temporal clustering to construct events out of the flaps seen in our trace. The idea here is to cluster flaps that occur close by in time on the assumption that they are probably due to the same routing event. Hence, we clustered the flaps observed at all the clients such that flaps within 10 seconds of each other were in the same cluster. Since flaps in the cluster are assumed to be due to the same routing event, we limited the maximum cluster size to 180 seconds, i.e. no two flaps more than 180 seconds apart can be in the same cluster. This is similar to the BGP update clustering used in [49]. The clustering results presented here are not very sensitive to the maximum cluster size – clustering with a maximum cluster size of 120 and 240 seconds yielded similar results. Using this approach, the 290,814 flaps yielded 22,319 clusters with the largest cluster containing 1,634 flaps.

Figure 4.11 shows the distribution of the fraction of the total number of flaps that occur in the cluster. The figure has a small number of moderate-to-large clusters and these correspond to infrequent BGP events near the servers or in

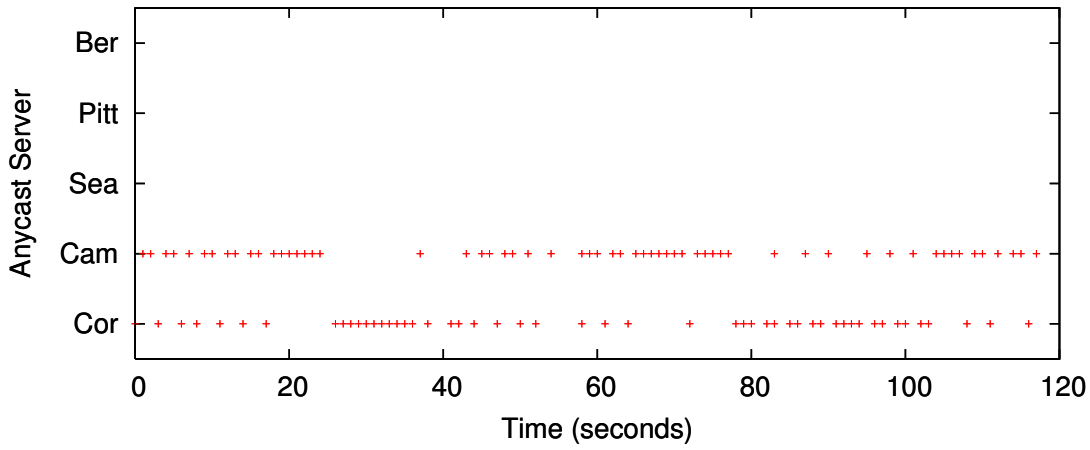


Figure 4.12: Probes at a rate of once per second from an unstable client. Each plotted point in the figure represents a probe and shows the server it is routed to – as can be seen, the client flaps very frequently between the anycast servers at Cornell and Cambridge.

core ISPs. This further buttresses our argument that IP Anycast does not have any harmful interactions with inter-domain routing. More importantly, the figure has a long tail depicting a very large number of very small clusters. These correspond to events near clients. As a matter of fact, a large majority of these clusters comprise of flaps seen at the highly unstable clients. This points to the fact that even the very small fraction of clients that observe poor affinity do so due to events that are close to them.

Further, the frequency at which the unstable clients observe flaps leads us to believe that these clients are multihomed and are using some kind of dynamic load balancing across multiple upstream providers. For example, the client which observed the most flaps belongs to AS# 15710 which, in turn, has two upstream providers – AS# 3356 (Level3) and AS# 8928 (INTEROUTE). This client observed almost continuous flaps between the server at Cornell and Cam-

bridge in our trace. To investigate further, we probed this client at a rate of once per second for a period of two hours. Figure 4.12 shows the server each probe is routed to (for clarity, we show the probes only for a two minute duration – the rest of the trace is similar). The figure shows that the client experiences very frequent flaps. In many cases, consecutive probes are routed to different servers. Given that BGP events occur at much coarser time scale, this high frequency of flaps suggests dynamic load balancing by the client.

To validate our conjecture, we used the view of inter-domain routing available through Route-Views [149] and CID R-Report [142] to determine the AS-level connectivity of the unstable clients. The 58 unstable clients belong to 47 distinct ASs and at least 42 of these have multiple upstream ASs. Since we do not have control over these clients, and hence cannot determine if the client ASs are indeed load balancing across their upstream ASs, we conducted an e-mail survey of the client ASs to determine if this is indeed the case. The survey yielded just five responses, though all the five ASs claimed to be using some form of load balancing across their provider. While the exact set-up of these clients begs further investigation, all the evidence at hand leads us to conclude that dynamic load balancing by clients is the root cause of the poor affinity observed by them.

4.8 Client Load Distribution

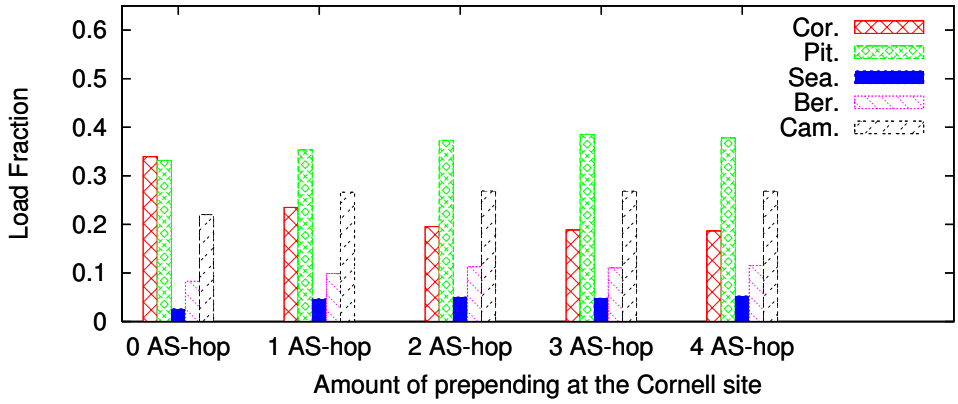
Clients access an IP Anycast deployment simply by sending packets to its anycast address, and it is the routing infrastructure that is responsible for delivering the packets to the one of the servers. Consequently, anycast operators don't

have any control over the number of clients that each server handles. In this section, we study the distribution of client load across the internal deployment and evaluate means by which this can be controlled.

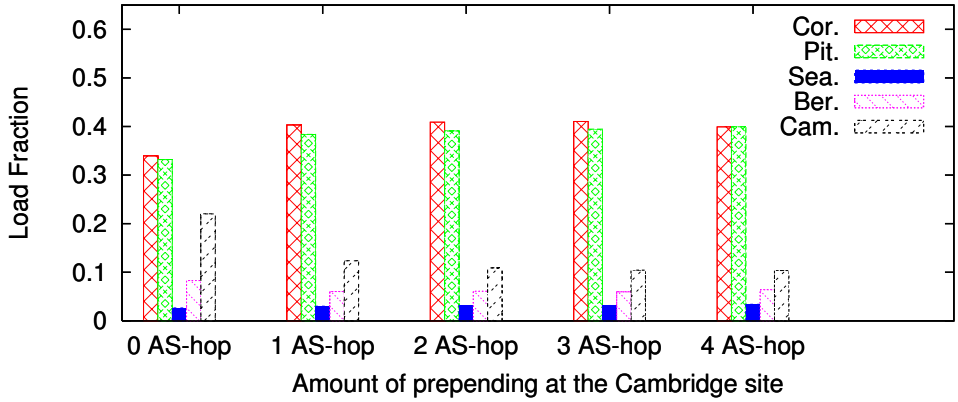
We used the TXT-record based querying described in section 3.5.3 to determine the distribution of all clients in our list across the servers in the internal deployment. The set of bars corresponding to "0 AS-hop" in figure 4.13(a) shows this distribution. As can be seen, in the default set-up, most of the clients (≈90%) are routed to the server at Cornell, Cambridge or Pittsburgh. This result is consistent with the uneven load distribution of clients across the anycasted J root servers [20] and K root servers [33]. Hence, IP Anycast, by itself, does not balance client load across the anycast servers.

It is interesting to note that the distribution of clients is skewed even amongst the servers that have ATT as their upstream, because many more clients are routed to Pittsburgh than to Seattle and Berkeley. On the other hand, the proximity measurements in Section 4.5 showed that anycast packets coming into the ATT network are routed to the closest of these three anycast servers. This would imply that most of the clients in our list that are routed to ATT are closer to the server at Pittsburgh than to the other two servers. In effect, this brings out the implicit trade-off between proximity and load balance. An anycast deployment which offers optimal latency-based selection of the anycast server is unlikely to achieve an even distribution of clients across the servers in the deployment.

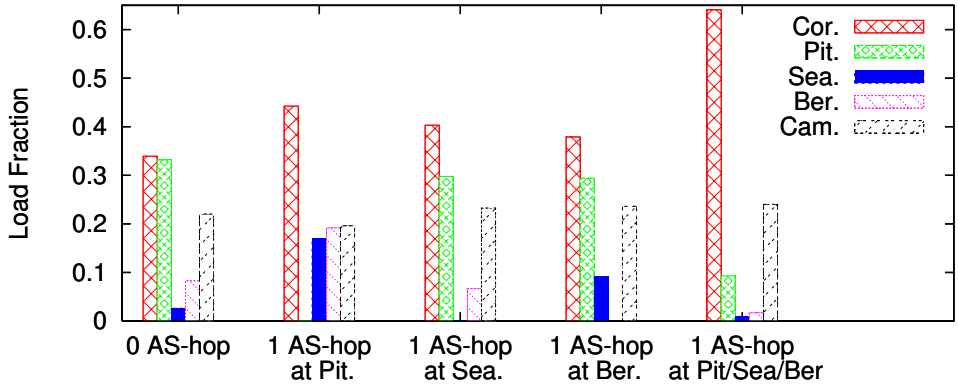
Given the uneven load distribution, we would like to investigate if anycast operators can manipulate routing advertisements at individual server sites to control the number of clients routed to them. In the rest of this section we evaluate the effectiveness of AS-PATH prepending as a means of controlling the client



(a) Prepending at Cornell



(b) Prepending at Cambridge



(c) Prepending at Pittsburgh/Seattle/Berkeley

Figure 4.13: Load on the anycast sites of anycast deployment in the default case and with various kinds of AS path prepending at the sites. Here, Load Fraction for a site is the ratio of the number of clients using the site to the total number of clients (≈20000).

load on anycast servers. AS-PATH prepending at a server involves increasing the length of the AS-PATH in the BGP advertisement that the server propagates and hence, should wean some of the clients away from it. For example, when "1 AS-hop" prepending is used at the Cornell site, the AS-PATH seen by the Cornell server's peer is [33207 33207].[8] Similarly, "2 AS-hop" prepending implies that the Cornell server advertises the internal deployment anycast prefix with the AS-PATH as [33207 33207 33207].

Figure 4.13(a) shows the distribution of client load across the anycast servers when path prepending is used at the Cornell server. As can be seen, prepending 1 AS-hop causes the fraction of clients being routed to the Cornell server to reduce from 34% to 23%. However, the reduction in load tapers off beyond this with ≈18% of the clients being routed to the Cornell server irrespective of the amount of prepending used. This can be explained in terms of typical ISP policies and the BGP decision process: the ISP policy (expressed as weights and local preferences) has higher priority than the AS-PATH length in the BGP decision process. Thus, ASs that choose to use the Cornell server as dictated by their routing policies are oblivious to the amount of prepending being done and the impact of path prepending soon runs into diminishing returns. Figure 4.13(b) shows the variation of client load with path prepending at the Cambridge server. These results follow a pattern similar to what is described above with the client load on the Cambridge server reducing and then tapering off.

Figure 4.13(c) shows the distribution of client load with path prepending at Pittsburgh, Seattle and Berkeley. These are the three anycast sites with the same upstream provider (ATT). The figure shows that prepending the AS-PATH of the advertisement at the Pittsburgh server causes the number of clients routed

[8] As mentioned earlier, 33207 is the origin AS used for the internal deployment.

to Pittsburgh to drop to zero. Since the servers at Seattle and Berkeley also have ATT as the upstream provider, using AS-PATH prepending at Pittsburgh causes routers in ATT's network to prefer both the Berkeley and the Seattle servers to the Pittsburgh server. Hence, all anycast traffic reaching ATT's network is split between the Seattle and Berkeley servers only. Results for 1-AS hop prepending at Seattle and Berkeley are analogous.

These results imply that if some servers in an anycast deployment have the same upstream provider, all of them need to prepend the AS-PATH in their advertisements in order to divert clients away from them. In the internal deployment, diverting clients away from the three servers with ATT as their upstream ISP (Pittsburgh, Seattle and Berkeley) would require all of them to use path prepending. The final set of bars in figure 4.13(c) shows the load distribution across our deployment in such a scenario.

Hence, while AS-PATH prepending can be used to manipulate load across servers with different providers, it is not effective for manipulating load within the set of servers with the same provider. Alternatively, in an IP Anycast service deployed according to the model presented in Section 4.5, AS-PATH prepending can only be used for balancing load across between groups of servers that have the same upstream ISP. For example, consider a deployment with a few servers having ATT as their upstream provider and the rest having WCG as their upstream provider. The servers with ATT as their upstream need to use AS-PATH prepending together to divert clients towards the servers with WCG as their upstream. For balancing the client load across servers with the same upstream provider, the servers need some sort of traffic engineering arrangement with their common provider. For instance, many ISPs allow their customers

to manipulate incoming traffic through the use of specific BGP community attributes in their routing advertisements [29,150]. Anycast servers with the same provider can thus use such mechanisms to coarsely control the number of clients routed to them. We are currently in the process of talks with ISPs who would allow us to host anycast servers and experiment with such mechanisms.

Finally, note that these mechanisms provide operators with a coarse-grained control over the distribution of clients across server sites (or groups of sites). For instance, this could be used by anycast operators in the face of a DoS attack on the deployment to redistribute traffic away from server sites under strain. Beyond this, anycast operators can use load balancing devices at server sites for a fine grained control over the distribution of clients being served by the site across the hosts that are part of the site. As a matter of fact, current commercial IP Anycast deployments use such mechanisms for balancing the number of clients served by individual cluster hosts at each site.

4.9 Discussion

Section 4.2 described previous IP Anycast measurement studies. Here we discuss other research efforts relevant to IP Anycast and relate our study to this broader context.

In addition to its implementation at the network layer, anycast can also be implemented at the application layer. *Application-layer anycast* provides a one-to-any service that maps a high-level name, such as a DNS name, into one of multiple servers, returning the selected server's IP address to the client. Such an approach offers a number of advantages over IP Anycast: it is easier to

deploy, offers fine-grained control over the load on the servers and can provide very fast failover to clients. And indeed, these advantages have led to the widespread adoption of application layer anycast as a service discovery primitive. For example, commercial CDNs [139] use DNS-based redirection (in combination with URL-rewriting) to direct clients to an appropriate server. Many proposals [47,53,121,131,137] in the academic community have also explored such designs.

In spite of these advantages, application-layer anycast is not a panacea. The fact that IP Anycast operates at the network layer implies that it is the only form of anycast that can be used by low-level protocols; e.g., the use of anycast in IPv4-to-IPv6 transition [67]. As importantly, operating at the network layer gives IP Anycast a "ground level" resilience not easily achieved by application-layer anycast — e.g., using DNS-based redirection to achieve resilience across a group of web servers requires first that the DNS servers themselves be available. It is this that makes IP Anycast particularly well suited for replicating critical infrastructures such as the DNS. For applications that do use IP Anycast, our deployment proposal can be used to build an anycast service that offers good proximity, fast failover, and control over the distribution of client load.

While IP Anycast functionality is available even today, it scales poorly in the number of anycast groups. Recognizing its many advantages, GIA [75] and PIAS (as discussed in the next chapter) seek to make IP Anycast more broadly usable and propose solutions to improve the scalability of IP Anycast. Our study focusses on the basic effectiveness of IP Anycast, not its scalability, and our results are relevant to the performance one might expect of an IP Anycast service whether implemented as a proxy-based service (PIAS) or a more scalable

IP-layer implementation (GIA).

Past studies have analyzed the use of AS-PATH prepending as a traffic engineering tool for multi-homed stub sites [88,104] and have proposed automated mechanisms for this [28]. However, we are not aware of any studies analyzing the use of AS-PATH prepending as a mechanism for controlling the distribution of client load across anycasted servers. Similarly, we conjecture other traffic engineering techniques can also be used as a load distribution mechanism by anycast operators. For example, some of the F root servers use the BGP `no-export` attribute as part of their anycast advertisements. This restricts the scope of the advertisement emanating from the server and hence reduces the number of clients served by it.

Our study is limited in several aspects. First, the size of the internal deployment raises concerns regarding the generality of our results and benefits of our proposed deployment model for larger deployments. Second, our internal deployment setup does not allow us to evaluate the effectiveness of certain other traffic engineering techniques for controlling client load. Finally, the use of external DNS nameservers as clients in our study restricted the amount and rate of probing that could be done. For the same reason, our conjectures regarding load balancing at clients had to be verified using heuristics and survey data. Nonetheless, we hope the measurement techniques presented in this chapter can serve in the large-scale evaluation of experimental anycast deployments along the lines presented here.

4.10 Summary

This chapter presented the results of a detailed measurement study of IP Anycast. Our study differs from previous efforts on two fronts. First, we evaluate IP Anycast deployments from a large number (>20,000) number of client vantage points. Second, we deploy our own IP Anycast service to perform controlled experiments that, for example, allow us to study the failure-mode behavior of IP Anycast. Our findings include:

1. IP Anycast, by itself, does not route clients to servers that are close in terms of latency.
2. IP Anycast is affected by delayed routing convergence and may be slow in rerouting clients in the face of server failures.
3. IP Anycast offers good affinity to all clients with the exception of a small fraction that explicitly load balance traffic across multiple upstream providers. In other words, we find IP Anycast does not interact poorly with inter-domain routing and hence, should not significantly impact stateful services.
4. IP Anycast services experience a skewed distribution of client load across the anycast servers.

Based on these measurements, we hypothesize that an IP Anycast deployment with a single upstream provider and with servers spread across this provider would offer good latency-based proximity. Our evaluation shows that this holds in our internal anycast deployment. Further, we generalize this model and argue that for good proximity in an IP Anycast deployment with multiple

upstream providers, each major upstream provider should be geographically spread and well covered by anycast servers. Our evaluation also suggests that such a deployment model provides fast failover to clients. However, an evaluation of this approach over larger deployments and fully characterizing the proximity within such a model is a topic of future work.

We also evaluate the effectiveness of AS-PATH prepending to manipulate the distribution of client load across servers and find that it can be used for controlling the number of clients routed to groups of anycast servers with the same upstream provider. Overall, we find that an IP Anycast service can be deployed to offer good proximity and fast failover to clients while allowing for coarse-grained control over the distribution of client load across the deployment. In the next chapter, we build upon these insights to design a new anycast architecture that addresses the limitations of native IP Anycast.

CHAPTER 5

PROXY IP ANYCAST SERVICE

5.1 Overview

IP Anycast, with its innate ability to find nearby resources, has long been considered an important means of service discovery. Further, our measurements show that with a planned deployment, IP Anycast can provide robust and efficient one-to-any communication. The growth of P2P applications presents appealing new uses for IP Anycast. Unfortunately, as detailed in section 1.2, IP Anycast suffers from serious problems: it is very hard to deploy globally, it scales poorly in the number of anycast groups, and it lacks important features like load balancing. As a result, its use is limited to a few critical infrastructure services such as DNS root servers.

The primary contribution of this chapter is the detailed description of a deployment architecture for an IP Anycast service that overcomes the limitations of today's "native" IP Anycast while adding new features, some typically associated with application-level anycast, and some completely new. This architecture, called **PIAS (Proxy IP Anycast Service)**, is based on the thesis that the limitations of native IP Anycast arise from the very fact that provides most of its strengths – the anycast functionality is provided by the routing infrastructure. Consequently, PIAS uses tunnels to decouple the anycast service offered by it from native IP Anycast. PIAS is composed as an overlay, and utilizes but does not impact the IP routing infrastructure. This allows it to combine the best of both worlds.

At a high-level, PIAS comprises of a set of proxy nodes that advertise IP Anycast addresses on behalf of group members. PIAS allows an endhost in an anycast group (anycast group member, or **anycast target**) to receive anycast packets for that group via its normal unicast address and normal protocol stack. The anycast target *joins* the anycast group simply by transmitting a request packet to an anycast address (again, via its unicast interface). The target may likewise *leave* the group through a request packet, or by simply becoming silent.

Packets from clients trying to access an anycast group are routed by native IP Anycast to a proxy which then tunnels the packets to the appropriate group member. The fact that PIAS is an IP Anycast service means that *clients* use the service completely transparently—that is, with their existing IP stacks and applications.

PIAS utilizes the IP address space efficiently: thousands of IP anycast groups may be identified through a single IP address. It scales well by the number of groups, group size and group churn with virtually no impact on the IP routing infrastructure. It provides fast failover in response to failures of both target hosts and PIAS infrastructure nodes.

PIAS can select targets based on criteria other than proximity to the sending host, notably including the ability to load balance among targets. PIAS has the unique feature that an anycast group member can also transmit packets to other members of the same anycast group. This is in contrast to native IP Anycast, where a group member would receive its own packet if it transmitted to the group. This feature makes IP Anycast available to P2P applications, something not possible if a host can't both send to and receive from the anycast group.

5.2 Design Goals

This section specifically lays out the design goals of PIAS, and briefly comments on how well PIAS meets those goals. The subsequent design description section refers back to these goals as needed. The goals are listed here in two parts. The first part lists those goals that are accomplished by native IP Anycast, and that we wish to retain. The second part lists those goals that are not accomplished by native IP Anycast. In this way, we effectively highlight the weaknesses of IP anycast, and the contributions of PIAS.

1. *Backwards Compatible*: Native IP Anycast is completely transparent to clients and routers , and we believe that this transparency is critical to the success of a new IP Anycast service. Because PIAS is an overlay technology that uses native IP Anycast, it does not change clients and routers.
2. *Scale by group size*: By virtue of being totally distributed among routers, native IP Anycast scales well by group size. PIAS has no inherent group size limitation. PIAS is deployed as an overlay infrastructure, and can scale arbitrarily according to the size of that infrastructure.
3. *Efficient packet transfer*: Because native IP Anycast uses IP routing, its paths are naturally efficient. As an overlay, PIAS imposes some stretch penalty on the paths packets take. The penalty imposed by PIAS is small (section 5.4.2), and shrinks as the PIAS infrastructure grows.
4. *Robustness*: Native IP Anycast's robustness properties (including packet loss) are similar to IP unicast. PIAS is engineered to be similarly robust.
5. *Fast failover*: Failover speed in Native IP Anycast depends on the convergence speed of the underlying routing algorithms, and can be fast (OSPF)

or somewhat slow (BGP). PIAS can be engineered to almost always rely on OSPF for certain types of failover (section 5.3.6). The PIAS overlay exposes additional failover situations that go beyond IP routing, and these are handled accordingly (Section 5.3.6).

The following are the goals that native IP Anycast does not satisfy.

6. *Easy* joins *and* leaves: Target hosts must not have to interact with IP routing to join and leave.
7. *Scale by the number of groups*: In addition to scaling by the usual metrics of memory and bandwidth, we require that PIAS also make efficient use of the IP address space. PIAS is able to accommodate thousands of groups within a single address by incorporating TCP and UDP port numbers as part of the group address.
8. *Scale by group dynamics*: Globally, IP routing behaves very badly when routes are frequently added and withdrawn. The PIAS overlay hides member dynamics from IP routing, and can handle dynamics caused both by continuous member churn and flash crowds (including those caused by DDoS attacks).
9. *Support target selection criteria*: IP Anycast can only select targets based on proximity. At a minimum, we wish to add load and connection affinity as criteria.

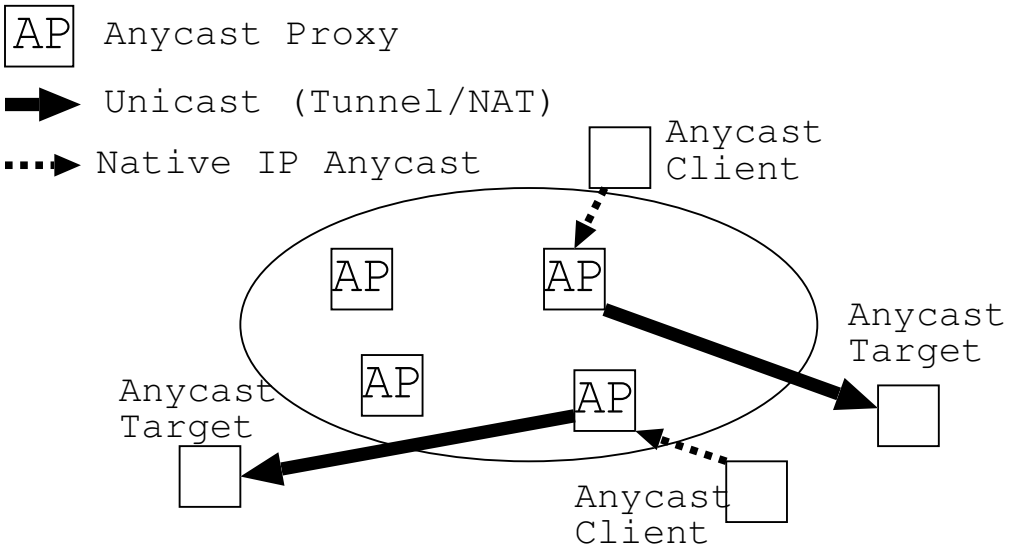


Figure 5.1: Proxy Architecture: the client packets reaching the proxies through native IP Anycast are tunneled to the targets

5.3 Design Description

This section gives a detailed description of PIAS. We take a "layered" approach to the description—we start with the core concepts and basic design and then step-by-step describe additional functionality that satisfies specific goals listed in section 5.2.

PIAS is deployed as an overlay infrastructure. It may be deployed by a CDN company like Akamai, by multiple cooperating ISPs, or even by a single ISP (though the efficacy of proximity discovery would be limited by the ISP's geographic coverage). Multiple distinct PIAS infrastructures may be deployed. In this case, each operates using distinct blocks of IP Anycast addresses, and they do not interact with each other.[1] In the remainder of this document, for simplicity of exposition, we assume a single PIAS infrastructure.

[1]Indeed, a single operator could deploy multiple distinct PIAS infrastructures as a way to scale.

The basic idea of PIAS, illustrated in Figure 5.1, is very simple. Router-like boxes, hereon referred to as ***anycast proxies*** (AP or simply *proxies*), are deployed at various locations in the Internet, for example at POPs (Point of Presence) of different ISPs. These proxies advertise the same block of IP addresses, referred to as the ***anycast prefix***, into the routing fabric (BGP, IGP). As such, the proxies are reachable by native IP Anycast—a packet transmitted to the anycast prefix will reach the closest proxy. However, these proxies are not the actual ***anycast target destinations*** (AT).[2] Rather, true to their name, they proxy packets that reach them via native IP Anycast to the true target destinations using unicast IP. This proxying can take the form of lightweight tunnels or NAT. NAT allows for backwards compatibility with the protocol stack at target hosts, but increases processing at the proxy.

This novel combination of native IP Anycast with tunneling to the unicast addresses of the targets allows PIAS to fulfill three critical design goals and drives the rest of the system design. First, it allows for efficient use of the address space as all the IP addresses in the prefix advertised by the proxies can be used by different anycast groups. In fact, PIAS does one better. It identifies an anycast group by the full ***transport address*** (TA), i.e. IP address and TCP/UDP port, thus allowing thousands of anycast groups per IP address. Second, it solves the IP routing scaling problem by allowing many anycast groups to share a single address prefix and hence, fulfills goal 7. Finally, it relieves targets from the burden of interacting with the routing substrate. They can join an anycast group by registering with a nearby proxy that is discovered using native IP Anycast. This fulfills goal 6.

The reader may notice two suspicious claims in the last paragraph. First, we

[2]The members of the anycast group; hereon referred to as **anycast targets** or simply **targets**.

claim to ease deployment by running unicast at the target instead of anycast, and yet the proxies still must run anycast. So, how is this an improvement? The benefit is that the difficult work of deploying IP Anycast is borne by the anycast provider once, and amortized across many anycast groups. Second, we claim to improve scaling by allowing thousands of IP Anycast groups to share a single IP address prefix. All we've really done, however, is to move the scaling problem from the IP routing domain to the PIAS infrastructure domain. This is quite intentional. As we argue later on, the scaling issues are much easier to deal with in the overlay than in IP routing.

PIAS offers two primitives to the members of an anycast group, which involve sending messages to a nearby proxy:

- *join(IP_A:$port_A$,IP_T:$port_T$,options)*: this message instructs the proxy to forward packets addressed to the anycast group identified by the TA IP_A:$port_A$ to the joining node's unicast TA IP_T:$port_T$. The *options* may specify additional information such as the selection criteria (load balance etc.), delivery semantics (scoping etc.), or security parameters needed to authenticate the target host. These are discussed later.
- *leave(IP_A:$port_A$,IP_T:$port_T$,options)*: this message informs the proxy that the target identified by TA IP_T:$port_T$ has left the group IP_A:$port_A$. *options* are the security parameters.

The join and leave messages are transmitted to the anycast address IP_A (that belongs to the anycast prefix) at some well-known port that is dedicated to receiving registration messages. This means that no extra configuration is required for a target to discover a nearby proxy.

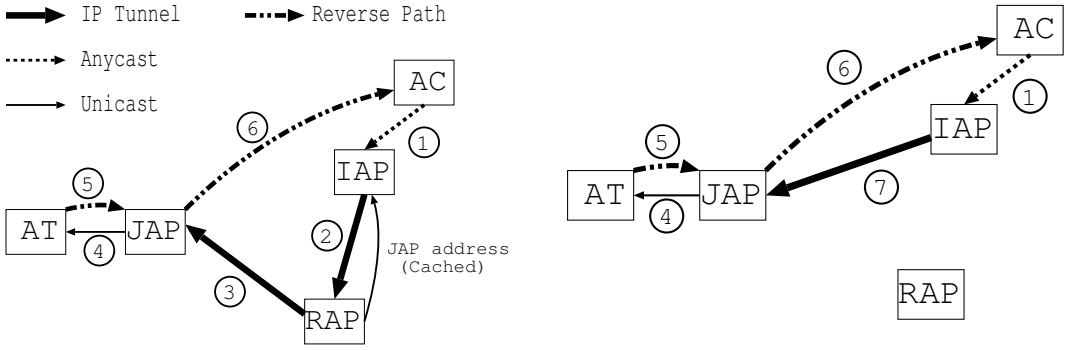
Note that we don't specify a "create group" primitive. For the purpose of this thesis, we assume that the first join essentially results in the creation of the group. In practice, a subscriber to the service would presumably have entered into a contract with the anycast service provider, which would have resulted in the assignment of anycast TAs to that subscriber. The subscriber would also have obtained authentication information using which targets may join the group. While the issues surrounding this sort of group creation are important, they are not central to the PIAS architecture, and we don't discuss them further.

5.3.1 The Join Anycast Proxy (JAP)

A target may leave a group either through the leave primitive, or by simply falling silent (for instance, because the target is abruptly shut off or loses its attachment to the Internet). This means that the ***Join AP*** (JAP—the nearby proxy with which the target registers; shown in figure 5.2) must monitor the health of its targets, determine when they are no longer available, and treat them as having left the group. The proximity of the JAP to the target makes it ideal for this.

The JAP must also inform zero or more other anycast proxies (APs) of the target(s) that have registered with it. This is because not all APs may be JAPs for a given group (that is, no target joined through them), but **anycast clients** (ACs) may nevertheless send them packets destined for the group. A proxy that receives packets directly from a client is referred to as the ***Ingress AP*** (IAP) for the client.[3] Note that the client-IAP relation is established using native IP

[3]In figure 5.1 the proxies in the client-target path are IAPs.



No.	Source	Dest	Comment
1	AC:p	AA:g	Native IP Anycast
2	IAP:AC:p	RAP:AA:g	IP-IP tunnel
3	RAP:AC:p	JAP:AA:g	IP-IP tunnel
4	JAP:q	AT:r	Unicast IP JAP:q = NAT(AC:p)
5	AT:r	JAP:q	Unicast IP
6	AA:g	AC:p	Unicast IP AC:p = NAT^{-1}(JAP:q)
7	IAP:AC:p	JAP:AA:g	IP-IP tunnel

Figure 5.2: Initial (left) and subsequent (right) packet path. The table shows the various packet headers. Symbols in block letters represent IP addresses, small letters represent ports. AA(Anycast Address) is one address in the address block being advertised by PIAS, AA:g is the transport address assigned to the group the target belongs to, while AT:r is the transport address at which the target wants to accept packets. Here, the target joined the group by invoking join(AA:g,AT:r,options)

anycast. As an IAP, the proxy must know how to forward packets towards a target; even though the IAP may not explicitly know of the target.

One possible way to achieve this would have the JAP spread information about targets associated with it to all proxies. This allows the IAP to tunnel packets directly to clients (as in Figure 5.1). However, such an approach would hamper PIAS's ability to support a large number of groups. In fact, Figure 5.1 is

conceptual—PIAS's approach for spreading group information is described in the next section and the actual paths taken by packets are shown in Figure 5.2.

5.3.2 Scale by the number of groups

In the previous section, we mentioned the need for a scheme that would allow PIAS to manage group membership information while scaling to a large number of groups. For any given group, we designate a small number of APs (three or four) to maintain a list of JAPs for the group. When acting in this role, we call the AP a ***Rendezvous Anycast Proxy*** (RAP). All APs can act as RAPs (as well as as JAPs and IAPs).

The RAPs associated with any given group are selected with a consistent hash [71] executed over all APs. This suggests that each proxy knows all other proxies, and maintains their current up/down status. This is possible, however, because we can assume a relatively small number of global APs ($\leq 20,000$, a number we derive later). We also assume that, like infrastructure routers, APs are stable and rarely crash or go out of service. The APs can maintain each other's up/down status through flooding, gossip [109] or a hierarchical structure [60]. The current implementation uses flooding. Such an arrangement establishes a simple one-hop DHT and hence, limits the latency overhead of routing through the proxy overlay.

When a proxy becomes a JAP for the group (i.e. a target of the group registers with it), it uses consistent hashing to determine all the RAPs for the group and informs them of the join. This allows the RAP to build a table of JAPs for the group.

The concept of the RAP leads to a packet path as shown on the left side of Figure 5.2. When an IAP receives a packet for an anycast group that it knows nothing about, it hashes the group TA, selects the nearest RAP for the group, and transmits the packet to the RAP (path segment 2). The RAP receives the packet and selects a JAP based on whatever selection criteria is used for the group. For instance, if the criteria is proximity, it selects a JAP close to the IAP. The RAP forwards the packet to the selected JAP (path segment 3), and at the same time informs the IAP of the JAP (the RAP sends a list of JAPs, for failover purposes).

The use of RAPs unfortunately introduces another overlay hop in the path from client to target. We mitigate this cost however by having the IAP cache information about JAPs. Once the IAP has cached this information, subsequent packets (not only of this connection, but of subsequent connections too) are transmitted directly to the JAP. This is shown in the right-hand side of Figure 5.2. The time-to-live on this cache entry can be quite large. This is because the cache entry can be actively invalidated in one of two ways. First, if the target leaves the JAP, the JAP can inform the IAP of this when a subsequent packet arrives. Second, if the JAP disappears altogether, inter-AP monitoring will inform all APs of this event. In both cases, the IAP(s) will remove the cached entries, failover to other JAPs it knows of, or failing this, go back to the RAP. Because of this cache invalidation approach, the IAP does not need to go back to the RAP very often.

Note that in figure 5.2, the JAP is responsible for transmitting packets to and receiving packets from its targets. The reasoning for this is not obvious and goes as follows. We aim to support legacy clients that expect to see return packets

coming from the same address and port to which they sent packets. In general, targets cannot source packets from anycast addresses and so at least one proxy must be inserted into the target-client path. Furthermore, if NAT is being used to forward packets to the target, then the proxy with the NAT state should be the proxy that handles the return packets.

This might argue for traversing the IAP in the reverse direction too, since by necessity it must be traversed in the forward direction. The argument in favor of using the JAP however, boils down to the following two points. First, it is highly convenient to keep all target state in one proxy rather than two or more. Since the JAP in any event must monitor target health, it makes sense to put all target state in the JAP. Second, the JAP is close to the target, so the cost of traversing the JAP in terms of path length is minimal (Section 5.4.2). Also, by seeing packets pass in both directions, the JAP is better able to monitor the health of the target. For the most part, when a packet passes from client to target, the JAP may expect to soon see a packet in the reverse direction. Rather than force the JAP to continuously ping each target, the lack of a return packet can be used to trigger pings.

The use of proxies implies that the PIAS path (AC⇒IAP⇒ JAP⇒AT) might be longer than the direct path (AC⇒AT).[4] However, the proximity of the client to the IAP and of the target to the JAP should ensure that PIAS imposes minimal stretch and hence fulfills goal 3. This has been substantiated by simulating the stretch imposed by PIAS across a tier-1 topology map of the Internet.

The introduction of the RAP to allow scaling by the number of groups is somewhat equivalent to the extra round-trip imposed by application-level any-

[4]The PIAS path may actually be shorter as inter-domain routing is not optimal [8].

cast schemes, for instance in the form of the DNS lookup or the HTTP redirect. This is one aspect of PIAS that falls short of native IP Anycast, which has no such extra hop. Having said that, it would be possible for a small number of groups with minimal target churn to operate without RAPS—that is, to spread JAP information among all APs. This might be appropriate, for instance, for a CDN or for 6to4 gateways [26]. By-and-large, however, we can expect most groups to operate with RAPs as described here, and in the remainder of the design section, we assume that is the case.

5.3.3 Scale by group size and dynamics

If the only selection criteria used by a RAP to select a JAP were proximity to the client, then the RAP could ignore the number of targets reachable at each JAP. In order to load balance across targets, however, RAPs must know roughly how many targets are at each JAP. In this way, RAPs can select JAPs in a load balanced way, and each JAP can subsequently select targets in a load balanced way. Unfortunately, requiring that RAPs maintain counts of targets at JAPs increases the load on RAPs. This could be a problem for very large groups, or for groups with a lot of churn.

We mitigate this problem by allowing the JAP to give the RAP an approximate number of targets, for example within 25% or 50% of the exact number. For instance, if 25% error is allowed, then a JAP that reported 100 targets at one time would not need to report again until the number of targets exceeded 125 or fell below 75. This approach allows us to trade-off the granularity of load balancing for scalability with group size and dynamics. Indeed, this trade-off can

be made dynamically and on a per-group basis. A RAP that is lightly loaded, for instance, could indicate to the JAP that 100% accuracy reporting is allowed (i.e., in its acknowledgment messages). As the RAP load goes up, it would request less accuracy, thus reducing its load. The combination of the two-tiered approach with inaccurate information in a system with 2 groups is illustrated in Figure 5.3 (the figure assumes that there is just one RAP for each group). Section 5.4.1 presents simulations that show the benefits of this approach in the case of a large, dynamic group.

In any event, the number of targets is not the only measure of load. Individual targets may be more-or-less loaded due to differing loads placed by different clients. Ultimately, the JAP may simply need to send a message to the RAPs whenever its set of targets are overloaded for whatever reason.

5.3.4 Scale by number of proxies

Given that we have laid out the basic architecture of PIAS, we can now specifically look at PIAS deployment issues. A central question is, how many proxies may we reasonably expect in a mature PIAS deployment, and can we scale to that many proxies?

A key observation to make here is that the scaling characteristics of PIAS are fundamentally different from the scaling characteristics of IP routing. While the traffic capacity of the Internet can be increased by adding routers, the scalability of IP routing per se is not improved by adding routers. All routers must contain the appropriate routing tables. For instance, all Tier1 routers must contain the complete BGP routing table no matter how many Tier1 routers there are. For

the most part, IP routing is scaled by adding hierarchy, not adding routers.

With PIAS, on the other hand, scaling does improve by adding proxies. With each additional proxy, there are lower ratios of target-to-JAP and group-to-RAP. Growth in the number of groups and targets can be absorbed by adding proxies. However, an increase in the number of proxies presents its own scaling challenge. Among other things, every proxy is expected to know the up/down status of every other proxy.

The following describes a simple divide-and-conquer approach that can be used if the number of proxies grows too large. In a typical deployment, a given anycast service provider starts with one anycast prefix, and deploys proxies in enough geographically diverse POPs to achieve good proximity. As more anycast groups are created, or as existing anycast groups grow, the provider expands into more POPs, or adds additional proxies at existing POPs. With continued growth, the provider adds more proxies, but it also obtains a new address prefix (or splits the one it has), and splits its set of proxies into two distinct groups. Because the IP routing infrastructure sees one address prefix per proxy group, and because a proxy group can consist of thousands of proxies and tens of thousands of anycast groups, the provider could continue adding proxies and splitting proxy groups virtually indefinitely.

The size of a mature proxy deployment may be roughly calculated as follows. There are about 200 tier-1 and tier-2 ISPs [123]. An analysis of the ISP topologies mapped out in [117] shows that such ISPs have ~25 POPs on average. Assuming that we'd like to place proxies in all of these POPs, this leads to 5000 POPs. Assuming 3-4 proxies per POP (for reliability, dicussed later), we get a conservative total of roughly 20,000 proxies before the infrastructure can

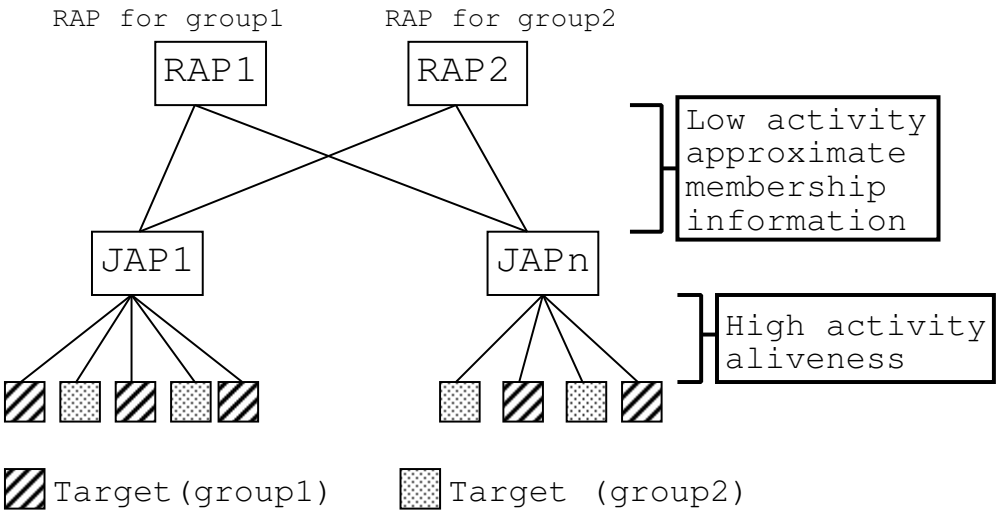


Figure 5.3: 2-tier membership management: the JAPs keep the aliveness status for the associated targets; the RAP for a group tracks the JAPs and an approximate number of targets associated with each JAP

be split.

While 20,000 proxies is not an outrageous number, it is large enough that we should pay attention to it. One concern not yet addressed is the effect of the number of proxies on IP routing dynamics. In particular, BGP reacts to route dynamics (flapping) of a single prefix by "holding down" that prefix—ignoring any advertisements about the prefix for a period of at most one hour [91]. A naive proxy deployment where each proxy advertises the anycast prefix directly into BGP would imply that a proxy failure necessitates a BGP withdrawal for the prefix (from the site where the proxy is located) that could lead to hold downs. While the proxy stability ensures that such events do not occur often, even the occasional prefix instability and the consequent service disruptions that a large proxy deployment would entail are not acceptable.

Hence, the deployment model involves more than one proxy being placed

inside every POP where the proxies are deployed. Such an arrangement is referred to as an ***anycast cluster*** and is based on the model used by the anycasted f-root server [145].[5] The approach involves connecting one or more routers and more than one proxy to a common subnet. All the proxies in the cluster advertise the anycast prefix into IGP while the routers advertise it into BGP and hence, a proxy-failure does not lead to a BGP withdrawal.

5.3.5 Proximity

The introduction of the proxies into the IP path negates the natural ability of native IP Anycast to find the nearest target. Therefore, we require explicit mechanisms in PIAS to regain this capability.

As mentioned before, native IP Anycast sets the client-IAP and target-JAP path segments. The RAP, on the other hand, selects the JAP, and therefore sets the IAP-JAP path segment (on forward packets) and the JAP-client path segment (on return packets). To ensure the proximity of the target to the client, the RAP must choose a JAP close to the IAP and hence, every AP must know the distance (in terms of latency) between every pair of APs. This could be accomplished using a system like Meridian [132] or a proximity addressing scheme like GNP [99] or Vivaldi [34].

Another possibility is to use a simple, brute-force approach, similar to what is used in RON [8]. Here, every AP occasionally pings every other AP and advertises the minimum measured round trip time (RTT) to all other APs. This is feasible because, with the cluster deployment approach, RAPs only need to

[5] Hereon, anycast cluster is referred to as proxy cluster or simply, cluster.

Table 5.1: Failover along the PIAS forward path (AC⇒IAP⇒JAP⇒AT) and reverse path (AT⇒JAP⇒AC)

Segment	Failure of	Failover through	Section
AC⇒IAP	IAP	IGP, onto a proxy within the same cluster	5.3.6
IAP⇒JAP	JAP	proxy health monitoring system	5.3.6
JAP⇒AT	AT	pings between target and JAP, passive monitoring by JAP	5.3.1,5.3.2
AT⇒JAP	JAP	pings routed to a different proxy who becomes JAP	5.3.6
JAP⇒AC	AC	no failover needed	-

know the distance between each pair of clusters. While validating the above claim would require experimentation with the actual deployment, back of the envelope calculations do paint a promising picture for the simple approach.

5.3.6 Robustness and fast failover

The introduction of proxies between client and target might have a negative impact on the robustness of PIAS as compared to native IP Anycast. On the other hand, RON [8] has shown how an overlay structure can be used to improve the resiliency of communication between any two overlay members. Extending the same thought, PIAS, by ensuring the robustness of packet traversal through the proxy overlay, can improve the resiliency of communication between clients and group members. We believe that given the stable nature of the proxies, their deployment in well connected parts of the Internet (tier-1 and tier-2 ISPs) and the engineering that would go into their set-up, PIAS should be able to match,

if not better, the robustness offered by native IP Anycast.

A related requirement is that of fast fail-over. ”E2E” native IP anycast has to achieve failover when a group member crashes, so that clients that were earlier accessing this member are served by some other group member. Given the way native IP Anycast works, this failover is tied to IP routing convergence. Our measurements show that a planned native IP Anycast deployment can achieve fast failover. Since PIAS uses native IP anycast to reach the proxies, it is subject to the same issues and would require planned proxy deployment. The process of overcoming the failure of a proxy is termed as **proxy failover**. In addition, the proxies must themselves be able to fail over from one target to another which is termed as **target failover**. Thus the failover problem seems worse with PIAS than with native IP Anycast; however, this is not the case.

Target failover

As discussed in Sections 5.3.1 and 5.3.2, the JAP is responsible for monitoring the aliveness of its targets. It does this through pinging and tracking data packets to and from the target. The JAP is also responsible for directing IAPs to delete their cache entries when enough targets have failed.

Proxy failover

There is still the question of clients failing over onto a different proxy when their IAP crashes, and targets failing over when their JAP crashes. And there are two levels at which this must be achieved: at the routing level and at the overlay level.

At the routing level, the system must be engineered such that when a proxy fails, clients that were using this proxy as an IAP are rerouted to some other proxy quickly. PIAS's deployment of proxies in a cluster means that this failover is across proxies within the same cluster. Also, since the proxies advertise the prefix into IGP, PIAS relies on IGP for convergence after a proxy failure and hence can achieve faster failover. Typically, this is of the order of a few seconds and can be reduced to sub-second times [6].

At the overlay level, to monitor the health of proxies, we use a 2-tier health monitoring system. At the first tier, the proxies within the same proxy cluster are responsible for monitoring each other. At the next level, each proxy in a cluster monitors the health of a small number of other clusters. When either an individual proxy or an entire cluster fails, it is detected quickly and communicated to all remaining proxies.

Section 5.3.2 had described IAP behavior when a JAP goes down. The only thing left to discuss is target behavior when a JAP goes down. In this case, IGP convergence will cause ping packets from the target to reach another proxy in the same cluster as the failed JAP. The proxy will ask the target to re-register. Table 5.1 sums up the way PIAS achieves failover across various segments of the client-target path.

5.3.7 Target selection criteria

As described earlier, the RAP may select the JAP based on a number of criteria, including proximity, load balancing, and connection affinity. The JAP subsequently selects a target. It is this selection process, divorced from IP routing,

that allows PIAS to offer richer target selection criteria.

How PIAS achieves load balance and proximity has already been discussed. Connection affinity is discussed later in this section. We wish to point out here that these three important selection criteria are in fact at odds with each other. For example, if both load balance and proximity are important criteria, and the JAP nearest to the IAP is heavily loaded, then one of the other criteria must be compromised. This basic set of trade-offs applies to application-level anycast as well.

By never selecting the source of a packet as the target, PIAS allows a host to be both a target and a client for a given group. Packets sent by the target to the group address would be forwarded to some group target other than the sender. Note that this is not possible with native IP Anycast and it allows PIAS to support new P2P applications (section 5.6.1).

Proxies could potentially base their target selection on various scoping criteria. These selection criteria can be expressed by overloading the transport address, i.e. a group can have separate TAs for each type of scoping. For instance, an anycast packet could be administratively scoped. That is, it could indicate that the target should be in the same site, belong to the same DNS domain, or have the same IP address prefix (or be from different sites, DNS domains, or IP prefixes). While how this would be configured and operated is a good topic for further study, the selection functionality of the RAP allows for the possibility of many such features.

Another form of selection would be to pick a random target rather than the nearest target - the RAP would pick a random JAP who would then pick a ran-

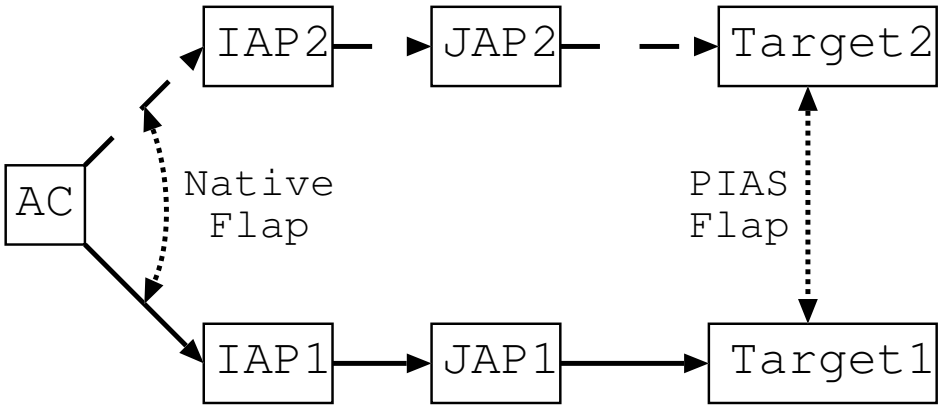


Figure 5.4: Lack of native IP Anycast affinity can cause flaps in the PIAS model

dom target. Random selection among a group can be useful for various purposes such as spreading gossip [54] or selecting partners in multicast content distribution [82]. Indeed, in the PIAS architecture, there is no reason an anycast packet cannot be replicated by the RAP and delivered to a small number of multiple targets. The salient point here is that once IP Anycast functionality is divorced from IP routing, any number of new delivery semantics are possible if the benefits justify the cost and complexity.

Connection affinity

Lack of connection affinity in native IP Anycast has long been considered one of its primary weak points. This issue spills over into PIAS. Specifically, the issue is how to maintain affinity when native IP Anycast causes a different IAP to be selected during a given client connection. If the same IAP is always used, then packets will be sent to the same JAP that was initially cached by the IAP. However, a change in the IAP could lead to a change in the target the packets are delivered to, as shown by Figure 5.4. Application-layer anycast doesn't have

this problem, because it always makes its target selection decision at connection start time, and subsequently uses unicast.

A simple solution would be to have RAPs select JAPs based on the identity of the client, such as the hash of its IP address. This way, even if IP routing caused packets from a given client to select a different IAP, they would be routed to the same JAP and hence the same target. Unfortunately, this approach completely sacrifices proximity and load balance. Broadly, another approach would be to modify the host application by making it anycast aware, and redirect the host to the unicast address of a selected target (either PIAS or the target itself could do this redirect). There are some security issues here—the redirect must be hard to spoof—but these are surmountable.

We can also imagine complex schemes whereby JAPs and IAPs coordinate to insure affinity. Alternatively, it is possible to circumvent the problem by ensuring that flaps between anycast group members don't impact the client operation. For instance, Trickles [115] proposes a TCP-like transport protocol wherein connection state is maintained at the clients and hence, would obviate the need for route stability for connection-oriented applications running on PIAS. Similarly, Al-Qudah et. al. [5] propose a lightweight mechanism that handles session disruptions due to anycast flaps and allows clients to continue normal operation.

However, the need for such mechanisms and their complexity depends on a fundamental question: how good or bad is the affinity offered by native IP Anycast? Our measurement results show that the affinity offered by native IP Anycast is very good. For instance, 95% of clients we measured suffer less than one flap per day. For these clients, the probability that a minute-long connection

breaks due to a flap is less than 1 in 1440 and the probability that an hour-long connection breaks is less than 1 in 24. Note that it is the short connections that, in order to avoid the overhead of anycast to unicast redirect, need to rely on anycast affinity. Long connections can incur the overhead of a redirect and hence, could use anycast for discovery and unicast for the actual communication.

These numbers paint an encouraging picture; the probability that a short connection breaks due to a routing flap is very small compared to the probability of the connection breaking due to other factors. Further, many applications involving long connections today account for the possibility that the connection might break and have built-in mechanisms to pick up where they left off. These factors imply that IP Anycast affinity is quite good and for most applications, PIAS should be able to avoid using the aforementioned mechanisms that obviate the anycast affinity requirement.

5.4 Evaluation

In this section we evaluate the PIAS architecture using simulations and measurements. Sections 5.4.1 and 5.4.2 present simulation results that show the scalability (by group characteristics) and the efficiency of the PIAS deployment while section 5.4.3 discusses our PIAS implementation.

5.4.1 Scalability by group size and dynamics

In this experiment, we evaluate PIAS's ability to handle large and dynamic groups (as described in 5.3.3). We simulate the load imposed by a large group

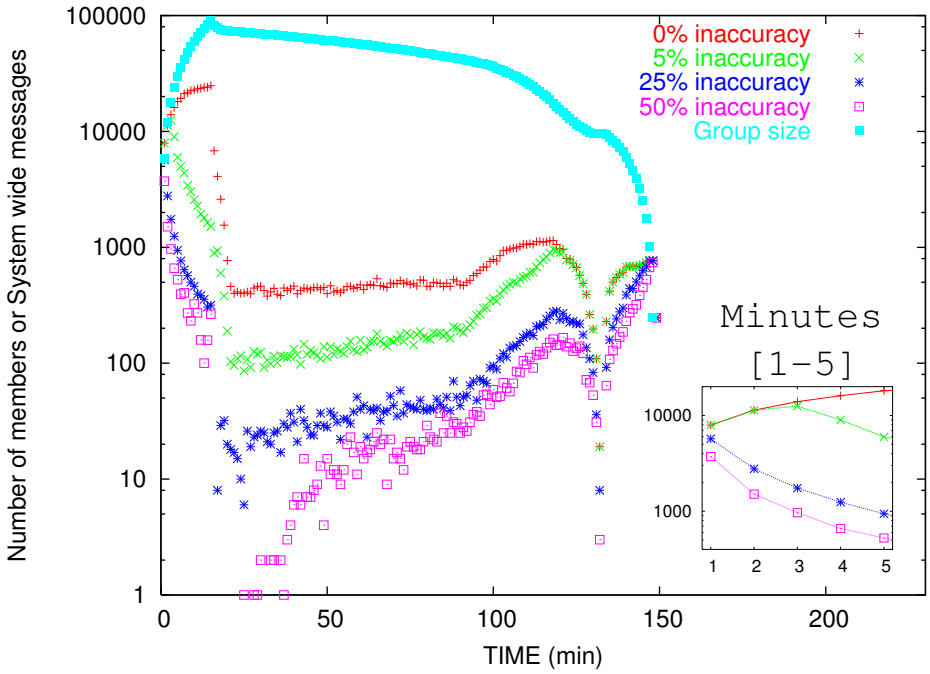


Figure 5.5: System-wide messages from the all the JAPs to the 4 RAPs during the event for varying degrees of inaccuracy

with high churn on the proxy infrastructure. The dynamics of the simulated group - the arrival rate of group members and the session duration cumulative distribution function - resemble the dynamics of the largest event observed in a study of large-scale streaming applications [119]. Simulation of just one such group is sufficient as the load imposed varies linearly with the number of such groups supported.

The PIAS infrastructure in the simulation has varying number of proxies and maximum group size. We simulate four RAPs per group. We want to measure the number of messages required to keep the 2-tier membership hierarchy updated in face of the group dynamics. This is the number of messages from the JAPs of the group to the 4 RAPs and is referred to as **'system-wide messages'**.

Figure 5.5 plots the system-wide messages produced with a proxy deployment of size 1000 and the group size bounded by 90000. The topmost curve in the figure shows how the group size varies with the time. A flash crowd, at a rate of ~100 members/second, leads to a sudden rise in the group size in the first 10 minutes. The other curves plot the number of messages produced in the corresponding minute (as plotted along the X-axis) for varying degrees of inaccuracy. The degree of inaccuracy, as explained in section 5.3.3, implies that a JAP only informs a RAP of a change in the number of members associated with it if the change is more than a certain percentage of the last value sent.

The inaccuracy of information offers only a small benefit in the nascent stages of the group (the first minute). This is because no matter what inaccuracy percentage we use, the JAP must inform the RAP of the first group member that contacts it. In the next couple of minutes, as the group increases in size and more members join their corresponding JAPs, the inaccuracy causes the traffic towards the 4 RAPs to drop rapidly (see the embedded graph in figure 5.5). Overall, the average number of messages over the duration of the entire event reduces from 2300 per min. with the naive approach to 117 per min. with 50% inaccuracy.

Figure 5.6 plots the average system-wide messages (per second) versus the percentage of inaccuracy for varying number of proxies and varying maximum group size. Each plotted point is obtained by averaging across 20 runs. All curves tend to knee around an inaccuracy mark of 50%−60%. The closeness of the curves for different sized groups (given a fixed number of proxies) points to the scalability of the system by the group size even in the face of high churn.

More interesting is the variation of the load on the RAPs with the number

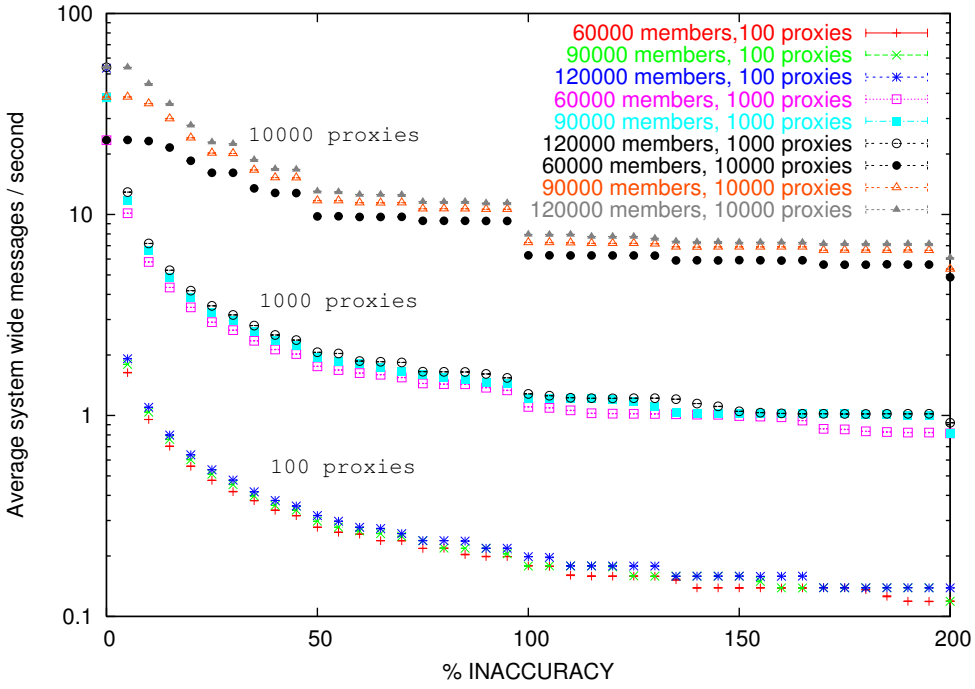


Figure 5.6: Average system-wide messages (per second) versus the percentage of inaccuracy with varying number of proxies and varying maximum group size.

of proxies. As the number of proxies increase, the number of JAPs increase; an offshoot of the assumption that the group members are evenly distributed across the proxy infrastructure. For a given group size, each JAP is associated with lesser number of group members. Hence, there is lesser benefit due to the inaccuracy approach. This shows up as the increase in the average number of messages directed towards the RAPs with the number of proxies.

The figure shows that such an extreme group in a 100 proxy deployment with 100% inaccuracy would require an average of ~0.18 messages/second. As a contrast the same setup in a 10000 proxy deployment would necessitate an average of ~7.25 messages/second. The low message overhead substantiates the PIAS scalability claim. Note that a larger number of proxies implies that

each proxy is a RAP for a smaller number of groups. The number of targets associated with each proxy (as a JAP) reduces too. Thus, increasing the number of proxies would indeed reduce the overall load on the individual proxies.

5.4.2 Stretch

PIAS causes packets to follow a longer path (client ⇒ IAP ⇒ JAP ⇒ target). We have argued that the combination of native IP anycast and proxy-to-proxy latency measurements minimizes the effect of this longer path. This section simulates the stretch introduced by PIAS along the end-to-end path.

For the simulation, we use a subset of the actual tier-1 topology of the Internet, as mapped out in the Rocketfuel project [117]. This subset consists of 22 ISPs, 687 POPs, and 2825 inter-POP links (details in [138]). The use of only the tier-1 topology can be justified on two grounds. First, a large proportion of traffic between a randomly chosen client-target pair on the Internet would pass through a tier-1 ISP. Second, such a simulation gives us an approximate idea about the overhead that a PIAS deployment restricted to tier-1 ISPs would entail.

The topology was annotated with the actual distance between POPs (in Kms) based on their geographical locations. We then used SSFNET [120] to simulate BGP route convergence. This allowed us to construct forwarding tables at each of the POPs and hence, determine the forwarding path between any two POPs.

The simulated PIAS deployment involves placing a variable number of proxies at random POPs, one proxy per POP. These POPs are referred to as the *proxy*

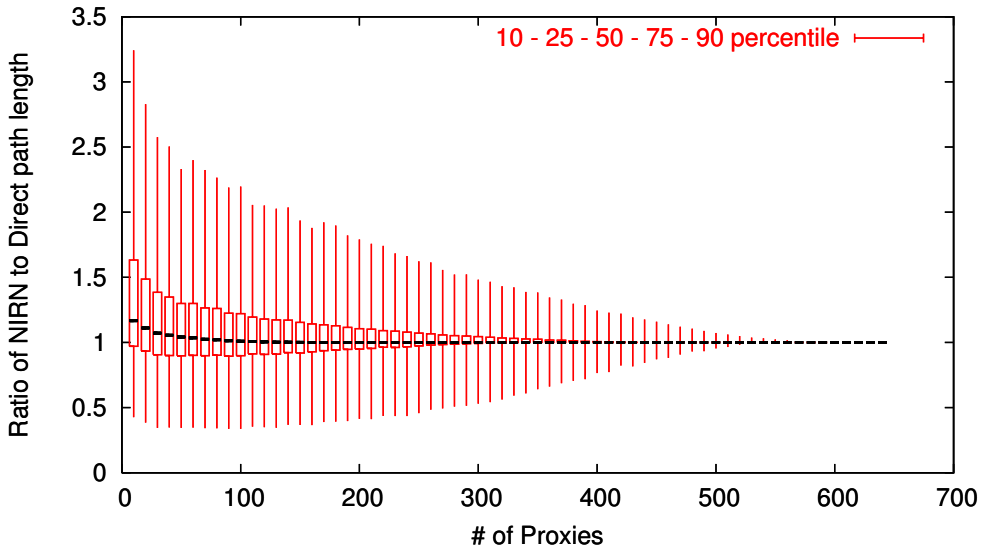


Figure 5.7: Percentiles for the stretch with varying number of proxies

POPs. For every client-target pair to be simulated, we choose a POP through which the client's packets enter the topology (the *client POP*) and a POP through which the target's packets enter the topology (the *target POP*). The forwarding paths between the client and the target through these POPs represents the *direct path*. The IAP is assumed to be in the *proxy POP* closest to the client POP—this is the *IAP POP*. Similarly, the JAP is in the *proxy POP* closest to the target POP—this is the *JAP POP*. The *PIAS path* comprises of the following three segments: from the *client POP* to the *IAP POP*, from the *IAP POP* to the *JAP POP* and from the *JAP POP* to the *target POP*.

Figure 5.7 plots the percentiles for the stretch with varying number of proxies. For a given number of proxies, we simulated 100000 runs. Each run comprised of simulating a client-target pair and finding the *direct* and the *PIAS path* length (in kms). Note that the well-documented non-optimal nature of inter-domain routing [8] is reflected in the cases where the PIAS path turns out to be shorter than the direct path. The figure shows that with a deployment of just

100 proxies (a mature deployment might encompass 50 times more POPs), the median stretch is 1.01 with the 90^{th} percentile being 2.2. Hence, even with a small size deployment, PIAS performs well with regards to the direct path.

5.4.3 Implementation

We have implemented the PIAS system. The current implementation of PIAS proxies comprises of a user-space component responsible for the overlay management tasks, such as handling proxy failures, target join/leaves, health monitoring etc. and a kernel-space component responsible for the actual forwarding of packets through the use of Netfilter hooks [98]. This involves tunneling of the packets when sending them between 2 proxy nodes, and using a NAT when handling packets to/from a target.

We deployed a simplified (and completely user-space version) of the PIAS system detailed above on the five node internal-anycast deployment described in section 4.3.

5.5 Related work

Table 5.2 summarizes the pros and cons of PIAS, application level anycast, and other related approaches described below.

Partridge et. al. [102] originally proposed the IPv4 anycast service. It involves assigning an otherwise unicast IP address IP_A to multiple hosts, and advertising it into the routing infrastructure from all the hosts. Packets addressed

Table 5.2: The Anycast Design Space

Criterion (related to goal number)	IPv4	IPv6	IP + GIA	App. Level	*i*3	PIAS
Router Modification(1)	No	No	**Yes**	No	No	No
Client Modification(1)	No	No	No	No	**Yes**	No
Scalability by group size(2)	Very Good	Very Good	Very Good	**Poor**	**Poor/Good**[6]	Good
Stretch(3)	No	No	Little/No	No	Little	Little
Robustness(4)	No Issues	No Issues	No Issue	Mixed	Mixed	Mixed[7]
Failover(5)	Fast[8]	Fast[12]	Fast[12]	Fast	Fast	Fast
Target Deployment(6)	**Difficult**	**Difficult**	**Difficult**	Easy	Easy	Easy
Scalability by no. of groups(7)	**No**	**No**	Yes	Yes	Yes	Yes
Scalability by group dynamics(8)	**Poor**	**Poor**	**Poor**	**Poor**	**Poor/Good**[10]	Good
Cost of Proximity(9)	None	None	Small	**Large**	**Large**	Small
Low-level access	Yes	Yes	Yes	**No**	Yes	Yes

to IP_A will be forwarded to the host nearest to the packet source in terms of metrics used by the routing protocol. Later, IPv6 incorporated anycast into its addressing architecture [64]. It allowed for scoped anycast addresses for groups confined to a topological region, which does not burden the global routing system. However, a globally spread group still poses scalability problems. Besides, IPv6 anycast also inherits all the other limitations of IPv4 anycast. Despite the shortcomings, there has been work detailing the relevance of anycast as a tool for service discovery and other applications, both for IPv4 [22] and for IPv6 [93].

Katabi and Wroclawski [75] proposed an architecture that allows IP Anycast to scale by the number of groups. Their approach is based on the observation that services have a skewed popularity distribution. Hence, making sure that the unpopular groups do not impose any load on the routing infrastructure addresses the scalability issue. However, the need to change routers puts a severe dent on the practical appeal of the approach. Besides, being a router-based approach, it suffers from most other limitations of IPv4 anycast.

Because of the limitations of these approaches, anycast today is typically implemented at the application layer. This offers what is essentially anycast service

discovery—DNS-based approaches use DNS redirection while URL-rewriting approaches dynamically rewrite the URL links as part of redirecting a client to the appropriate server. Related proposals in the academic community include [47,53,137]. The idea behind these is to identify the group using an application level name that, at the beginning of the communication, is mapped to the unicast address of a group member. The reliance on unicast support from the underlying IP layer implies that these approaches circumvent all limitations of IP Anycast. The challenge here is to collect the relevant selection metrics about the group members in an efficient and robust fashion.

Another element in this design space is anycast built on top of the indirection architecture offered by i3 [121]. i3 uses identifiers as a layer of indirection that generically gives the receiver tremendous control over how it may (or may not) be reached by senders. One of the services i3 can provide is anycast. There are two main advantages of PIAS over i3 for the anycast service. First, PIAS requires no changes in the protocol stack, whereas i3 requires a new layer inserted below transport. A PIAS client, on the other hand, can use PIAS with no changes whatsoever. Second, because PIAS uses native IP Anycast, it is easier to derive proximity from PIAS than from i3. PIAS only has to measure distances between proxies—i3 has to measure distances to clients and targets. The main advantage of i3 over PIAS is that it is easier to deploy an i3 infrastructure than a PIAS infrastructure, precisely because i3 doesn't require IP Anycast. Indeed, this has been a source of frustration for us—we can't just stick a PIAS proxy on Planetlab and start a service.

[10]Note that the way i3 has described their anycast, it wouldn't scale to very large or very dynamic groups, because a single node holds all the targets and receives pings from the targets. It may be possible that i3 could achieve this with a model closer to how they do multicast, but we're not sure.

[11]For reasons described in first paragraph of section 5.3.6.

[8]They can be engineered to be fast by relying on IGP for convergence.

As far as the broader notion of indirection is concerned, there is no question that i3 is more general. Its ability for both the sender or receiver to chain services is very powerful. The addressing space is essentially infinite, and hosts can create addresses locally. Finally the security model (that supports the chaining) is elegant and powerful. Having said that, PIAS does provide indirection from which benefits other than just anycast derive. For unicast communications, it could be used to provide mobility, anonymity, DoS protection, and global connectivity through NATs. In the best of all worlds, we'd want something like i3 running over PIAS. But IPv6 and NAT have taught us that you don't always get the best of all worlds, and considering PIAS's backwards compatibility, it may after all be the more compelling story.

5.6 Anycast applications

Given that PIAS offers an easy-to-use global IP Anycast service that combines the positive aspects of both native IP Anycast and application-layer anycast, it is interesting to consider new ways in which such a service could be used.

5.6.1 Peer Discovery

Though IP Anycast has long been regarded as a means of service discovery, this has always been in the context of clients finding servers. PIAS opens up discovery for P2P networks, where not only is there no client/server distinction, but peers must often find (and be found by) multiple peers, and those peers can

come and go rapidly. Examples of such cases include BitTorrent and network games.

One reason that traditional IP Anycast has not worked for peer discovery (other than difficulty of deployment), is that an IP anycast group member cannot send to the group—packets are just routed back to themselves. With the right selection characteristics, PIAS can support a wide-range of P2P applications. Random selection would allow peers to find arbitrary other peers, and is useful to insure that unstructured P2P networks are not partitioned. Proximity is obviously also important, but to insure that a peer can find multiple nearby peers (rather than the same peer over and over), a selection service whereby a node can provide a short list of targets to exclude (i.e. already-discovered targets) could be used.

5.6.2 Reaching an Overlay network

A very compelling application of PIAS would allow a RON [8] network to scale to many thousands of members, and would allow those members to use RON not only for exchanging packets with each other, but with any host on the Internet! What follows is a high-level description of the approach. Assume a set of 50-100 RON "infrastructure" nodes that serve many thousands of RON clients. The RON nodes all join a large set of anycast groups—large enough that there is an anycast transport address (TA) for every possible client connection. The RON nodes also partition the anycast TAs so that each TA maps to a single RON node. Clients discover nearby RON nodes (or a couple of them) using one of the anycast groups, and establish a unicast tunnel (for instance, a VPN tunnel) with

the RON node. We call this the *RON tunnel*, and the RON node is referred to as the *local RON*.

When a client wishes to establish a connection with some remote host on the Internet, it does so through its RON tunnel. The local RON assigns one of its TAs to the connection using NAT, and forwards the packet to the remote host. When the remote host returns a packet, it reaches a nearby RON node, called the *remote RON*. Because the transport address of the return packet maps to the local RON node, the remote RON node can identify the local RON node. The remote RON tags the packet with its own identity, and transmits the packet through the RON network to the local RON node, which caches the identity of the remote RON, and delivers the packet to the client. Now subsequent packets from the client to the remote host can also traverse the RON network.

This trick isn't limited to RONs. It could also work for route optimization in Mobile IP (for v4 or v6, see [96] for a description of the problem), or simply as a way to anonymize traffic without sacrificing performance.

5.7 Discussion

In this chapter, we have presented the basic aspects of PIAS. A "practical" IP Anycast service, however, requires a number of features that we don't describe in detail. For example, the need for scoping whereby packets from clients in a domain (enterprise) are always served by targets within the domain. This can be achieved by deploying a PIAS proxy in the domain, or simply by deploying intra-domain native IP Anycast. We comment on a few other concerns below.

Security. An important issue regarding PIAS is security. The IP routing infrastructure is secured router-by-router through human supervision of router configuration. This makes routing security error-prone and unreliable. Since PIAS involves advertising a prefix into inter-domain routing, it is afflicted by the same issues. However, it is important to note that PIAS does not worsen the situation. Also, because an anycasted autonomous system is akin to a multi-homed autonomous system from the routing point of view, any future solution for routing security would apply directly to the PIAS deployment.

PIAS, however, does need to explicitly secure its *join* and *leave* primitives. Because these primitives are to be used by group members who have an explicit contract with the anycast service provider, we could use standard admission control schemes; for example, PIAS could adapt any of a number of network or wireless authentication protocols like EAP [3]. Previous work on using overlays to protect specific targets from DOS attacks [78] described some approaches to allow controlled access to the overlay.

Proximity. An assumption implicit in PIAS's claim of incurring minimal stretch (section 5.4.2) is the proximity of the client to the IAP and of the server to the JAP. This assumption is justified because these relations are discovered using native IP Anycast and hence, the distances are small in terms of metrics used by inter-domain routing. However, this does not necessarily imply that the distances are small in terms of latency. As a matter of fact, our measurements detailed in Section 4.5 show that naive IP Anycast deployments can suffer from poor proximity. However, we also found that planned native IP Anycast deployments can ensure that clients are directed to a close-by server. Hence, a planned PIAS deployment can achieve minimal stretch.

Perhaps a more important and common situation would be one whereby two targets were on the same LAN. This would be the case for instance at a gaming LAN party. It would obviously be nice if the game could automatically determine if other gamers were on the local LAN, and if not, then discover peers over the Internet. Without getting into details, it is easy to envision an approach whereby a PIAS TA could be algorithmically mapped into an IP multicast address and UDP port. PIAS targets would try this local IP multicast-based discovery before falling back on global PIAS.

5.8 Summary

We propose a proxy based IP anycast service that addresses most of the limitations of native IP Anycast. Specifically, this chapter presents the design of PIAS, a practically deployable IP Anycast architecture. The unique features of PIAS such as the scalability by the size and dynamics of groups mean that it opens up new avenues of anycast usage. The purported scalability has been substantiated through simulations representing extreme, but real, workloads. Simulations on the real tier-1 topology of the Internet point to the efficiency of our approach.

The fact that PIAS uses native IP Anycast means that it can be used as a simple and general means of discovery and bootstrapping. Internet measurements detailed in the previous section show that the reliance on native IP Anycast does not undermine PIAS's ability to support connection oriented services. A PIAS prototype has been built and deployed. We feel confident that PIAS has the potential of fulfilling the need for a generic Internet-wide anycast service that can serve as a building block of many applications, both old and new.

CHAPTER 6

CONCLUDING REMARKS

We take the stand that dirty-slate solutions should be an important part of the networking research agenda. Many network problems arise due to a mismatch between design and use and hence, any solution needs to address this. We hypothesize that the ability to create virtual networks through the use of tunnels represents an opportunity to mend such mismatches and hence, alleviate network problems. And the rapidly growing use of tunnels in the Internet entails that such tunnel-driven solutions have a much lower bar for deployment. In this thesis, we illustrate our argument by harnessing tunnels to tackle two important problems facing the Internet.

The growing size of the Internet routing table poses vexing challenges, both technical and business, for Internet ISPs. Consequently, reducing the routing table size and improving routing scalability has been an extremely well-researched problem area with several architectural fixes proposed. Instead of focusing on reducing the size of global routing table, we note that important immediate gains can be made by reducing the routing state maintained by individual routers. Guided by this, the ViAggre proposal presented in this thesis uses tunnels to create virtual networks within a given ISP's network that are responsible for a fraction of the IPv4 address space. This ensures that individual routers only need to keep a fraction of the global routing table and is shown to significantly increase router lifetimes. Further, ViAggre is a dirty-slate solution since it does not require changes to network infrastructure and can be employed by ISPs today. This has allowed it to have real world impact, both in terms of standardization and industry interest. While not a complete solution,

we think that ViAggre provides ISPs with an attractive alternative to upgrading their routers.

IP Anycast is a one-to-any communication primitive supported by the Internet today. It derives several advantages by operating at the network layer, foremost amongst which is its robustness. It is these advantages that have led to IP Anycast being used for the transparent replication of the DNS root servers. However, the very fact that IP Anycast operates at the routing layer has also been its Achilles heel and has limited its deployment to a few infrastructure services. Based on this observation, the PIAS architecture uses a virtual network of proxies to offer anycast service to clients while the proxies themselves rely on native IP Anycast. This separates the anycast service from the routing infrastructure, while amortizing the deployment effort across all the anycast groups that the PIAS infrastructure supports. PIAS is also a dirty-slate solution with no changes required to network infrastructure – hardware or software. Consequently, we were able to deploy PIAS and have shown that it can be used as a simple and general means of discovery and bootstrapping on the Internet.

Looking ahead, we hope that ViAggre can help with the Internet's routing pains in the short to medium term till more fundamental, long-term changes can be agreed upon and deployed in the Internet. Similarly, we believe that a decent-sized PIAS deployment has the potential to serve as a general anycast service across the Internet. These possibilities are evidence of the fact that there are still significant gains to be made through dirty-slate network solutions.

BIBLIOGRAPHY

[1] J. Abley. Hierarchical Anycast for Global Service Distribution. ISC Technical Note ISC-TN-2003-1 www.isc.org/tn/isc-tn-2003-1.html, March 2003.

[2] Joe Abley. A Software Approach to Distributing Requests for DNS Service Using GNU Zebra, ISC BIND 9, and FreeBSD. In *Proc. of USENIX Annual Technical Conference*, June 2004.

[3] B. Aboba, L. Blunk, J. Vollbrecht, J. Carlson, and H. Levkowetz. RFC 3748, Extensible Authentication Protocol (EAP), June 2004.

[4] Akamai Technologies Inc. Internet Bottlenecks: the Case for Edge Delivery Services, October 2000. www.akamai.com/en/resources/pdf/whitepapers/Akamai_Internet_Bottlenecks_Whitepaper.pdf.

[5] Zakaria Al-Qudah, Seungjoon Lee, Michael Rabinovich, Oliver Spatscheck, and Jacobus Van der Merwe. Anycast-Aware Transport for Content Delivery Networks. In *Proc. of World Wide Web Conference*, April 2009.

[6] Cengiz Alaettinoglu and Stephen Casner. Detailed Analysis of ISIS Routing Protocol on the Qwest Backbone. NANOG 24 meeting, February 2002. www.nanog.org/meetings/nanog24/presentations/cengiz.pdf.

[7] Hussein Alzoubi, Seungjoon Lee, Michael Rabinovich, Oliver Spatscheck, and Jacobus Van der Merwe. Anycast CDNs Revisited. In *Proc. of World Wide Web Conference*, April 2008.

[8] David Andersen, Hari Balakrishnan, Frans Kaashoek, and Robert Morris. Resilient Overlay Networks. In *Proc. of ACM Symposium on Operating Systems Principles*, October 2001.

[9] L. Andersson, I. Minei, and B. Thomas. RFC 5036, LDP Specification, Jan 2006.

[10] Hitesh Ballani. Internal Anycast Service deployment, May 2006. http://pias.gforge.cis.cornell.edu/deployment.php.

[11] Hitesh Ballani, Yatin Chawathe, Sylvia Ratnasamy, Timothy Roscoe, and Scott Shenker. Off by Default! In *Proc. of Workshop on Hot Topics in Networks (HotNets)*, November 2005.

[12] Hitesh Ballani, Andrey Ermolinskiy, Sylvia Ratnasamy, and Paul Francis. An Experiment In Deploying Next Generation Network Protocols. Technical report, Cornell University, May 2006.

[13] Hitesh Ballani and Paul Francis. Towards a Deployable IP Anycast Service. In *Proc. of USENIX Workshop on Real, Large Distributed Systems (WORLDS)*, December 2004.

[14] Hitesh Ballani and Paul Francis. Towards a Global IP Anycast Service. In *Proc. of ACM SIGCOMM*, August 2005.

[15] Hitesh Ballani and Paul Francis. Understanding IP Anycast. Technical report, Cornell University, May 2006.

[16] Hitesh Ballani and Paul Francis. CONMan: A Step Towards Network Manageability. In *Proc. of ACM SIGCOMM*, August 2007.

[17] Hitesh Ballani, Paul Francis, Tuan Cao, and Jia Wang. ViAggre: Making Routers Last Longer! In *Proc. of Workshop on Hot Topics in Networks (HotNets)*, October 2008.

[18] Hitesh Ballani, Paul Francis, Tuan Cao, and Jia Wang. Making Routers Last Longer with ViAggre. In *Proc. of USENIX/ACM Symposium on Networked Systems Design and Implementation (NSDI)*, April 2009.

[19] Hitesh Ballani, Paul Francis, and Sylvia Ratnasamy. A Measurement-based Deployment Proposal for IP Anycast. In *Proc. of ACM SIGCOMM Conference on Internet Measurement*, October 2006.

[20] Piet Barber, Matt Larson, Mark Kosters, and Pete Toscano. Life and Times of J-Root. NANOG 32 meeting, October 2004. `http://www.nanog.org/mtg-0410/kosters.html`.

[21] Paul Barford. Wisconsin Advanced Internet Laboratory (WAIL), Dec 2007. `http://wail.cs.wisc.edu/`.

[22] E. Basturk, R. Haas, R. Engel, D. Kandlur, V. Peris, and D. Saha. Using IP Anycast For Load Distribution And Server Location. In *Proc. of IEEE Globecom Global Internet Mini Conference*, November 1998.

[23] Andy Bavier, Nick Feamster, Mark Huang, Jennifer Rexford, and Larry Peterson. In VINI Veritas: Realistic and Controlled Network Experimentation. In *Proc. of ACM SIGCOMM*, September 2006.

[24] Peter Boothe and Randy Bush. Anycast Measurements Used to Highlight Routing Instabilities. NANOG 34 meeting, May 2005. `http://www.nanog.org/mtg-0505/boothe.html`.

[25] Matthew Caesar, Donald Caldwell, Nick Feamster, Jennifer Rexford, Aman Shaikh, and Jacobus van der Merwe. Design and Implementation of a Routing Control Platform. In *Proc. of USENIX/ACM Symposium on Networked Systems Design and Implementation (NSDI)*, May 2005.

[26] B. Carpenter and K. Moore. RFC 3056, Connection of IPv6 Domains via IPv4 Clouds, February 2001.

[27] Martin Casado, Tal Garfinkel, Aditya Akella, Michael Freedman, Dan Boneh, Nick McKeown, and Scott Shenker. SANE: A Protection Architecture for Enterprise Networks. In *Proc. of Usenix Security*, August 2006.

[28] R. Chang and M. Lo. Inbound Traffic Engineering for Multi-homed AS's Using AS Path Prepending. *IEEE Network*, pages 18–25, March 2005.

[29] E. Chen and T. Bates. RFC 1998, An Application of the BGP Community Attribute in Multi-home Routing, August 1996.

[30] Yang Chu, Sanjay Rao, Srinivasan Seshan, and Hui Zhang. Enabling Conferencing Applications on the Internet Using an Overlay Muilticast Architecture. In *Proc. of ACM SIGCOMM*, August 2001.

[31] Brent Chun, David Culler, Timothy Roscoe, Andy Bavier, Larry Peterson, Mike Wawrzoniak, and Mic Bowman. PlanetLab: An Overlay Testbed for Broad-Coverage Services. *ACM SIGCOMM Computer Communication Review*, 33(3), July 2003.

[32] Danny Cohen and Ed Taft. IEN 140, Mutual Encapsulation of Internetwork Protocols, April 1980.

[33] Lorenzo Colitti. Effect of Anycast on K-root. DNS-OARC Workshop, July 2005. http://www.ripe.net/info/ncc/presentations/anycast-kroot.pdf.

[34] Frank Dabek, Russ Cox, Frans Kaashoek, and Robert Morris. Vivaldi: A Decentralized Network Coordinate System. In *Proc. of ACM SIGCOMM*, August 2004.

[35] E. Davies and A. Doria. Analysis of Inter-Domain Routing Requirements and History. Internet Draft draft-irtf-routing-history-07.txt, January 2008.

[36] Suran de Silva. 6500 FIB Forwarding Capacities. NANOG 39 meeting, February 2007. http://www.nanog.org/mtg-0702/presentations/fib-desilva.pdf.

[37] S. Deering and R. Hinden. IPv6 Metro Addressing. Internet Draft draft-deering-ipv6-metro-addr-00.txt, March 1996.

[38] Steve Deering. The Map & Encap Scheme for Scalable IPv4 Routing with Portable Site Prefixes, March 1996. http://www.cs.ucla.edu/~lixia/map-n-encap.pdf.

[39] Nandita Dukkipati, Nick McKeown, and Alexander G. Fraser. RCP-AC: Congestion Control to Make Flows Complete Quickly in Any Environment. In *Proc. of High-Speed Networking Workshop: The Terabits Challenge*, April 2006.

[40] E. Rosen and Y. Rekhter. RFC 2547, BGP/MPLS VPNs, March 1999.

[41] G. Montenegro (Editor). RFC 3024, Reverse Tunneling for Mobile IP, January 2001.

[42] W. Fang and L. Peterson. Inter-AS Traffic Patterns and Their Implications. In *Proc. of Global Internet*, December 1999.

[43] D. Farinacci, V. Fuller, D. Oran, and D. Meyer. Locator/ID Separation Protocol (LISP). Internet Draft draft-farinacci-lisp-02.txt, July 2007.

[44] D. Farinacci, T. Li, S. Hanks, D. Meyer, and P. Traina. RFC 2784, Generic Routing Encapsulation (GRE), March 2000.

[45] Nick Feamster, Hari Balakrishnan, and Jennifer Rexford. Some Foundational Problems in Interdomain Routing. In *Proc. of Workshop on Hot Topics in Networks (HotNets)*, November 2004.

[46] Nick Feamster, Jay Borkenhagen, and Jennifer Rexford. Guidelines for Interdomain Traffic Engineering. *SIGCOMM Comput. Commun. Rev.*, October 2003.

[47] Zongming Fei, Samrat Bhattacharjee, Ellen W. Zegura, and Mostafa H. Ammar. A Novel Server Selection Technique for Improving the Response Time of a Replicated Service. In *Proc. of IEEE INFOCOM*, March 1998.

[48] Anja Feldmann, Albert Greenberg, Carsten Lund, Nick Reingold, Jennifer Rexford, and Fred True. Deriving Traffic Demands for Operational IP Networks: Methodology and Experience. *IEEE/ACM Trans. Netw.*, 9(3), June 2001.

[49] Anja Feldmann, Olaf Maennel, Z. Morley Mao, Arthur Berger, and Bruce Maggs. Locating Internet Routing Instabilities. In *Proc. of ACM SIGCOMM*, September 2004.

[50] Pierre Francios and Olivier Bonaventure. An Evaluation of IP-based Fast Reroute Techniques. In *Proc. of CoNEXT*, October 2005.

[51] Paul Francis. Comparision Of Geographical And Provider-rooted Internet Addressing. *Computer Networks and ISDN Systems*, 27(3), December 1994.

[52] Paul Francis, Xiaohu Xu, and Hitesh Ballani. FIB Suppression with Virtual Aggregation and Default Routes. Internet Draft draft-francis-idr-intra-va-01.txt, September 2008.

[53] Michael J. Freedman, Karthik Lakshminarayanan, and David Mazires. OASIS: Anycast for Any Service. In *Proc. of USENIX/ACM Symposium on Networked Systems Design and Implementation (NSDI)*, May 2006.

[54] Ayalvadi J. Ganesh, Anne-Marie Kermarrec, and Laurent Massoulie. SCAMP: Peer-to-Peer Lightweight Membership Service for Large-Scale Group Communication. In *Proc. of International COST264 Workshop on Networked Group Communication*, November 2001.

[55] Bob Gillian. VYATTA: Linux IP Routers, Dec 2007. `http://freedomhec.pbwiki.com/f/linux_ip_routers.pdf`.

[56] Albert Greenberg, Gisli Hjalmtysson, David A. Maltz, Andy Meyers, Jennifer Rexford, Geoffrey Xie, Hong Yan, Jibin Zhan, and Hui Zhang. A Clean Slate 4D Approach to Network Control and Management. *ACM SIGCOMM Computer Communications Review*, October 2005.

[57] B. Greene and D. McPherson. ISP Security: Deploying and Using Sinkholes. NANOG meeting, June 2003. `www.nanog.org/mtg-0306/sink.html`.

[58] Saikat Guha and Paul Francis. An End-Middle-End Approach to Connection Establishment. In *Proc. of ACM SIGCOMM*, August 2007.

[59] Krishna P. Gummadi, Stefan Saroiu, and Steven D. Gribble. King: Estimating Latency between Arbitrary Internet End Hosts. In *Proc. of ACM SIGCOMM Internet Measurement Workshop*, November 2002.

[60] Anjali Gupta, Barbara Liskov, and Rodrigo Rodrigues. One Hop Lookups for Peer-to-Peer Overlays. In *Proc. of Workshop on Hot Topics in Operating Systems (HotOS)*, May 2003.

[61] T. Hain. An IPv6 Provider-Independent Global Unicast Address Format. Internet Draft draft-hain-ipv6-PI-addr-02.txt, September 2002.

[62] K. Hamzeh, G. Pall, W. Verthein, J. Taarud, W. Little, and G. Zorn. RFC 2637, Point-to-Point Tunneling Protocol (PPTP), July 1999.

[63] T. Hardy. RFC 3258, Distributing Authoritative Name Servers via Shared Unicast Addresses, April 2002.

[64] R. Hinden and S. Deering. RFC 3513, Internet Protocol Version 6 (IPv6) Addressing Architecture, April 2003.

[65] R. Housley and S. Hollenbeck. RFC 3378, EtherIP: Tunneling Ethernet Frames in IP Datagrams, September 2002.

[66] David Hughes, Dec 2004. PACNOG list posting `http://mailman.apnic.net/mailing-lists/pacnog/archive/2004/12/msg00000.html`.

[67] C. Huitema. RFC 3068, An Anycast Prefix for 6to4 Relay Routers, June 2001.

[68] Geoff Huston. BGP Reports, December 2007. `http://bgp.potaroo.net/`.

[69] Geoff Huston and Grenville Armitage. Projecting Future IPv4 Router Requirements from Trends in Dynamic BGP Behaviour. In *Proc. of ATNAC*, December 2006.

[70] D. Jen, M. Meisel, D. Massey, L. Wang, B. Zhang, and L. Zhang. APT: A Practical Transit Mapping Service. Internet Draft draft-jen-apt-01.txt, November 2007.

[71] David R. Karger, Eric Lehman, Frank Thomson Leighton, Rina Panigrahy, Matthew S. Levine, and Daniel Lewin. Consistent Hashing and Random Trees: Distributed Caching Protocols for Relieving Hot Spots on the World Wide Web. In *Proc. of ACM STOC*, May 1997.

[72] Elliott Karpilovsky and Jennifer Rexford. Using Forgetful Routing to Control BGP Table Size. In *Proc. of CoNext*, December 2006.

[73] Dina Katabi. The Use of IP-Anycast for Building Efficient Multicast Trees. In *Proc. of Global Telecommunications Conference*, December 1999.

[74] Dina Katabi, Mark Handley, and Charlie Rohrs. Congestion Control for High Bandwidth-Delay Product Networks. In *Proc. of ACM SIGCOMM*, August 2002.

[75] Dina Katabi and John Wroclawski. A Framework for Scalable Global IP-Anycast (GIA). In *Proc. of ACM SIGCOMM*, August 2000.

[76] S. Kent. RFC 4303, IP Encapsulating Security Payload (ESP), December 2005.

[77] S. Kent, C. Lynn, and K. Seo. Secure Border Gateway Protocol (S-BGP). *IEEE Journal on Selected Areas in Communication*, 18(4), April 2000.

[78] Angelos D. Keromytis, Vishal Misra, and Dan Rubenstein. SOS: Secure Overlay Services. In *Proc. of ACM SIGCOMM*, August 2002.

[79] Changhoon Kim, Matthew Caesar, Alexandre Gerber, and Jennifer Rexford. Revisiting Route Caching: The World Should Be Flat. In *Proc. of PAM*, April 2009.

[80] Changhoon Kim, Alexandre Gerber, Carsten Lund, Dan Pei, and Shubhabrata Sen. Scalable VPN Routing via Relaying. In *Proc. of ACM SIGMETRICS*, June 2008.

[81] D. Kim, D. Meyer, H. Kilmer, and D. Farinacci. RFC 3446, Anycast Rendevous Point (RP) mechanism using Protocol Independent Multicast (PIM) and Multicast Source Discovery Protocol (MSDP), January 2003.

[82] Dejan Kostic, Adolfo Rodriguez, Jeannie Albrecht, and Amin Vahdat. Bullet: High Bandwidth Data Dissemination Using an Overlay Mesh. In *Proc. of ACM Symposium on Operating Systems Principles*, October 2003.

[83] D. Krioukov, K. Fall, and X. Yang. Compact Routing on Internet-like graphs. In *Proc. of IEEE INFOCOM*, March 2004.

[84] Dmitri Krioukov and kc claffy. Toward Compact Interdomain Routing, August 2005. `http://arxiv.org/abs/cs/0508021`.

[85] Balachander Krishnamurthy, Craig Wills, and Yin Zhang. On the Use and Performance of Content Distribution Networks. In *Proc. of ACM SIGCOMM Workshop on Internet Measurement*, November 2001.

[86] Craig Labovitz, Abha Ahuja, Abhijit Bose, and Farnam Jahanian. Delayed Internet Routing Convergence. *IEEE/ACM Trans. Netw.*, June 2001.

[87] Tony Li. Router Scalability and Moore's Law, October 2006. `http://www.iab.org/about/workshops/routingandaddressing/Router_Scalability.pdf`.

[88] Samantha S. M. Lo and Rocky K. C. Chang. Active Measurement of the AS Path Prepending Method. In *Proc. of ICNP (Poster)*, November 2005.

[89] Z. Morley Mao. Routing Research Issues. In *Proc. of WIRED*, October 2003.

[90] Z. Morley Mao, Randy Bush, Timothy G. Griffin, and Matthew Roughan. BGP Beacons. In *Proc. of ACM SIGCOMM Conference on Internet Measurement*, October 2003.

[91] Zhuoqing Morley Mao, Ramesh Govindan, George Varghese, and Randy H. Katz. Route Flap Damping Exacerbates Internet Routing Convergence. In *Proc. of ACM SIGCOMM*, August 2002.

[92] D. Massey, L. Wang, B. Zhang, and L. Zhang. A Proposal for Scalable Internet Routing & Addressing. Internet Draft draft-wang-ietf-efit-00, February 2007.

[93] S. Matsunaga, S. Ata, H. Kitamura, and M. Murata. Applications of IPv6 Anycasting. draft-ata-ipv6-anycast-app-00, February 2005.

[94] David Meyer, Lixia Zhang, and Ed. Kevin Fall. RFC 4984, Report from the IAB Workshop on Routing and Addressing, September 2007.

[95] Kevin Miller. Deploying IP Anycast. NANOG 29 meeting, October 2003. http://www.net.cmu.edu/pres/anycast/.

[96] Mobility for IPv6 (mip6), IETF Working Group Charter, June 2007. http://www.ietf.org/html.charters/OLD/mip6-charter.html.

[97] T. Narten. Routing and Addressing Problem Statement. Internet Draft draft-narten-radir-problem-statement-02.txt, April 2008.

[98] Netfilter, May 2009. www.netfilter.org.

[99] T. S. Eugene Ng and Hui Zhang. Predicting Internet Network Distance with Coordinates-Based Approaches. In *Proc. of IEEE INFOCOM*, June 2002.

[100] Mike O'Dell. GSE–An Alternate Addressing Architecture for IPv6. Internet Draft draft-ietf-ipngwg-gseaddr-00.txt, February 1997.

[101] Hiroshi Ohta and Marshall Eubanks. MBONE Deployment (mboned) charter, March 2006. http://www.ietf.org/html.charters/mboned-charter.html.

[102] C. Partridge, T. Mendez, and W. Milliken. RFC 1546, Host Anycasting Service, November 1993.

[103] Paul Francis. All About Tunnels, September 2004. www.cs.cornell.edu/Courses/CS619/2004fa/talks/092104-tunnels-francis.pdf.

[104] Bruno Quoitin, Cristel Pelsser, Olivier Bonaventure, and Steve Uhlig. A Performance Evaluation of BGP-based Traffic Engineering. *Intl. Journal of Network Management*, 15(3), May 2005.

[105] Sylvia Ratnasamy, Scott Shenker, and Steven McCanne. Towards an Evolvable Internet Architecture. In *Proc. of ACM SIGCOMM*, August 2005.

[106] R. Ravi and A. Sinha. Multicommodity Facility Location. In *Proc. of ACM-SIAM SODA*, January 2004.

[107] Y. Rekhter, T. Li, and S. Hares (Editor). RFC 4271, A Border Gateway Protocol 4 (BGP-4), January 2006.

[108] Jennifer Rexford, Jia Wang, Zhen Xiao, and Yin Zhang. BGP Routing Stability of Popular Destinations. In *Proc. of ACM SIGCOMM Internet Measurment Workshop*, November 2002.

[109] R. Rodrigues, B. Liskov, and L. Shrira. The Design of a Robust Peer-to-Peer System. In *Proc. of ACM SIGOPS European Workshop*, September 2002.

[110] E. Rosen, A. Viswanathan, and R. Callon. RFC 3031, Multiprotocol Label Switching Architecture, January 2001.

[111] Sandeep Sarat, Vasileios Pappas, and Andreas Terzis. On the Use of Anycast in DNS. Technical report, HiNRG, Johns Hopkins University, December 2004.

[112] Sandeep Sarat, Vasileios Pappas, and Andreas Terzis. On the Use of Anycast in DNS. In *Proc. of ACM SIGMETRICS (Poster)*, June 2005.

[113] John Scudder. Router Scaling Trends. APRICOT Meeting, February 2007. `http://submission.apricot.net/chatter07/slides/future_of_routing`.

[114] Anees Shaikh, Renu Tewari, and Mukesh Agrawal. On the Effectiveness of DNS-based Server Selection. In *Proc. of IEEE INFOCOM*, April 2001.

[115] Alan Shieh, Andrew C. Myers, and Emin Gün Sirer. Trickles: A Stateless Network Stack for Improved Scalability, Resilience and Flexibility. In *Proc. of USENIX/ACM Symposium on Networked Systems Design and Implementation (NSDI)*, May 2005.

[116] W. Simpson. RFC 1853, IP in IP Tunneling, October 1995.

[117] Neil Spring, Ratul Mahajan, and Thomas Anderson. Quantifying the Causes of Path Inflation. In *Proc. of ACM SIGCOMM*, August 2003.

[118] Neil Spring, Ratul Mahajan, and David Wetherall. Measuring ISP topologies with Rocketfuel. In *Proc. of ACM SIGCOMM*, August 2002.

[119] Kunwadee Sripanidkulchai, Aditya Ganjam, Bruce Maggs, and Hui Zhang. The Feasibility of Supporting Large-scale Live Streaming Applications with Dynamic Application Rnd-points. In *Proc. of ACM SIGCOMM*, September 2004.

[120] SSFNet, May 2009. `www.ssfnet.org/homePage.html`.

[121] Ion Stoica, Daniel Adkins, Shelley Zhuang, Scott Shenker, and Sonesh Surana. Internet Indirection Infrastructure. In *Proc. of ACM SIGCOMM*, August 2002.

[122] L. Subramanian, V. Roth, I. Stoica, S. Shenker, and R. Katz. Listen and Whisper: Security mechanisms for BGP. In *Proc. of USENIX/ACM Symposium on Networked Systems Design and Implementation (NSDI)*, March 2004.

[123] Lakshminarayanan Subramanian, Sharad Agarwal, Jennifer Rexford, and Randy H. Katz. Characterizing the Internet Hierarchy from Multiple Vantage Points. In *Proc. of IEEE INFOCOM*, June 2002.

[124] Lakshminarayanan Subramanian, Matthew Caesar, Cheng Tien Ee, Mark Handley, Morley Mao, Scott Shenker, and Ion Stoica. HLP: A Next Generation Inter-domain Routing Protocol. In *Proc. of ACM SIGCOMM*, August 2005.

[125] Lakshminarayanan Subramanian, Ion Stoica, Hari Balakrishnan, and Randy Katz. OverQoS: An Overlay Based Architecture for Enhancing Internet QoS. In *Proc. of USENIX/ACM Symposium on Networked Systems Design and Implementation (NSDI)*, March 2004.

[126] Wei Sun, Z. Morley Mao, and Kang Shin. Differentiated BGP Update Processing for Improved Routing Convergence. In *Proc. of ICNP*, November 2006.

[127] N. Taft, S. Bhattacharyya, J. Jetcheva, and C. Diot. Understanding Traffic Dynamics at a Backbone PoP. In *Proc. of Scalability and Traffic Control and IP Networks SPIE ITCOM*, August 2001.

[128] W. Townsley, A. Valencia, A. Rubens, G. Pall, G. Zorn, and B. Palter. RFC 2661, Layer Two Tunneling Protocol (L2TP), August 1999.

[129] Patrick Verkaik, Andre Broido, kc claffy, Ruomei Gao, Young Hyun, and Ronald van der Pol. Beyond CIDR Aggregation. Technical Report TR-2004-1, CAIDA, 2004.

[130] David X. Wei, Cheng Jin, Steven H. Low, and Sanjay Hegde. FAST TCP: Motivation, architecture, algorithms, performance. *IEEE/ACM Transactions on Networking*, December 2006.

[131] Bernard Wong and Emin Gün Sirer. ClosestNode.com: An Open-Access, Scalable, Shared Geocast Service for Distributed Systems. *SIGOPS Operating Systems Review*, 40(1), January 2006.

[132] Bernard Wong, Aleksandrs Slivkins, and Emin Gün Sirer. Meridian: A Lightweight Network Location Service without Virtual Coordinates. In *Proc. of ACM SIGCOMM*, August 2005.

[133] R. Woodburn and D. Mills. RFC 1241, A Scheme for an Internet Encapsulation Protocol: Version 1, July 1991.

[134] Hong Yan, David A. Maltz, T. S. Eugene Ng, Hemant Gogineni, Hui Zhang, and Zheng Cai. Tesseract: A 4D Network Control Plane. In *Proc. of USENIX/ACM Symposium on Networked Systems Design and Implementation (NSDI)*, April 2007.

[135] Xiaowei Yang. NIRA: A New Internet Routing Architecture. In *Proc. of ACM SIGCOMM Workshop on Future Directions in Network Architecture*, August 2003.

[136] Xiaowei Yang and David Wetherall. Source Selectable Path Diversity Via Routing deflections. In *Proc. of ACM SIGCOMM*, September 2006.

[137] Ellen W. Zegura, Mostafa H. Ammar, Zongming Fei, and Samrat Bhattacharjee. Application-layer Anycasting: A Server Selection Architecture and Use in a Replicated Web Service. *IEEE/ACM Trans. Netw.*, 8(4), August 2000.

[138] Xinyang Zhang, Paul Francis, Jia Wang, and Kaoru Yoshida. Scaling Global IP Routing with the Core Router-Integrated Overlay. In *Proc. of ICNP*, November 2006.

[139] Akamai, May 2006. `http://www.akamai.com/`.

[140] AS112 Project Home Page, May 2006. www.as112.net.

[141] CacheFly, May 2006. http://www.cachefly.com.

[142] CIDR Report, May 2006. http://www.cidr-report.org/.

[143] Global Server Load Balancing, May 2006. http://www.tenereillo.com/GSLBPageOfShame.htm.

[144] History of DNS Root Anycast controversy, May 2006. http://www.av8.net/IETF-watch/DNSRootAnycast/History.html.

[145] ISC F Root-Server, May 2006. http://www.isc.org/index.pl?/ops/f-root/.

[146] National Lambda Rail, April 2006. www.nlr.net.

[147] Report on DNS Amplification attacks, May 2006. http://www.circleid.com/posts/report_on_dns_amplification_attacks/.

[148] Root-Server Technical Operations, May 2006. http://www.root-servers.org/.

[149] Route Views Project Page, May 2006. www.route-views.org.

[150] SprintLink's BGP Policy, May 2006. http://www.sprintlink.net/policy/bgp.html.

[151] clogin Manual Page, October 2008. http://www.shrubbery.net/rancid/man/elogin.1.html.

[152] Foundry Router Reference, July 2008. http://www.foundrynetworks.co.jp/services/documentation/srcli/BGP_cmds.html.

[153] JunOS Route Preferences, July 2008. http://www.juniper.net/techpubs/software/junos/junos60/swconfig60-routing/html/protocols-overview4.html.