IMPROVING CLOUD MIDDLEBOX INFRASTRUCTURE FOR ONLINE SERVICES

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I want to dedicate this thesis to my family for their continuous support, encouragement, patience and love.

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ABSTRACT

Gandhi, Rohan S. PhD, Purdue University, August 2016. Improving Cloud Middlebox Infrastructure for Online Services. Major Professor: Y. Charlie Hu.

Middleboxes are an indispensable part of the datacenter networks that provide high availability, scalability and performance to the online services. Using load balancer as an example, this thesis shows that the prevalent scale-out middlebox designs using commodity servers are plagued with three fundamental problems: (1) The server-based layer-4 middleboxes are costly and inflate round-trip-time as much as 2x by processing the packets in software. (2) The middlebox instances cause traffic detouring en route from sources to destinations, which inflates network bandwidth usage by as much as 3.2x and can cause transient congestion. (3) Additionally, existing cloud providers do not support layer-7 middleboxes as a service, and third-party proxy-based layer-7 middlebox design exhibits poor availability as TCP state stored locally on middlebox instances are lost upon instance failure. This thesis examines the root causes of the above problems and proposes new cloud-scale middlebox design principles that systemically address all three problems.

First, to address the performance problem, we make a key observation that existing commodity switches have resources available to implement key layer-4 middlebox functionalities such as load balancer, and by processing packets in hardware, switches offer low latency and high capacity benefits, at no additional cost as the switch resources are idle. Motivated by this observation, we propose the design principle of using idle switch resources to accelerate middlebox functionailites. To demonstrate the principle, we developed the complete L4 load balancer design that uses commodity switches for low cost and high performance, and carefully fuses a few software load balancer instances to provide for high availability.

Second, to address the high network overhead problem from traffic detouring through middlebox instances, we propose to exploit the principles of locality and flexibility in placing the middlebox instances and servers to handle the traffic closer to the sources and reduce the overall traffic and link utilization in the network.

Third, to provide high availability in a layer 7 middleboxes, we propose a novel middlebox design principle of decoupling the TCP state from middlebox instances and storing it in persistent key-value store so that any middlebox instance can seamlessly take over any TCP connection when middlebox instances fail. We demonstrate the effectiveness of the above cloud-scale middlebox design principles using load balancers as an example. Specifically, we have prototyped the three design principles in three cloud-scale load balancers: Duet, Rubik, and Yoda, respectively. Our evaluation using a datacenter testbed and large scale simulations show that Duet lowers the costs by 12x and latency overhead by 1000x, Rubik further lowers the datacenter network traffic overhead by 3x, and Yoda L7 Load balancer-as-a-service is practical; decoupling TCP state from load balancer instances has a negligible (<1%) overhead.

1. CLOUD MIDDLEBOX INFRASTRUCTURE

Cloud computing has become an integral part of our lives. The online services that provide important services such as email, news, banking, maps, search and music are now residing in the cloud. The success of such cloud online services is incomplete without middleboxes on which online services heavily rely for several network operations including distributing load (load balancer, proxy), access control (firewall), remote access (VPN). For example, without load balancers, the online services cannot easily scale out to thousands of servers and support millions of user requests per second. Additionally, without firewalls the online services are easily exposed to several attacks and malicious activities that threaten to steal user confidential information. Together, middleboxes simplify deployment, monitoring, security and management of the online services making them an indispensable component of the online services. Given these important benefits, the middlebox market is expected to cross $10B by 2016 [1], and become comparable to the router market.

1.1 Requirements

All user traffic coming to the online services first goes through the middleboxes that perform the network operations before sending the traffic to the servers assigned to the online services (Figure 1.1). For example, load balancer (LB) in Figure 1.1 splits and forwards the requests among hundreds of web front ends, or the firewall inspects the incoming traffic for malicious activities.

In addition to handling the traffic coming from Internet, the middleboxes also handle a large fraction of the total traffic within the datacenter (DC) because many services inside the DC also take advantage of several benefits provided by the middleboxes. For example, middleboxes such as a load balancer provide a *virtual IP (VIP)*

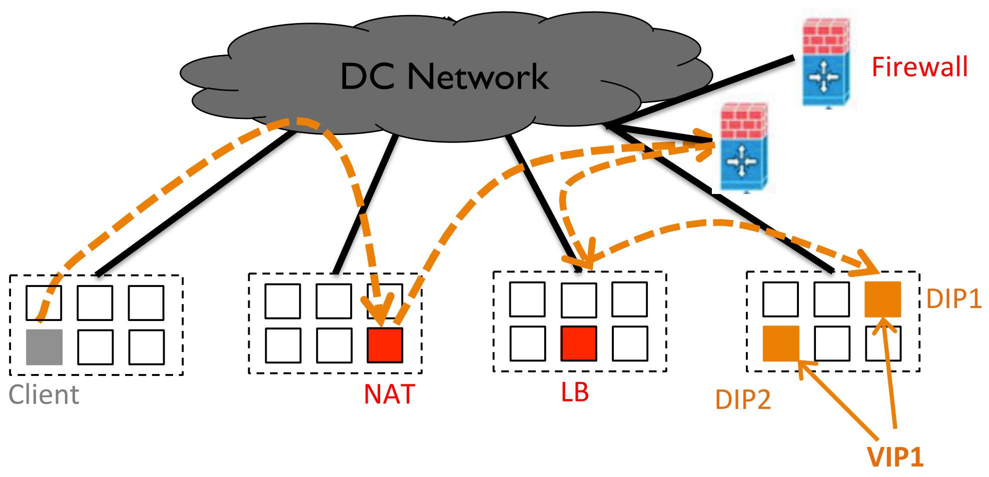


Fig. 1.1.: Typical middlebox deployment. The red boxes denote the software middleboxes, whereas the blue symbols denote the specialized hardware middleboxes. The arrows show the traffic flow when the client is within the same DC.

abstraction, where an online service receives traffic on one or a small number of VIPs, and the load balancer splits the VIP traffic across servers assigned to the online services. This indirection through the VIP provides many advantages, foremost is the isolation, where it masks the dynamics such as server failure, migration and maintenance within the online services from the users and other dependent online services. To take advantage of such benefits, all the traffic between online services within the same datacenter goes through the load balancer. It is reported that 44% of the total datacenter traffic is VIP traffic.

Supporting such a high traffic volume adds significant strain on the middlebox data-plane. Additionally, as the user traffic directly flows through the middleboxes, middleboxes face stringent requirements on the availability and performance. Precisely, the middleboxes face the following requirements:

- **High scalability:** The middleboxes are expected to provide high capacity and scalability to support large and highly dynamic nature of the user traffic. For example, a load balancer is expected to handle traffic as high as 44 Tbps in a mid-size datacenter with 40K servers [2], and this number is expected to grow rapidly as new online services are being added and as the traffic to existing online services grows. The middleboxes are also expected to provide high capacity to individual tenants, *e.g.,* handle 100+ Gbps traffic or 1+ million simultaneous connections. Additionally, the online traffic is highly volatile, where the max. to min. traffic ratio in a single day on average is 6x [3]. In such settings, the middleboxes should be able to dynamically scale up/down to adapt to the dynamic traffic volume.

- **High availability:** The middleboxes are required to provide high availability as any downtime in the middleboxes can result in user traffic being dropped or loss of connectivity of the online services, either of which directly reduces online service availability and affects their revenue.
- **Low latency:** The online services have strict latency deadlines in responding to the user requests, which also translates into middleboxes minimizing latency inflation in processing the packets as any time lost in the middlebox processing can otherwise be used by the online services to improve the response quality.
- **Low cost:** The middleboxes are expected to be low cost, as the money spent on the middleboxes can otherwise be used to improve the online services infrastructure. Prior work estimates that the cost of the middleboxes should not exceed 1% of the total server costs [2].
- **Low network bandwidth overhead:** Typically, the middleboxes run separately from the servers and the clients, and the traffic from clients first need to detour to the middleboxes en route to the destinations. The network operators want to reduce such traffic detour as: (1) The middleboxes handle a huge traffic volume, such detour can unnecessarily inflate the total traffic in the network, which may require network operators to increase network capacity which is costly; (2) The high traffic detour may also trigger transient congestion, which can damage the performance of the latency-sensitive applications.

- **Layer-2 to layer-7 support:** The middleboxes are required to do network operations based on any of the layer-2 to layer 7 (OSI model) fields. For example, a network firewall should be able to block traffic from certain MAC/IP addresses or to certain ports. Similarly, layer-7 load balancer should be able to split the traffic based on the HTTP header.

1.2 Specialized hardware middleboxes and limitations

Traditionally, specialized hardware middleboxes are used to provide middlebox functionalities. Specialized hardware middleboxes are optimized for performance for individual middlebox operations using specially designed hardware. However, these middleboxes do not meet the requirements in cloud. First, these middleboxes are costly at scale. It is reported that the hardware middleboxes can cost $80K for 20Gbps traffic, far beyond the middlebox budget. Second, they have poor availability, where they provide only 1+1 redundancy which is not enough at cloud scale. Third, these middleboxes lack flexibility, as they use special hardware optimized only for certain middlebox functions, which cannot be extended easily.

1.3 Software middleboxes and limitations

Software middleboxes address the limitations of the specialized hardware middleboxes. Software middleboxes implement middlebox functions in the software running on commodity servers. For example, Ananta [2] software load balancer consists of a central controller, and several software Muxes (SMux) that provide a distributed data plane. Each SMux maintains mappings between the VIPs and servers for all online services, and implements traffic splitting and encapsulation functionality in software.

The software middleboxes scale out to hundreds of servers to support large traffic volume. The advantage of the software middleboxes is that they can easily scale-out, where the servers can be dynamically added or removed to match the traffic demand. Additionally, software middleboxes provide high flexibility due to software implementation, and provides n+1 redundancy as there are n instances active when one instance fails.

Despite these benefits, software middleboxes suffer from numerous limitations, which we enumerate below:

1.3.1 High Cost

First, software middleboxes are costly due to poor capacity at individual servers. This limitation stems from processing the packets in the software. In case of a load balancer, we observed that the individual servers can only handle 300K packets/sec as the CPU saturates at this rate. This limited capacity translates into requiring 4K servers to handle 15Tbps of the traffic only for load balancing, which is typical in a midsize 40K datacenter [2, 4], which far exceeds the middlebox budget.

1.3.2 High latency inflation

Second, the software middleboxes incurs high and highly variable latency again by processing the packets in software. Again, in case of a load balancer, we observed latency inflation anywhere between 200-1000 μsec when processing packets at rate as low as 100K packets per sec, which is significant as the typical datacenter RTTs are around 330μsec. This results in high end-to-end latency and low throughput as every packet coming to the online service suffers from this high latency. Applications such as algorithmic stock trading and high performance distributed memory caches demand ultra-low (a few microseconds) latency within the data center. For such applications, the latency inflation by the software middleboxes is not acceptable.

1.3.3 High network bandwidth overhead

In addition to the limitations in cost and latency, software middleboxes also inflate the network bandwidth usage and maximum link utilization (MLU) indicating congestion due to large traffic volume. The middleboxes are placed oblivious to the location of the sources and destinations. Thus, even if the sources and destination are in the same rack, the traffic first has to traverse many hops to reach the middleboxes. Second, the middleboxes typically choose the destination end-point independent of its location (using hash calculated over IP 5-tuples). Thus, even if the destinations are in the same rack as the middleboxes, the middleboxes send the traffic to destinations many hops away. Lastly, existing designs do not take advantage of the flexibility in placing the end-points. End-points can be placed closer to the middleboxes and sources to shorten the path.

Our analysis using real traffic traces showed that indeed all the traffic hits the datacenter network core even if most of it can be contained in individual racks. This detour to/from the middleboxes increase the number of hops and inflates traffic on the individual links and max. link utilization beyond 90% indicating congestion. Such high bandwidth usage not only requires the datacenter (DC) operator to provision high network bandwidth which is costly, but also makes the network more prone to transient congestion which affect latency-sensitive services.

1.3.4 Poor availability

The previous discussion was focused on a general class of middleboxes called layer-4 middleboxes that process packets based on the up to layer-4 fields such as TCP/IP fields. A special class of middleboxes called layer-7 middleboxes that process packets based on the up to layer 7 fields such as HTTP or FTP suffer from an additional problem of poor availability, where the client and server connections break when individual instances fail. The root-causes for this undesired problem are deeply embedded into a proxy design heavily used by the layer-7 middleboxes especially load balancer, where for an incoming client request, each load balancer instance first establishes a TCP connection with the client to get the HTTP request, then selects the server and starts a new connection with the server. However, the proxy instance stores the client and server TCP connection state locally only, and is lost on failure. Even if the packets get rerouted to another instance, it cannot recover the TCP state required to maintain the connection. As a result, the client and server connections break, which results in poor user experience and affects the revenue of online services.

1.3.5 Costly and limited cloud offerings

Major cloud providers do not provide layer-7 middleboxes as a service, which prompts the tenants to either build and maintain their own service or use third party designs (such as HAProxy). In either alternatives, the tenants have to manage scalability of their middleboxes on their own. However, we observe that the tenants cannot scale down the middlebox instances easily as removing instances could break the existing connections as explained previously. Thus, the middlebox would remain provisioned for the peak traffic even when the average traffic is small, which unnecessarily incurs high operating costs.

In summary, existing specialized hardware and software based middleboxes fail to effectively meet all the requirements at the cloud scale, which calls for a new class of middlebox designs.

1.4 Thesis Overview

Using load balancer as an example, this thesis presents a new class of middleboxes that address all the limitations of their predecessors. As a first step, we design a new layer-4 load balancer that offers high capacity at ultra low latency while substantially reducing the load balancer cost. These improvements stem from a unique observation that the existing commodity switches in the datacenters have idle resources to implement load balancer and other major layer-4 middleboxes functionalities. We further improve the network bandwidth overhead using the principles of: (1) locality, *i.e.,* placing the load balancer instances closer to the sources and destinations to minimize the traffic detour, (2) end-point flexibility, *i.e.,* flexibility to place the end-point servers closer to the sources to further amplify the locality.

We show that the design choices that provided high availability in the layer-4 load balancers cannot be extended to the layer-7 load balancers. Thus, we build a new highly available layer-7 load balancer using a principle of decoupling the TCP state from middlebox instances and storing it persistently.

Below we briefly overview the journey in building such new class of the load balancers.

1.4.1 Load balancers using hardware and software

This thesis first proposes a new class of load balancer design, Duet, as a first step to address the poor performance and high cost limitations of the existing layer-4 load balancers. Duet is based on a key observation that existing commodity switches which are in abundance in today's datacenters have resources that can be re-purposed to implement key layer-4 middlebox functionalities including load balancer and firewall. Specifically, we show that using these resources we can build Hardware Multiplexer (HMux) on the switch that is functionally the same as the Software Multiplexer (SMux) that runs on commodity servers in software load balancing [2]. Moreover, these resources are idle, *i.e.,* not used by the existing applications in the datacenter. Thus, using these resources incurs no cost to the cloud provider. Moreover, by processing the packets in the hardware, switches offer low latency and high capacity benefits, where even a single switch can handle traffic in the order of 100 Gbps with latency inflation in the order of 10μsec.

While switches offer high capacity, low latency and low cost, the architecture is less flexible than software load balancers. Specifically, handling certain cases of switch failures is challenging (§2.4.1). Thus, our second idea is to integrate the switch-based load balancer with a small deployment of software load balancer instances, to achieve the best of both worlds. We make the integration seamless using simple routing mechanisms. In the combined design, most of the traffic is handled by the switch-based hardware load balancer, while software load balancer acts as a backstop, to ensure high availability and provide flexibility.

Compared to dedicated hardware load balancers, or pure software load balancers (Ananta), DUET is highly cost effective. It load-balances most of the traffic using *existing* switches, and needs only a small deployment of software load balancer as a backstop. Because most of the traffic is handled by the HMuxes, DUET has significantly lower latency than software load balancers. At the same time, use of software load balancer enables DUET to inherit high availability and flexibility of the software load balancer.

To design DUET, we addressed two main challenges. First, individual switches in the data center do not have enough memory to hold the load balancer mapping database for all online services. Thus, we need to partition the mappings among the switches. We devise a simple greedy algorithm to partition the mappings that attempts to minimize the "leftover" traffic (which is perforce handled by the software load balancer), while taking into account constraints on switch memory and demands of various traffic flows.

The second challenge is that this mapping must be regularly updated to adapt to the variety of datacenter conditions. For example, VIPs or DIPs are added or removed by customers, switches and links fail and recover. We devise a migration scheme that avoids memory deadlocks and minimizes unnecessary VIP movement.

Together, using these techniques DUET provides an organically scalable, high performance and highly available layer-4 load balancer.

1.4.2 Reducing network overhead by exploiting locality and end-point flexibility

Despite the benefits in Duet, it continues to suffer from high network bandwidth overhead problem (§1.3.3). In fact, it suffers from an additional problem that the traffic detouring through core links breaks the full-bisection bandwidth guarantees originally provided by full-provisioned networks such as Clos and FatTree [31].

We propose Rubik, a new LB that significantly reduces the high bandwidth usage by LB. Like Duet, Rubik uses a hybrid LB design consisting of the HMuxes and SMuxes, and aims to maximize the VIP traffic handled by HMuxes to reduce the LB costs. While doing that, Rubik reduces the bandwidth usage using two synergistic design principles. First, Rubik exploits the locality, *i.e.,* it tries to load balance VIP traffic generated within individual ToRs across the DIPs residing in the same ToRs. This reduces the total traffic entering the core network. Second, Rubik exploits end-point flexibility, *i.e.,* it tries to place the DIPs for a VIP in the same ToRs as the sources generating the VIP traffic.

To exploit locality, Rubik uses a novel architecture that splits the VIP-to-DIP mapping for a VIP into multiple “local” and a single “residual” mappings stored in different HMuxes. The local mapping stored at a ToR handles the traffic generated in the ToR across the DIPs in the same ToR. The residual mapping assigned to an HMux handles the traffic not handled by local mappings and maximizes the total VIP traffic handled by HMuxes.

To exploit locality and end-point flexibility, Rubik faces numerous challenges. First, there are limited resources – individual switches have limited memory (where VIP-to-DIP mappings are stored) and individual ToRs have limited servers (where DIPs can be assigned). Also, individual DIPs (servers) have limited capacities. Exploiting end-point flexibility further compounds the challenge as there are dependencies across services. The dependencies arise because many large services are multi-tiered; when a subservice at tier i receives a request, it spawns multiple requests to the subservices at tier $(i+1)$. Because of such dependencies, traffic sources at a lower tier are not known until DIPs in the higher tier are placed. Furthermore, Rubik needs to ensure that it assigns DIPs that satisfy SLAs.

We develop a practical two-step solution to address the above challenges. In the first step, we design an algorithm to jointly calculate the DIP placement and mappings to maximize the traffic contained in ToRs while satisfying various constraints using an LP solver. In the second step, we use a heuristic assignment to maximize the total traffic handled by HMuxes to reduce the costs.

Lastly, to adapt to the cloud dynamics such as changes in the VIP traffic, failures, *etc.*, RUBIK regularly updates its local, residual mappings and DIP placement while limiting the number of servers migrated.

Using all these techniques, RUBIK substantially reduces the network bandwidth overhead while maintaining all the benefits of DUET.

1.4.3 High Availability through decoupling TCP State

Although DUET and RUBIK provide highly available layer-4 load balancer, the design choices that enabled the high availability cannot be extended to provide high availability in the layer-7 load balancer. This is because layer-4 software middleboxes naturally offer high availability as all instances select the same back-end server using the hash calculated on the IP 5-tuples.

However, layer-7 middleboxes do not have such flexibility because the HTTP header is carried only in the first few packets of a given connection. When one middlebox instance fails after receiving the HTTP header, and packets get rerouted to another instance, the later instance cannot determine the the server assigned to that connection because it does not have the HTTP header required to select the back-end server.

We propose YODA, which enables L7 load balancing as a service to the tenants in public clouds. YODA is based on a scale-out design and uses existing VMs in the cloud to address the above drawbacks of current per-tenant, proxy-based L7 LB solutions. YODA addresses the availability challenge using two ideas. We observe that in current L7 LB solutions, the key reason a client flow handled by a failed L7 LB instance could not migrate to another instance is that the failed instance had end-to-end connections with the server/client, and hence the TCP states for the two connections stored locally at the L7 LB instance would be lost upon the instance failure. The first key idea of YODA is to have the L7 LB instances use the VIP in connections with the server and the client so that neither connection is tied to the instance's IP. This is a prerequisite for migrating the client flow to another LB instance in a way that is transparent to the client and the server. Our insight is that LB instances can easily use the VIP by leveraging existing L4 LB service in the cloud.

The above mechanism allows the LB instance to receive and forward packets between the client and the server using the VIP, but the *flow state*, consisting of the two TCP states and the selected server, are still stored locally on the LB instance and hence can be lost upon the instance failure. Our second idea is to decouple such flow state from the specific LB instance handling the connection and store it in a high performance persistent datastore, called TCPStore, that we built on top of Memcached [5]. As a result, when an LB instance fails, the flow state of the traffic it was handling can be retrieved from TCPStore by any of the remaining LB instances. Together with the first idea of using VIP, the new LB instance can seamlessly take over the flows. Using VIP and decoupling and sharing the flow state among LB instances this way also simplify providing LB scalability, since the flows can easily migrate to go through other LB instances as LB instances are added or removed.

The second goal in YODA is on how to provide YODA as a shared cloud service to reduce the operational costs as well as absorb the management overhead. YODA provides an effective rule-based interface to the online service operators to specify the policies on how to split the traffic. This rule-based interface can result in millions of rules for individual online services as there are numerous combinations possible on layer-3 to layer-7 fields, which triggers a new challenge of how to assign the rules to the YODA instances.

Assigning rules for all tenants to all LB instances provides high robustness, where upon one LB instance failure, any of the remaining LB instances can handle its traffic. However, this choice also increases the number of rules on individual LB instances which significantly inflates latency. To address this challenge, we develop a many-to-

Table 1.1.: Thesis contributions, new DC middlebox princoples and outcomes

Existing design limitations	New design principles	Outcomes
High cost	Use of hardware	12x reduction
High latency		10x reduction
High network bandwidth	Using locality and end-point flexibility	3x reduction in bandwidth 4x reduction in MLU
Poor availability	Replicating TCP state	Close to highest availability at low overhead in performance
Costly cloud offerings	Balancing costs and redundancy	3.7x reduction in costs 4x more redundancy

many VIP assignment algorithm that minimizes the LB instance cost while giving a level of guarantee on both latency and failure resilience.

1.5 Thesis contributions

In summary, this thesis makes the following three contributions, and establishes many "firsts" in advancing the middlebox infrastructure in the cloud (summarized in Table 1.1).

First, we characterize the conditions, design challenges, and design principles for moving layer-4 load balancing functionality directly into hardware switches which offer significantly lower latency and higher capacity than software servers. We present the design and implementation of a switch-based load balancer. To the best of our knowledge, this is the first such design. We show how to seamlessly combine the switch-based load balancer with a small scale software load balancer to achieve high availability and flexibility. Again, to the best of our knowledge, this is the first "hybrid" load balancer design.

Second, through careful analysis of the LB workload from one of our production DCs, we show that the existing LB incur high network bandwidth overhead. We present the design and implementation of Rubik that overcomes these inefficiencies by exploiting traffic locality and end-point flexibility. To the best of our knowledge, this is the first LB design that exploits these principles.

Third, we present the design and implementation of a highly available and scalable L7 load balancer as-a-service for public clouds. We present two key ideas for achieving high availability of a L7 LB: the design principle of decoupling the flow state from the LB instances and storing it in a persistent storage, and leveraging the L4 LB service to enable each L7 LB instance to use the VIP in interacting with both the client and the server. We present an effective algorithm for calculating VIP-LB instance assignment to minimize the cost while guaranteeing a given level of LB robustness and performance.

We believe the design choices and principles proposed in the context of the load balancer can be easily applied to the broader categories of layer-4 and layer-7 middleboxes.

Outcomes: We evaluate Duet using a testbed implementation as well as extensive, large-scale simulations. Our results show that Duet provides 12x more capacity than the pure software load balancer, at a fraction of the software load balancer cost, while also reducing the latency inflation by 10x or more. Additionally, we show that Duet quickly adapts to the network dynamics in the data center including failures. Through testbed experiments and extensive simulations, we show that Rubik reduces the DC network bandwidth usage by 3x and the MLU (max. link utilization) by over 4x while providing a high performance and highly available LB.

We evaluate Yoda using a prototype deployed on a 60-VM testbed in Windows Azure, and large scale simulations. Our results show that the flow state can be captured, and used on different VMs to maintain the flows, transparently to the servers and clients, and the overhead of decoupling TCP state is very small (<1 msec). Our simulation results using a one-day trace from production services show that YODA-as-a-service reduces L7 LB instance cost for the tenants by 3.7x while providing 4x more redundancy.

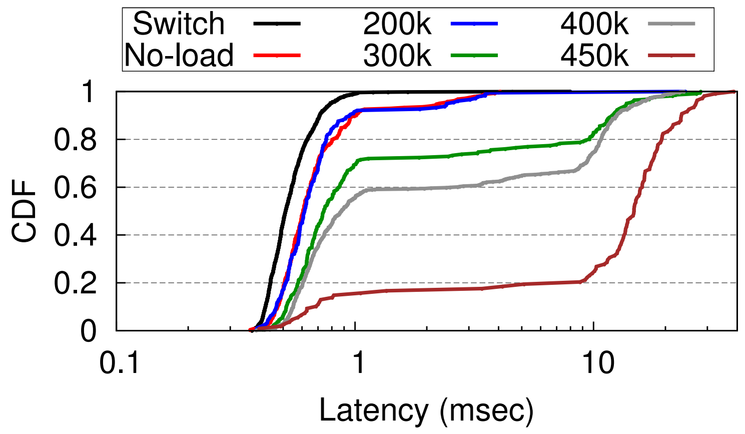
2. DUET: CLOUD SCALE LOAD BALANCING WITH HARDWARE AND SOFTWARE

2.1 Background and Motivation

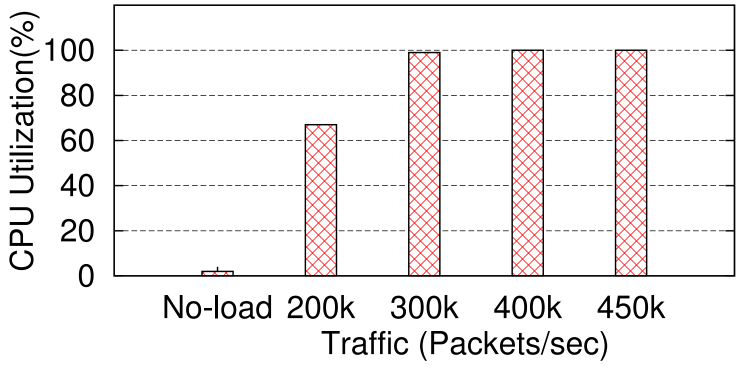
In this chapter, we present a brief background on load balancing functionality in datacenters (DCs), briefly describe a software-only load balancer architecture (Ananta), and point out its shortcomings.

A DC typically hosts multiple services. Each service is a set of servers that work together as a single entity. Each server in the set has a unique direct IP (DIP) address. Each service exposes one or more virtual IP (VIP) outside the service boundary. The load balancer forwards the traffic destined to a VIP to one of DIPs for that VIP. Even services within the same DC use VIPs to communicate with each other, since the indirection provided by VIPs offers several benefits. For example, individual servers can be maintained or upgraded without affecting dependent services. Management of firewall rules and ACLs is simplified by expressing them only in terms of VIPs, instead of DIPs, which are far more numerous and are subject to churn.

The key to the efficient functioning of the indirection architecture is the load balancer. A typical DC supports thousands of services [2, 6], each of which has at least one VIP and many DIPs associated with it. All incoming Internet traffic to these services and most inter-service traffic go through the load balancer. As in [2], we observe that almost 70% of the total VIP traffic is generated within DC, and the rest is from the Internet. The load balancer design must not only scale to handle this workload but also minimize the processing latency. This is because to fulfill a single user request, multiple back-end services often need to communicate with each other — traversing the load balancer multiple times. Any extra delay imposed by the load balancer could have a negative impact on end-to-end user experience. Besides that,



(a) End-to-end latency



(b) CPU Utilization

Fig. 2.1.: Performance of software Mux.

the load balancer design must also ensure high service availability in face of failures of VIPs, DIPs or network devices.

2.1.1 Ananta Software Load Balancer

We first briefly describe the Ananta [2] software load balancer. Ananta uses a three-tier architecture, consisting of ECMP on the routers, several software Muxes (SMuxes) that run on commodity servers, and are deployed throughout the DC, and a host agent (HA) that runs on each server.

Each SMux stores the VIP to DIP mappings for all the VIPs configured in the DC. Using BGP, every SMux announces itself to be the next hop for every VIP. Incoming packets for a VIP are directed to one of the SMuxes using ECMP. The SMux selects a DIP for the VIP, and encapsulates the packet, setting the destination address of the outer IP header to the chosen DIP. At the DIP, the HA decapsulates the incoming packet, rewrites the destination address and port, and sends it to server. The HA also intercepts outgoing packets, and rewrites their IP source addresses from the DIP to the VIP, and forwards the direct server return (DSR).

Ananta can support essentially an unlimited number of VIPs and DIPs, because it stores this mapping in the large main memory on commodity servers. While a single SMux in Ananta has limited capacity (due to software processing), Ananta can still scale to handle large volumes of traffic. First, Ananta deploys numerous SMuxs, and relies on ECMP to split the incoming traffic among them. Second, DSR ensures that only the incoming or the VIP traffic goes through the load balancer. Ananta also includes a mechanism called fast path to enhance scalability. Fast path allows all inter-service traffic to directly use DIPs, instead of using VIPs. However, this negates the benefits of the VIP indirection. For example, if fast path is enabled, service ACLs have to be expressed in terms of DIPs.

In summary, implementing parts of load balancing functionality in software allows Ananta to be highly scalable and flexible. However, processing packets in software is also the Achilles heel for Ananta, because it adds latency, and limits the throughput, as we discuss next.

2.1.2 Limitations of Software Load Balancer

Figure 2.1(a) shows the CDF of the RTTs for the VIP traffic load-balanced by a production Ananta SMux as traffic to the VIP varies between 0 and 450K packets/sec. Even at zero load the SMux adds a median latency of 196μsec. The latency variance is also significant, with the 90^{th} percentile being 1ms. The median RTT (without load balancer) in our production DCs is 381μsec, so the inflation in latency is significant for the intra-DC traffic, which accounts for 70% of the total VIP traffic. (For the remaining traffic from the Internet, it is a lesser problem due to larger WAN latencies). The high latency inflation and high latency variability result from processing the packets in software. We also see that the added latency and the variance get much worse at higher load.

The results also illustrate that an individual SMux instance has low capacity. Beyond 300K packets/sec, the CPU utilization reaches 100% (Figure 2.1(b)). Thus, for the hardware SKU used in our DCs, each SMux can handle only up to 300K packets/sec, which translates to 3.6 Gbps for 1,500-byte packets. At this rate, supporting 15 Tbps VIP traffic for a mid-sized (40K servers) DC would require over 4K SMuxes, or 10% of the DC size; which is unacceptable[1].

2.2 Duet: Core ideas

In the previous section, we saw that while software load balancers are flexible and scalable, they suffer from low throughput and high latency. In this thesis, we propose a new design called Duet that offers scalability, high throughput and low latency, at a small fraction of the software load balancer's cost.

Duet is based on two novel ideas. First, we leverage idle resources of modern, commodity data center switches to construct a hardware load balancer. We call this design Hardware Mux (HMux). HMux offers microsecond latency, and high capacity, without the need for *any* additional hardware. However, the HMux design suffers from certain shortcomings. Thus, our second idea is to combine the HMux with Ananta-like software Mux (SMux). The combined system is called Duet in which the SMux acts as a backstop for the HMux.

We now describe the design of HMux. To simplify the description, we will assume that the DC is not virtualized, *i.e.,* one DIP corresponds to one server. The changes required to support VMs are described in §2.4.2.

[1]Newer technologies such as direct-packet IO and RDMA may help match packet processing capacity of the SMux to that of the NIC (10 Gbps), but they may not match packet processing capacity of the switch (600 Gbps+) as we explain in § 2.2.1.

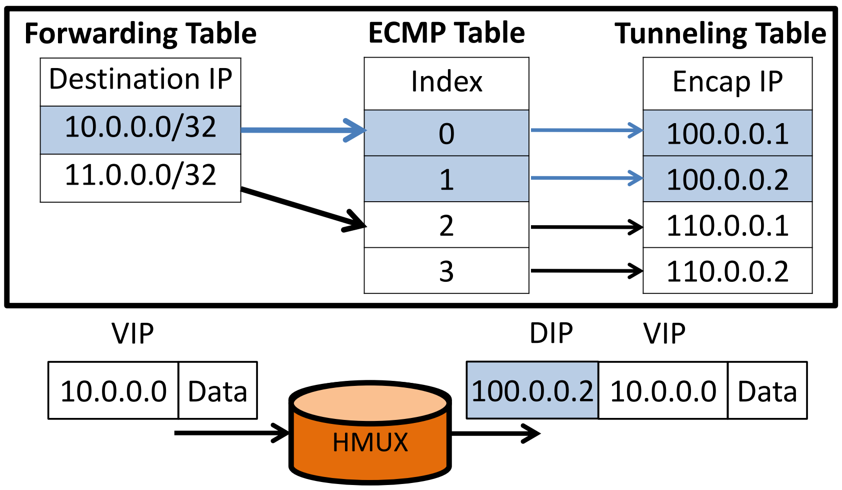


Fig. 2.2.: Storing VIP-DIP mapping on a switch.

2.2.1 HMux

Ananta's SMux implements two key functions to load balance traffic: (1) for each VIP, split traffic equally among its DIPs, and (2) use IP-in-IP encapsulation to route the VIP traffic to the corresponding DIPs. Both of these functions have long been available on commodity switches, *i.e.*, traffic splitting is supported using ECMP and IP-in-IP encapsulation is supported using tunneling. However, major switch vendors have only recently started to provide the APIs for fine-grained control over ECMP and tunneling functionality.

Our key insight is that by carefully programming the ECMP and tunneling tables using these new APIs, we can make a commodity switch act as a hardware Mux (HMux), in addition to its normal functionality. In fact, this can be easily done on most of the switches used in our DCs today.

Figure 2.2 shows the HMux design. A packet arriving at a switch goes through a processing pipeline. We focus on three tables used in the pipeline. The packet matches one entry in the host forwarding table which then points to multiple ECMP table entries. These ECMP table entries correspond to multiple next hops for the packet[2]. The actual next hop for the packet is selected by using the hash of the IP 5-tuple to index into the ECMP table. The tunneling table enables IP-in-IP encapsulation by storing the information needed to prepare the outer IP header for a given packet.

To construct HMux, we link the ECMP and tunneling functionalities. Consider a packet destined for VIP 10.0.0.0 that arrives at the HMux. There are two DIPs (100.0.0.1 and 100.0.0.2) for this VIP. The host forwarding table indicates that the first two entries in the ECMP table pertain to this VIP. The ECMP entries indicate that packets should be encapsulated, and point to appropriate entries in the tunneling table. The switch encapsulates the packet using IP-in-IP encapsulation, and the destination address in the outer IP header is set to the DIP address specified in the tunneling table entry. The packet is then forwarded to the appropriate interface.

Thus, at the expense of some entries in the host forwarding, ECMP and tunneling tables, we can build a load balancer using commodity switches. In fact, if all the VIP-to-DIP mappings are stored on every top-of-rack (ToR) switch as well as every access switch, this HMux design can provide load balancing functionality to all intra-DC and inter-DC traffic. However, the amount of space available in the three tables is limited, raising two distinct issues.

Number of VIPs: The first problem is the size of the host forwarding table. The switches in our DC have 16K entries in the host table. The host table is mostly empty, because it is used only for routing within a rack. But even the 16K entries may not be enough to hold all VIPs in a large DC. One way to address this problem is by using longest prefix match (LPM) forwarding table. However, LPM table is heavily used for routing within and across DCs, and is not available to be used for load balancing. We support higher number of VIPs using SMuxes as explained in §2.2.3.

[2]The information is split between ECMP group table and ECMP table; we omit such details due to lack of space.

Number of DIPs: The second problem concerns the sizes of the ECMP and tunneling tables. ECMP table typically holds 4K entries, and is mostly empty (see § 2.8). The tunneling table typically holds 512 entries. In our DC, few applications use tunneling, so these entries are mostly free as well. The number of DIPs an individual HMux can support is the minimum of the number of free entries in the ECMP and the tunneling tables (see Figure 2.2). Thus, an individual HMux can support at most 512 DIPs. This is orders of magnitude smaller than the total number of DIPs. We address this challenge next.

2.2.2 Partitioning

We address the problem of limited size of ECMP and tunneling tables using two mechanisms: (1) We divide the VIP-to-DIP mapping across multiple switches. Every switch stores only a small subset of all the VIPs, but stores all the DIPs for those VIPs. This way of partitioning ensures all the traffic for a particular VIP arrives at a single switch and the traffic is then equally split among the DIPs for that VIP. (2) Using BGP, we announce the VIPs that are assigned to the switches, so that other switches can route the VIP packets to the switch where the VIP is assigned.

Figure 2.3 illustrates this approach. VIP_1 has two DIPs (D_1 and D_2), whereas VIP_2 has one (D_3). We assign VIP_1 and VIP_2 to switches C_2 and A_6 respectively, and flood the routing information in the network. Thus, when a source S_1 sends a packet to VIP_1, it is routed to switch C_2, which then encapsulates the packet with either D_1 or D_2, and forwards the packet.

Another key benefit of partitioning is that it achieves *organic scalability* of HMuxes — when more servers are added in the DC and hence traffic demand increases, more switches will also be added and hence the aggregate capacity of HMuxes will also increase proportionally.

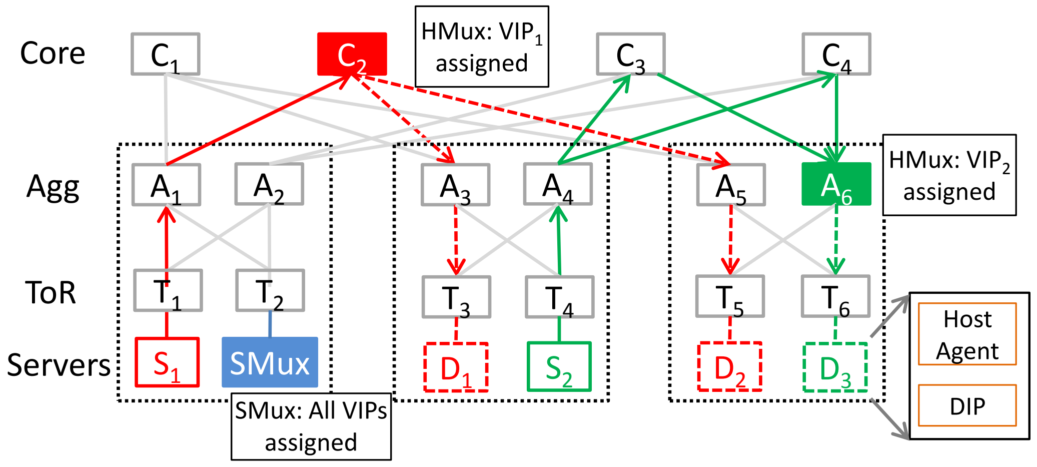


Fig. 2.3.: **Duet architecture**: VIPs are partitioned across different HMuxes — VIP_1 and VIP_2 are assigned to HMux C_2 and A_6. Additionally, SMuxes act as backstop for all the VIPs. Every server (apart from SMuxes) runs host-agent that decapsulates the packets and forwards to the DIP. Links marked with solid lines carry VIP traffic, and links with dotted lines carry DIP traffic.

2.2.3 DUET: HMux + SMux

While partitioning helps increase the number of DIPs HMux can support, that number still remains limited. The HMux design also lacks the flexibility of SMux, because VIPs are partitioned and "pinned" to specific HMuxes. This makes it challenging to achieve high VIP availability during network failures. Although replicating VIP across a few switches may help improve failure resilience, it is still hard to achieve the high availability of Ananta because Ananta stores the complete VIP-DIP mappings on a large number of SMuxes.

This motivates us to architect Duet— a new load balancer design to fuse the flexibility of SMux and the high capacity and low latency of HMux.

Design

Duet's goal is to maximize VIP traffic handled using HMux, while using SMux as a backstop. Thus, besides an HMux on each switch, Duet also deploys a small number of SMuxes on commodity servers (figure 2.3). The VIPs are partitioned among HMuxes as described earlier. In addition, each SMux announces all the VIPs. The routing protocol preferentially routes VIP traffic to HMux, ensuring that VIP traffic is primarily handled by HMux – thereby providing high capacity and low latency. In case of HMux failure, traffic is automatically diverted to SMux, thereby achieving high availability. To ensure that existing connections do not break as a VIP migrates from HMux to SMux or between HMuxes, all HMuxes and SMuxes use the same hash function to select DIPs for a given VIP.

The preferential routing to HMux can be achieved in several ways. In our current implementation, SMux announces the VIPs in aggregate prefixes, while HMux announces /32 routes to individual VIPs. Longest prefix matching (LPM) prefers /32 routes over aggregate prefix routes, and thus directs incoming VIP traffic to appropriate HMux, unless that HMux is unavailable.

The number of SMuxes needed depends on several factors including the VIP traffic that cannot be assigned to HMux due to switch memory or link bandwidth limits (§2.3), the VIP traffic that failovers to SMux due to HMux failure (§2.4.1), and the VIP traffic that is temporarily assigned to SMux during VIP migration (§2.3.2). We estimate it based on historical traffic and failure data in DC.

Benefits

The key benefits of Duet are summarized below.

Low cost: Duet does not require any additional hardware – it uses idle resources on existing switches to provide load balancing functionality. Duet also requires far fewer SMuxes than Ananta, since SMuxes are used only as a backstop for HMuxes, and hence carry far less traffic.

High capacity and low latency: this is because VIP traffic is primarily handled by HMux on switch.

High availability: by using SMux as a backstop during failures, Duet enjoys the same high availability as Ananta.

High limit on number of VIPs: If the number of VIPs exceeds the capacity of the host forwarding table (16K), the additional VIPs can be hosted on SMux. Traffic data (Figure 2.15) in our production DCs shows that VIP traffic distribution is highly skewed – most of the traffic is destined for a small number of "elephant" VIPs which can be handled by HMux. The remaining traffic to "mice" VIPs can be handled by SMux.

These benefits can only be realized through careful VIP-switch assignment. The assignment must take into account both memory and bandwidth constraints on individual switches, as well as different traffic load of different VIPs. The assignment must dynamically adapt to changes in traffic patterns and network failures. In the next two sections, we describe how Duet solves these problems, as well as provides other load balancing functions.

2.3 VIP Assignment Algorithm

We formalize the VIP-switch assignment problem using the notations listed in Table 2.1.

Input: The input to the algorithm includes the list of VIPs (V), the DIPs for each individual VIP v (d_v), and the traffic volume for each VIP. The latter is obtained from network monitoring. The input also includes the network topology, consisting of a set of switches (S) and a set of links (E). The switches and links constitute the two types of resources (R) in the assignment. Each resource instance has a fixed capacity C_i, *i.e.,* the link bandwidth for a link, and memory capacity that includes residual ECMP and tunneling table capacity available for Duet on a switch. To absorb the

Table 2.1.: Notations used in VIP assignment algorithm.

Notation	**Explanation**
V	Set of VIPs
d_v	Set of DIPs for the v-th VIP
S, E	Set of switches and links respectively
R	Set of resources (switches and links)
C_i	Capacity of i-th resource
$t_{i,s,v}$	v-th VIP's traffic on i-th link, when it is assigned to s-th switch
$L_{i,s,v}$	load (additional utilization) on i-th resource if v-th VIP is assigned to s-th switch
$U_{i,s,v}$	Cumulative utilization of i-th resource if v-th VIP is assigned to s-th switch
$U_{i,v}$	Cumulative utilization of i-th resource after v VIPs have been assigned
$MRU_{s,v}$	Max. Resource Utilization (MRU) after v-th VIP is assigned to s-th switch

potential transient congestion during VIP migration and network failures, we set the capacity of a link to be 80% of its bandwidth.

Objective: Find the VIP-switch assignment that maximizes the VIP traffic handled by HMux. As explained earlier, this will improve latency and reduce cost by cutting the number of SMux needed. We do not attempt to minimize the extra network propagation delay due to indirection because the propagation delay contributes only less than 30μsec of the 381μsec RTT in our DC.

Constraints: Any VIP-switch assignment should not exceed the capacity of any of the resources.

The VIP assignment problem is a variant of multi-dimensional bin-packing problem [7], where the resources are the bins, and the VIPs are the objects. Multi-dimensional bin-packing problems are NP-hard [7]. DUET approximates it with a greedy algorithm, which works quite well in our simulations based on real topology and traffic load of a large production network.

2.3.1 VIP Assignment

We define the notion of maximum resource utilization (MRU). We have two types of resource – switches and links. MRU represents the maximum utilization across all switches and links.

Algorithm sketch: We sort a given set of VIPs in decreasing traffic volume, and attempt to assign them one by one (*i.e.,* VIPs with most traffic are assigned first). To assign a given VIP, we consider all switches as possible candidates to host the VIP. Typically, assigning a VIP to different switches will result in different MRU. We pick the assignment that results in the smallest MRU, breaking ties at random. If the smallest MRU exceeds 100%, *i.e.,* no assignment can accommodate the load of the VIP, the algorithm terminates. The remaining VIPs are not assigned to any switch – their traffic will be handled by the SMuxes. We now describe the process of calculating MRU.

Calculating MRU: We calculate the additional utilization (load) on every resource for each potential assignment. If the v-th VIP is assigned to the s-th switch, the extra utilization on the i-th link is $L_{i,s,v} = \frac{t_{i,s,v}}{C_i}$ where traffic $t_{i,s,v}$ is calculated based on the topology and routing information as the source/DIP locations and traffic load are known for every VIP. Similarly, the extra switch memory utilization is calculated as $L_{s,s,v} = \frac{|d_v|}{C_s}$, *i.e.,* the number of DIPs for that VIP over the switch memory capacity.

The cumulative resource utilization when the v-th VIP is assigned to the s-th switch is simply the sum of the resource utilization from previously assigned (v-1) VIPs and the additional utilization due to the v-th VIP:

$$U_{i,s,v} = U_{i,v-1} + L_{i,s,v} \tag{2.1}$$

The MRU is calculated as:

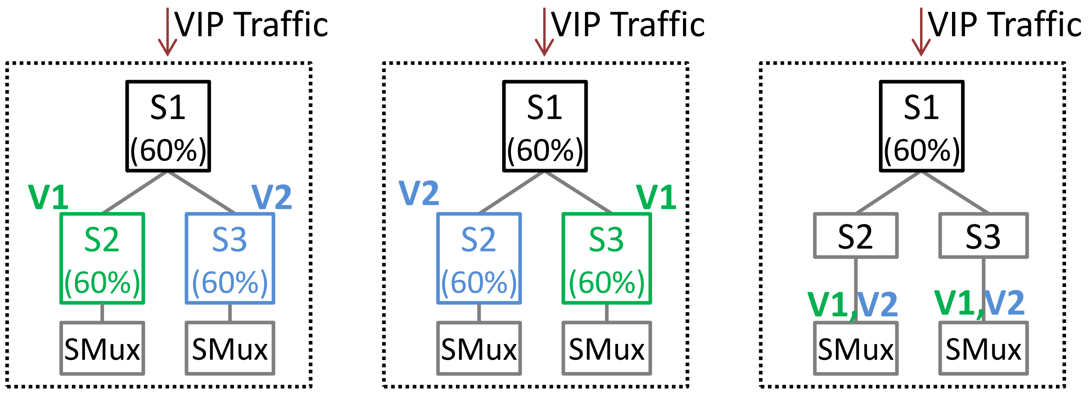
$$MRU_{s,v} = max(U_{i,s,v}), \forall i \in R \tag{2.2}$$

2.3.2 VIP Migration

Due to traffic dynamics, network failures, as well as VIP addition and removal, a VIP assignment calculated before may become out-of-date. From time to time, Duet needs to re-calculate the VIP assignment to see if it can handle more VIP traffic through HMux and/or reduce the MRU. If so, it will migrate VIPs from the old assignment to the new one.

There are two challenges here: (1) how to calculate the new assignment that can quickly adapt to network and traffic dynamics without causing too much VIP reshuffling, which may lead to transient congestion and latency inflation. (2) how to migrate from the current assignment to new one.

A simple approach would be to calculate the new assignment from scratch using new inputs (*i.e.,* new traffic, new VIPs etc.), and then migrate the VIPs whose assignment has changed between the current assignment and the new one. To prevent



(a) Initial (b) Final (c) Through SMux

Fig. 2.4.: Memory deadlock problem during VIP migration. VIPs V1 and V2 both occupy 60% of switch memory each. The goal of migration is to migrate the VIPs from assignment in (a) to (b); DUET eliminates this problem by migrating VIPs through SMuxes, as shown in (c).

routing black holes during VIP migration, we would use make-before-break — *i.e.,* a VIP would be announced from the new switch before it is withdrawn from the old switch. This simple approach is called *Non-sticky*.

The *Non-sticky* approach suffers from two problems. First, it may lead to transitional memory deadlock. Figure 2.4 shows a simple example where initially VIP V1 and VIP V2 are assigned to switches S_2 and S_3, respectively, but swap positions in the new assignment. Further, either VIP takes 60% of the switch memory. Because of limited free memory, there is no way to swap the VIPs under the make-before-break approach. When there are a large number of VIPs to migrate, finding a feasible migration plan becomes very challenging. Second, even if there was no such deadlock, calculating a new assignment from scratch may result in a lot of VIP reshuffling, for potentially small gains.

DUET circumvents transitional memory deadlocks by using SMux as a stepping stone. We first withdraw the VIPs that need to be moved from their currently as-

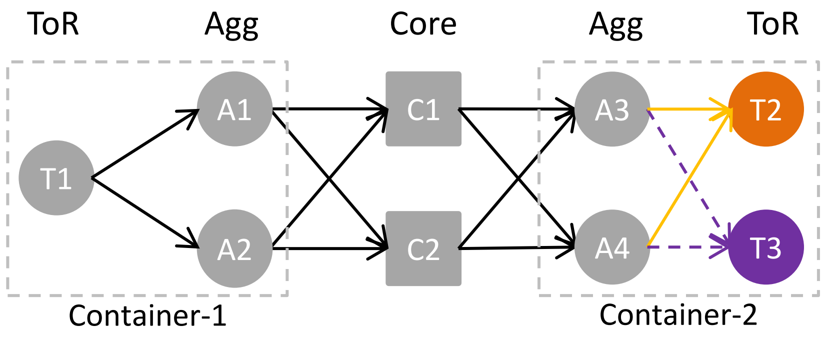


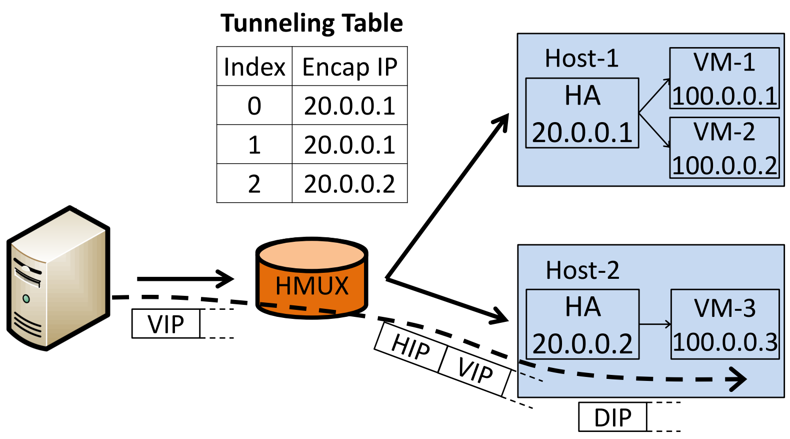
Fig. 2.5.: When the VIP assignment changes from ToR T_2 to T_3, only the links inside container-2 are affected. As a result, we can first select best ToR in a container based on the links within container, and then scan over all containers and remaining Core and Agg switches.

signed switches and let their traffic hit the SMux[3]. We then announce the VIPs from their newly assigned switches, and let the traffic move to the new switches. This is illustrated in Figure 2.4(c) where both VIP's (V1 and V2) traffic is handled by SMux during migration.

Because SMux is used as a stepping stone, we want to avoid unnecessary VIP reshuffling to limit the amount of VIP traffic that is handled by SMux during migration. Hence, we devise a *Sticky* version of the greedy VIP assignment algorithm that takes the current assignment into account. A VIP is moved only if doing so results in significant reduction in MRU. Let us say that VIP v was assigned to switch s_c in the current assignment, and the MRU would be the lowest if it is assigned to switch s_n in the new assignment. We assign v to s_n only if $(MRU_{s_c,v} - MRU_{s_n,v})$ is greater than a threshold. Else we leave v at s_c.

Complexity: It is important for Duet to calculate the new assignment quickly in order to promptly adapt to network dynamics. Since all $L_{i,s,v}$ can be pre-computed, the complexity to find the minimum MRU (Equation 2.2) for VIP-switch assignment is $O(|V| \cdot |S| \cdot |E|)$.

[3]Recall that SMux announces all VIPs to serve as a backstop (§2.2.3)



Tunneling Table

Index	Encap IP
0	20.0.0.1
1	20.0.0.1
2	20.0.0.2

Fig. 2.6.: Load balancing in virtualized clusters.

This complexity can be further reduced by leveraging the hierarchy and symmetry in the data center network topology. The key observation is that assigning a VIP to different ToR switches inside a container will only affect the resource utilization inside the same container (shown in Figure 2.5). Therefore, when assigning a VIP, we only need to consider one ToR switch with the lowest MRU inside each container. Because ToR switches constitute a majority of the switches in the data center, this will significantly reduce the computation complexity to $O(|V| \cdot ((|S_{core}| + |S_{agg}| + |C|) \cdot |E| + |S_{tor}| \cdot |E_c|))$. Here C and E_c denote the containers and links inside a container. S_{core}, S_{agg} and S_{tor} are the Core, Aggregation and ToR switches respectively.

2.4 Practical Issues

We now describe how Duet handles important practical issues such as failures and configuration changes.

2.4.1 Failure Recovery

A critical requirement for load balancer is to maintain high availability even during failures. Duet achieves this primarily by using SMuxes as a backstop.

HMux (switch) failure: The failure of an HMux is detected by neighboring switches. The routing entries for the VIPs assigned to the failed HMux are removed from all other switches via BGP withdraw messages. After routing convergence, packets for these VIPs are forwarded to SMuxes, since SMuxes announce all VIPs. All HMux and SMux use the same hash function to select DIPs for a given VIP, so existing connections are not broken, although they may suffer some packet drops and/or reorderings during convergence time (<40ms, see §2.6.2). Because in our production DCs we rarely encounter failures that are more severe than three switch failures or single container failures at a time, we provision sufficient number of SMuxes to handle the failover VIP traffic from HMuxes due to those failures.

SMux failure: SMux failure has no impact on VIPs assigned to HMux, and has only a small impact on VIPs that are assigned only to SMuxes. Switches detect SMux failure through BGP, and use ECMP to direct traffic to other SMuxes. Existing connections are not broken, although they may suffer packet drops and/or reorderings during convergence.

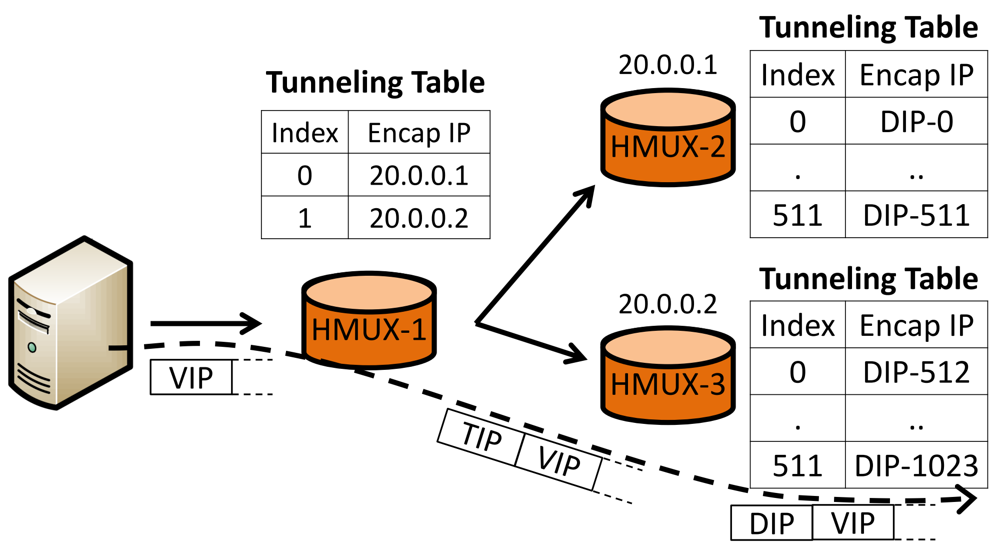
Link failure: If a link failure isolates a switch, it is handled as a switch failure. Otherwise, it has no impact on availability, although it may cause VIP traffic to re-route.

DIP failure: The Duet controller monitors DIP health and removes failed DIP from the set of DIPs for the corresponding VIP. Existing connections to the failed DIP are necessarily terminated. Existing connections to other DIPs for the corresponding VIP are still maintained using resilient hashing [8].

2.4.2 Other Functionalities

VIP addition: A new VIP is first added to SMuxes, and then the migration algorithm decides the right destination.

VIP removal: When a VIP assigned to an HMux is to be withdrawn, the controller removes it both from that HMux and from all SMuxes. VIPs assigned to *only*



Tunneling Table

Index	Encap IP
0	20.0.0.1
1	20.0.0.2

Tunneling Table

Index	Encap IP
0	DIP-0
.	..
511	DIP-511

Tunneling Table

Index	Encap IP
0	DIP-512
.	..
511	DIP-1023

Fig. 2.7.: Large fanout support.

SMuxes need to be removed only from SMuxes. BGP withdraw messages remove the corresponding routing entries from all switches.

DIP addition: The key issue is to ensure that existing connections are not remapped if DIPs are added to a VIP. For VIPs assigned to SMuxes, this is easily achieved, since SMuxes maintain detailed connection state to ensure that existing connections continue to go to the right DIPs. However, HMuxes can only use a hash function to map VIPs to DIPs (Figure 2.2). Resilient hashing only ensures correct mapping in case of DIP *removal* – not DIP *addition*. Thus, to add a DIP to a VIP that is assigned to an HMux, we first remove the VIP from the HMux, causing SMuxes to take it over, as described earlier. We then add the new DIP, and eventually move the VIP back to an appropriate HMux.

DIP removal: DIP removal is handled in a manner similar to DIP failure.

Virtualized clusters: In virtualized clusters, the HMux would have to encapsulate the packet twice – outer header carries the IP of the host (native) machine, while inner header carries IP of the VM hosting the DIP. However, today's switches cannot encapsulate a single packet twice. So, we use HA in tandem with HMux, as shown in Figure 2.6. The HMux encapsulates the packet with the IP of the host machine (HIP) that is hosting the DIP. The HA on the DIP decapsulates the packet and forwards it to the right DIP based on the VIP. If a host has multiple DIPs, the ECMP and tunneling table on the HMux holds multiple entries for that HIP (HIP 20.0.0.1 in Figure 2.6) to ensure equal splitting. At the host, the HA selects the DIP by hashing the 5-tuple.

Heterogeneity among servers: When the DIPs for a given VIP have different processing power, we can proportionally split the traffic using WCMP (Weighted Cost Multi-Path) where faster DIPs are assigned larger weights. WCMP can be easily implemented on commodity switches.

VIPs with large fanout: Typically the capacity of the tunneling table on a single-chip switch is 512. To support a VIP that has more than 512 DIPs, we use indirection, as shown in Figure 2.7. We divide the DIPs into multiple partitions, each with at most 512 entries. We assign a single transient IP (TIP) for each partition. As a VIP, a TIP is a routable IP, and is assigned to a switch. When assigning a VIP to an HMux, we store the TIPs (as opposed to DIPs) in the tunneling table (Figure 2.7). When a packet for such a VIP is received at the HMux, the HMux encapsulates the packet with one of the TIPs and forwards it to the switch to which the TIP is assigned. That switch decapsulates the TIP header and re-encapsulates the packet with one of the DIPs, and forwards it. The latency inflation is negligible, as commodity switches are capable of decapsulating and re-encapsulating a packet at line rate. This allows us to support up to $512 * 512 = 262,144$ DIPs for a single VIP, albeit with small extra propagation delay[4].

Port-based load balancing: A VIP can have one set of DIPs for the HTTP port and another for the FTP port. Duet supports this using the tunneling table and ACL rules. ACL (Access Control) Rules are similar to OpenFlow rules, but currently support a wider range of fields. We store the DIPs for different destination ports at different indices in the tunneling table (Figure 2.8). The ACL rules, match on the IP

[4]The VIP assignment algorithm also needs some changes to handle TIPs. We omit details due to lack of space.

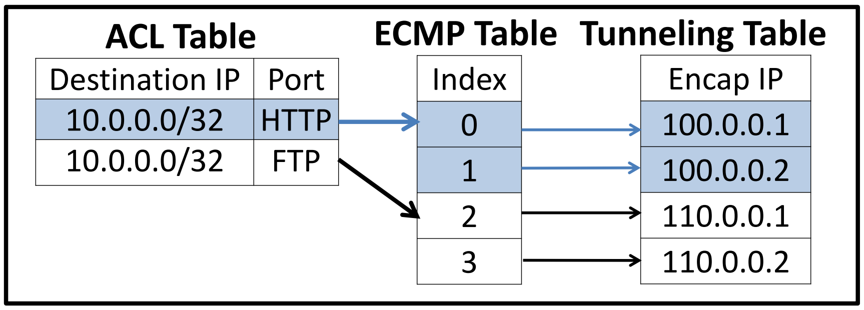


Fig. 2.8.: Port-based load balancing.

destination and destination port fields, and the action is forwarding the packet to the corresponding tunneling table entry. Typically the number of ACL rules supported is larger than the tunneling table size, so it is not a bottleneck.

SNAT: Source NAT (SNAT) support is needed for DIPs to establish outgoing connections[5]. Ananta supports SNAT by maintaining state on SMuxes [2]. However, as discussed earlier, switches cannot maintain such connection state. Instead, DUET supports SNAT by sharing the hash function used by HMux with the host agent (HA). Like Ananta, DUET assigns disjoint port ranges to the DIPs, but unlike Ananta, the HA on the DIP does not randomly choose an unused port number. Instead, it selects a port such that the hash of the 5-tuple would correctly match the ECMP table entry on HMux. The HA can do this easily since it knows the hash function used by HMux. Note that the HA needs to do this *only* during establishment (*i.e.*, first packet) of outgoing connections. If an HA runs out of available ports, it receives another set from the DUET controller.

2.5 Implementation

In this section, we briefly discuss the implementation of the key components in DUET: (1) DUET Controller, (2) Host Agent, and (3) Switch Agent, and (4) SMux, as shown in Figure 2.9.

[5]Outgoing packets on established connections use DSR.

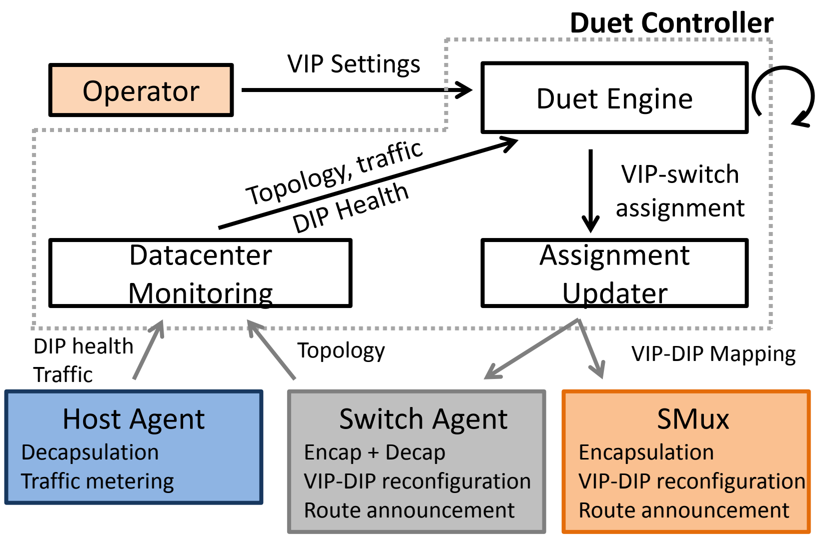


Fig. 2.9.: Components in DUET implementation.

DUET **Controller:** The controller is the heart of DUET. It performs three key functions: (1) Datacenter monitoring: It gathers the topology and traffic information from the underlying network. Additionally, it receives the VIP health status periodically from the host agents. (2) DUET Engine: It receives the VIP-to-DIP mapping from the network operator and the topology and traffic information from the DC-monitoring module, and performs the VIP-switch assignment as described in § 2.3. (3) Assignment Updater: It takes the VIP-switch assignment from the DUET engine and translates it into rules based on the switch agent interface. All these modules communicate with each other using RESTful APIs.

Switch Agent: The switch agent runs on every switch. It uses vendor-specific APIs to program the ECMP and tunneling tables, and provides RESTful APIs which are used by the assignment updater to add/remove VIP-DIP mapping. On every VIP change, the switch agent fires touting updates over BGP.

Host Agent and SMux: The host agent and SMux implementation are the same as in Ananta. The host agent primarily performs packet decapsulation, SNAT and DIP health monitoring. Additionally, the host agents perform traffic metering and report the statistics to the DUET controller.

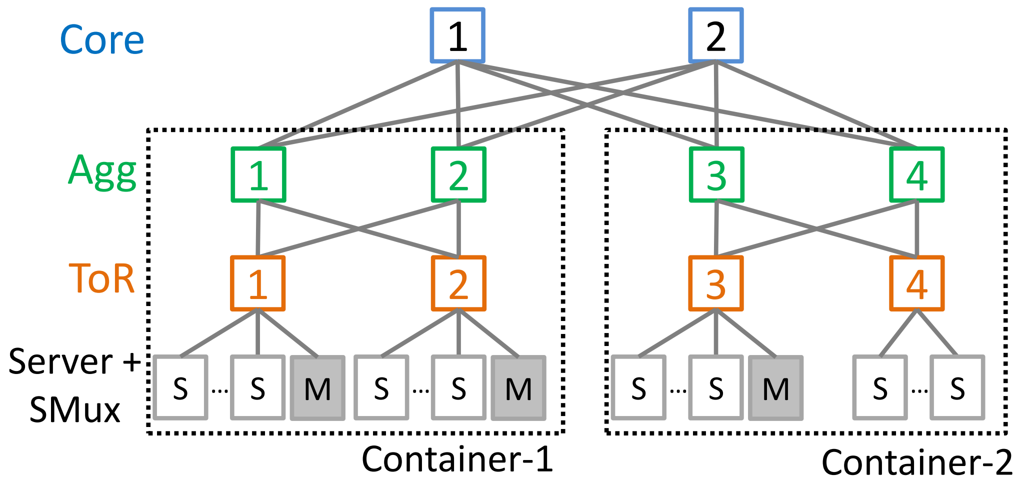


Fig. 2.10.: Our testbed. FatTree with two containers of two Agg and two ToR Switches each, connected by two Core switches.

Same as in Ananta, we run a BGP speaker along side of each SMux to advertise all the VIPs assigned to the SMux.

In total, the controller code consists of 4200 LOC written in C#, and the switch agent code has about 300 LOC in Python.

2.6 Testbed Experiments

Our testbed (Figure 2.10) consists of 10 Broadcom-based switches and 60 servers. Of the 60 servers, 34 act as DIPs and the others are used to generate traffic. Each of ToRs 1, 2 and 3 is also connected to a server acting as SMux.

Our testbed experiments show: (1) HMuxes provide higher capacity, (2) DUET achieves high availability during HMux failure as the VIP traffic seamlessly falls back to SMuxes, and (3) VIP migration is fast, and DUET maintains high availability during VIP migration.

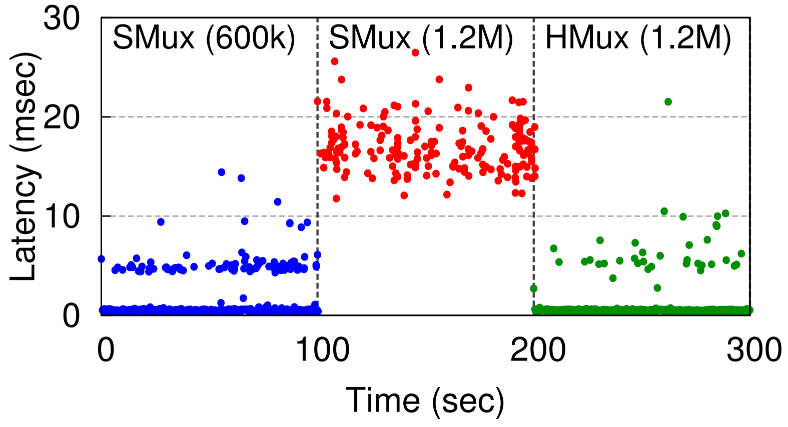


Fig. 2.11.: HMux has higher capacity.

2.6.1 HMux Capacity

If the load balancer instances have low capacity, packet queues will start building up, and traffic will experience high latency. This experiment illustrates that individual HMuxes instances (*i.e.,* a switch) have significantly higher capacity than individual SMux instances.

The experiment uses 11 VIPs, each with 2 DIPs. We send UDP traffic to 10 of the VIPs, leaving the 11th VIP unloaded.

The experiment has three steps. (1) All 11 VIPs are assigned to the SMuxes, and we generate a total traffic of 600K packets per second to the 10 VIPs (60K per VIP). Since each VIP is announced from every SMux, the traffic is split evenly between all SMuxes, and each SMux is handling 200K packets per second. (2) At time 100 sec, we increase the traffic to 1.2M packets per second, so each SMux is handling 400K packets per second. (3) Finally, at time 200 sec, we switch all VIPs to a *single* HMux hosted on ToR 1.

The metric of interest is the latency to the *unloaded* VIP, measured using pings sent every 3ms. We measure the latency to the unloaded VIP so that the latency *only* reflects the delay suffered at the SMux or HMux – the VIP or the DIP itself is not the bottleneck. The results shown in Figure 2.11.

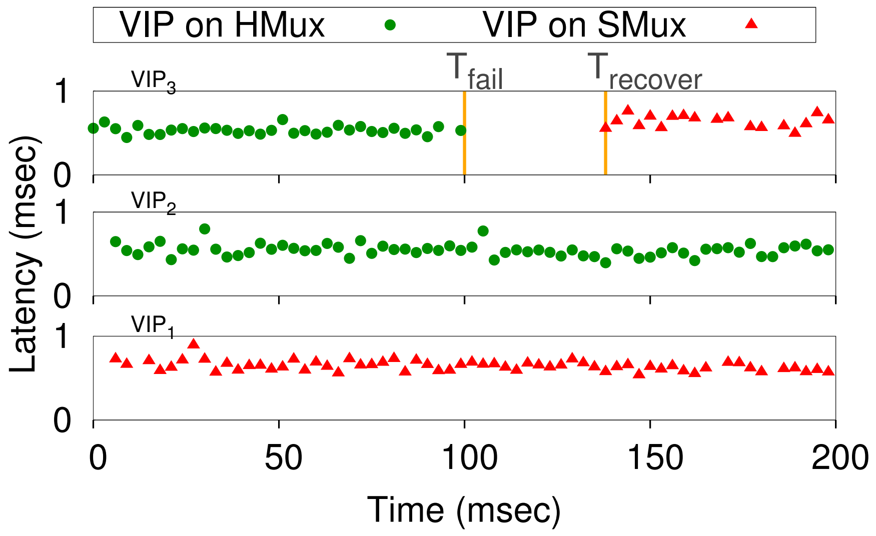


Fig. 2.12.: VIP availability during failure.

We see that until time 100 sec, the latency is mostly below 1ms, with a few outliers. This is because each SMux is handling only 200K packets per second, which is well within its capacity (300K packets per second – see §2.1), and thus there is no significant queue buildup. At time 100, the latency jumps up – now each SMux is handling 400K packets per second, which is well beyond its ability. Finally, at time 200 sec, when all VIPs are on a single HMux, the latency goes down to 1ms again. This shows that *a single* HMux instance has higher capacity than at least 3 SMux instances.

In fact, since HMux processes all packets in the *data plane* of the switch, it can handle packets at line rate, and no queue buildup will occur till we exceed the link capacity (10Gbps in this experiment).

2.6.2 HMux Failure Mitigation

One of the most important benefits of using the SMux as a backstop is automatic failure mitigation, as described in §2.4. In this experiment, we investigate the delay involved in failing over from an HMux to an SMux. This delay is important because *during* failover, traffic to the VIP gets disrupted.

We assign 7 VIPs across HMuxes and the remaining 3 to the SMuxes. We fail one switch at 100 msec. We measure the impact of HMux failure on VIP availability by monitoring the ping latency to all 10 VIPs every 3ms.

Figure 2.12 shows the ping latency for three VIPs: (1) One on the failed HMux (VIP_3), (2) One on a healthy HMux (VIP_2), and (3) One on an SMux (VIP_1), respectively.

We make three observations: (1) The traffic to VIP_3 falls over to SMux within 38 msec after HMux failure. The delay reflects the time it takes for other switches to detect the failure, and for the routing to converge. The VIP was not available during this period, *i.e.,* there is no response to pings. (2) After 38 msec, pings to VIP_3 are successful again. (3) The VIPs assigned to other HMuxes and SMuxes are not affected; their latency is unchanged during HMux failure. These observations demonstrate the effectiveness of using SMux as a backstop in the DUET design.

2.6.3 VIP Migration

Recall that we also use SMux as a backstop during VIP migration. We now investigate the delays involved in this process. This delay is important because it places a lower bound on how quickly DUET can react to network conditions.

In this experiment, we assign 7 VIPs to the HMuxes and the remaining 3 VIPs to the SMuxes. We migrate a VIP from HMux-to-SMux (VIP_1), SMux-to-HMux (VIP_2), and HMux-to-HMux through SMux (VIP_3) at different times. We measure the VIP availability by monitoring the ping latency (every 3ms) to these VIPs, and we also measure the migration delay.

Figure 2.13 shows the ping latency. At time T_1, the controller starts the first wave of migration by sending the migrate command (migrate to SMuxes) to the corresponding switch agents for VIP_1 and VIP_3. It takes about 450ms for the migration

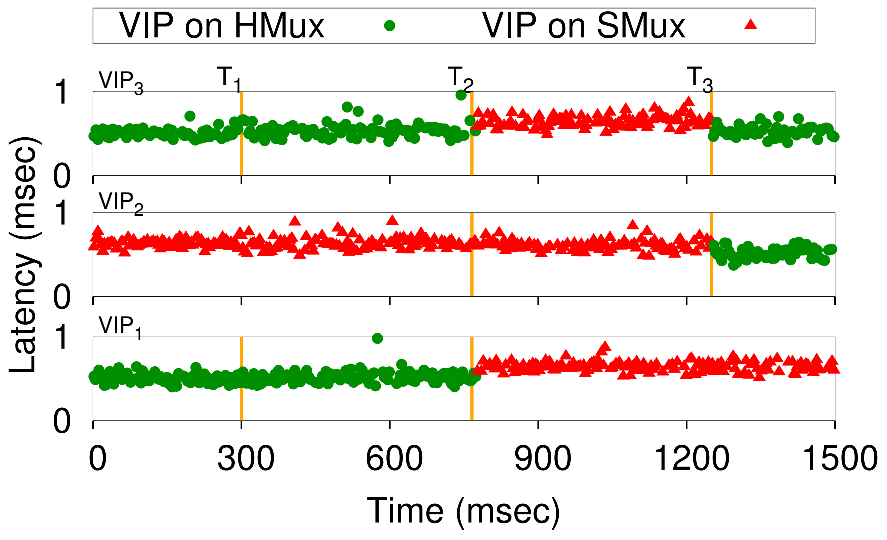
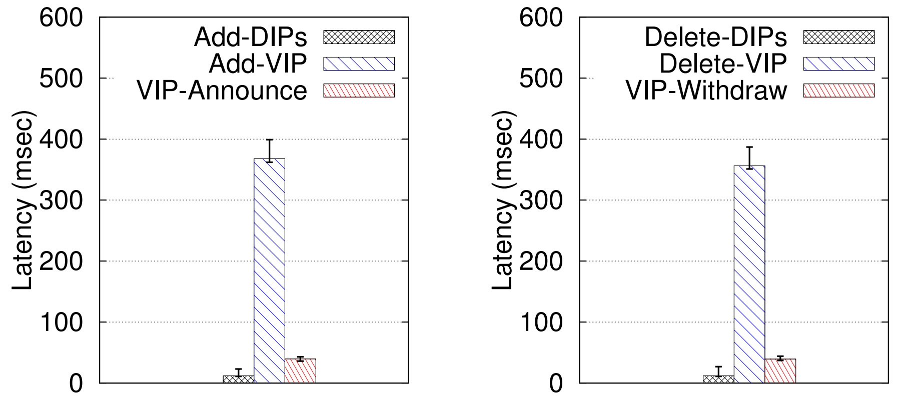


Fig. 2.13.: VIP availability during migration.

to finish (time T_2), at which time, the controller sends another migrate command (migrate back to HMux) to VIP_2 and VIP_3, which takes about 400ms to take effect (time T_3). We see that that all three VIPs remain fully available during the migration process. The VIPs see a very slight increase in latency when they are on SMux, due to software processing of packets on SMux.

Note that unlike the failure scenario discussed earlier, during the migration process, there is no "failure detection" involved. This is why we see no ping packet loss in Figure 2.13.

Figure 2.14 shows the three components of the migration delay: (1) latency to add/delete a VIP as measured from the time the controller sends the command to the time other switches receive the BGP update for the operation, (2) latency to add/delete DIPs as measured similarly as the VIPs, (3) latency for the BGP update (routing convergence), measured as the time from the VIP is changed in the FIB on one switch till the routing is updated in the remaining switches, *i.e.,* BGP update time on those switches.



(a) Add (b) Delete

Fig. 2.14.: Latency breakdown.

Almost all (80-90%) of the migration delay is due to the latency of adding/removing the VIP to/from the FIB. This is because our implementation of the switch agent is not fully optimized – improving it is part of our future work.

2.7 Evaluation

In this section, we use large-scale simulations to show that: (1) DUET needs far fewer SMuxes than Ananta to load balance the same amount of VIP traffic; (2) Despite using fewer SMuxes (and hence being cheaper), DUET incurs low latency on load balanced traffic; (3) The VIP assignment algorithm is effective; (4) Network component failures do not cause significant congestion, even though DUET's VIP assignment algorithm is oblivious to network component failures; (5) The migration algorithm is effective.

2.7.1 Simulation Setup

Network: Our simulated network closely resembles that of a production datacenter, with a FatTree topology connecting 50k servers connected to 1600 ToRs located in 40 containers. Each container has 40 ToRs and 4 Agg switches, and the 40 contain-

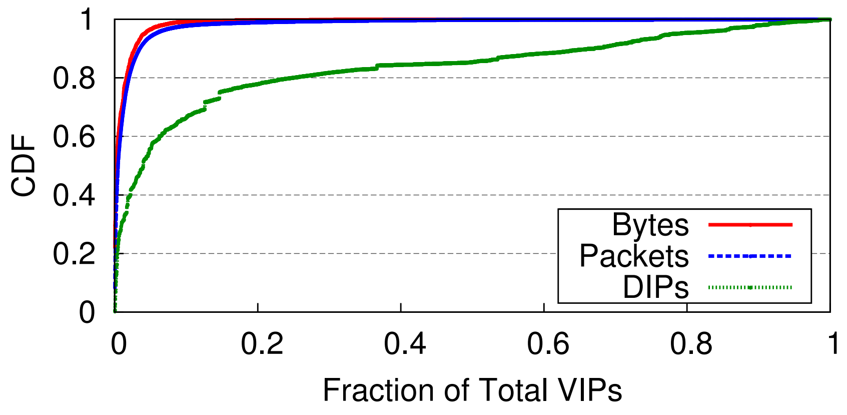


Fig. 2.15.: Traffic and DIP distribution.

ers are connected with 40 Core switches. The link and switch memory capacity were set with values observed in production datacenters: routing table and tunneling table sizes set to 16k and 512, respectively, and the link capacity set to 10Gbps between ToR and Agg switches, and 40 Gbps between Agg and Core switches.

Workload: We run the simulations using the traffic trace collected from one of our production datacenters. The trace consists of 30K VIPs, and the number of DIPs and the traffic distribution across the VIPs are shown in Figure 2.15. We divide the 3-hour trace into 10-minute intervals, and calculate the VIP assignment in each interval, based on the traffic demand matrix (the number of bytes sent and received between all sources and destinations), the topology and the forwarding tables.

2.7.2 SMux Reduction

We first compare the number of SMuxes needed in Duet and Ananta to load-balance same amount of traffic in the datacenter.

We calculate the number of SMuxes needed by Ananta such that no SMux receives traffic exceeding its capacity. We consider two SMux capacities: 3.6Gbps, as observed on the production SMuxes (§2.1), and 10Gbps, assuming the CPU will not be a bottleneck.

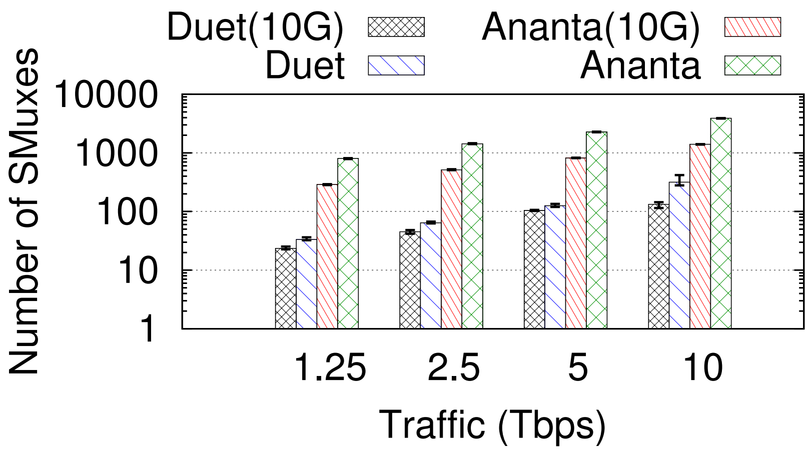


Fig. 2.16.: Number of SMuxes used in Duet and Ananta.

The number of SMuxes needed for Duet depends on the capacity of SMux, the traffic generated by VIPs that could not be assigned to HMuxes, and specifics of failure model, and migration probabilities (§2.2.3). In this experiment, we assign the VIPs to HMuxes using the algorithm described in §2.3, which tries to assign as many VIPs to HMuxes as it can, subject to switch memory and link bandwidth constraints. We have specified the memory and bandwidth details earlier.

Based on failure scenarios in [9,10], we provision the number of SMuxes to handle the maximum traffic under either (1) entire container failure, or (2) three random switch failures. For example, if an entire container fails, the total traffic T to all the VIPs assigned to the switches inside need to fail over to SMuxes. Thus the number of SMuxes needed is $\frac{T}{C_{smux}}$ where C_{smux} is SMux capacity.

We ignore migration – it is covered in §2.7.6.

Figure 2.16 shows that Duet requires far fewer SMuxes compared to Ananta at all traffic rates. *Note the log scale on Y axis.* For all the traffic rates, Duet was able to assign 16k VIPs to the HMuxes (routing table limit). Overall, compared to Ananta, Duet requires 12-24x times fewer SMuxes when the SMux capacity is 3.6 Gbps and 8-12x times fewer SMuxes when the SMux capacity is 10Gbps, across different traffic loads.

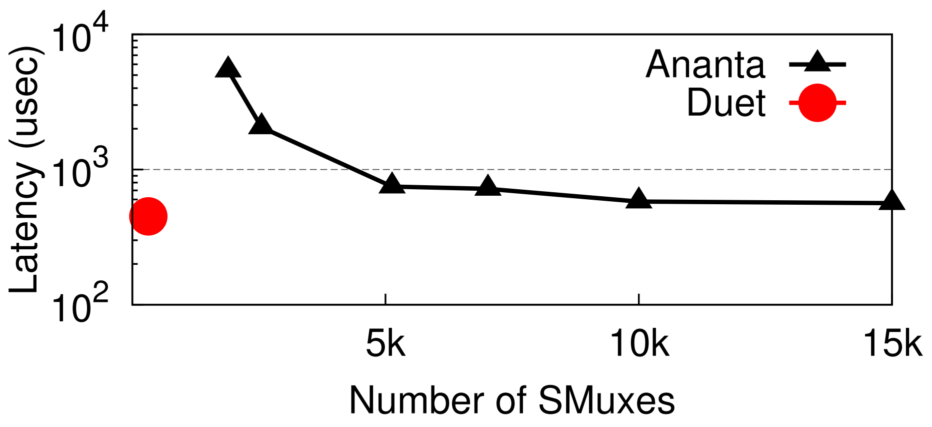


Fig. 2.17.: Latency (microseconds) vs. number of SMuxes in Ananta and DUET.

We note that for all traffic scenarios, majority of the SMuxes needed by DUET were needed to handle failure. The fraction of SMuxes needed to handle the traffic to the VIPs that could not be assigned to the HMux is small. This shows that the VIP assignment algorithm does a good job of "packing" VIPs into HMuxes.

2.7.3 Latency vs. SMuxes

Another way to look at the trade-off described in §2.7.2 is to hold the traffic volume constant, and see how many SMuxes Ananta needs to provide the same latency as DUET. This is shown in figure 2.17.

We hold the traffic at 10Tbps, and vary the number of SMuxes for Ananta from 2000 to 15,000. The black line shows median latency for Ananta. The red dot represents DUET. DUET used 230 SMuxes, and achieved median latency of 474 μsec.

We see that if Ananta were to use the same number of SMuxes as DUET (230), the median latency would be many times higher (over 6 ms). On the other hand, Ananta needs 15,000 SMuxes to achieve latency comparable to DUET.

The absolute latency numbers may appear small – however, recall that median DC RTTs are of the order of 381 μsec[6], and in many cases, to satisfy a single user

[6]Newer technologies such a RDMA lower this to 2-5 μsec!

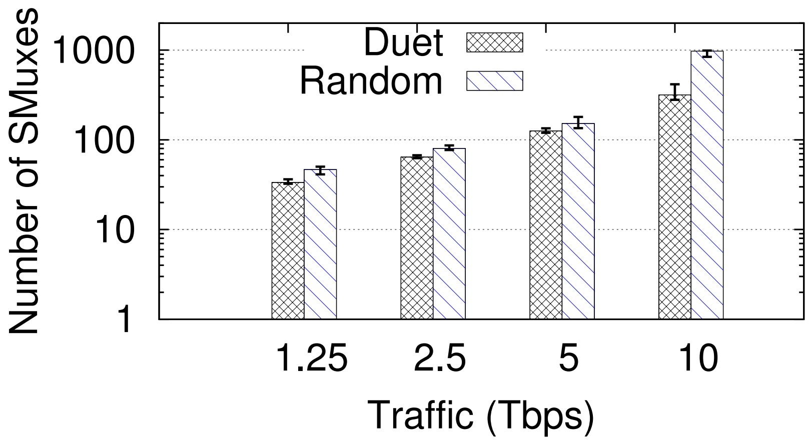


Fig. 2.18.: Number of SMuxes used by Duet and Random.

request, an application like Search traverses load balancer multiple times. Any time lost in the network is *wasted* time – which could have otherwise been used by the application to improve user experience [11–13].

2.7.4 Duet vs. Random

To understand the impact of assigning VIPs based on the maximum resource utilization, we compare the performance of Duet in terms of the number of SMuxes against a random strategy (Random) that selects the *first feasible* switch that does not violate the link or switch memory capacity. This assignment algorithm can be viewed as a variant of FFD (First Fit Decreasing) as the VIPs are assigned in the sorted order of decreasing traffic volume.

Figure 2.18 shows the total number of SMuxes needed by Duet and Random (note the log scale). We see that Random results in 120%–307% more SMuxes compared to Duet as the traffic load varies from 1.25 to 10 Tbps. This shows that by taking resource utilization into account, Duet ensures that only a small fraction of VIPs traffic is left to be handled by the SMuxes.

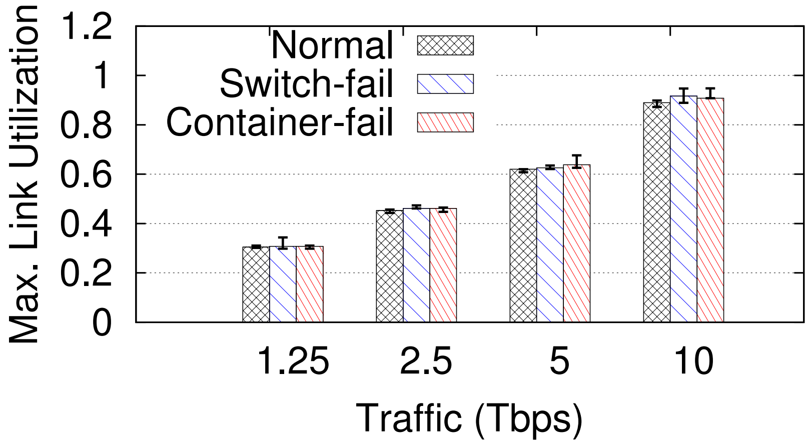


Fig. 2.19.: Impact of failures on max. link utilization.

2.7.5 Impact of Failure

Microbenchmark results in §2.6.2 showed that Duet can handle HMux failures well – the VIPs fall back to SMux, and the disruption to the VIP traffic is minimal. In §2.7.2, we considered the number of SMuxes Duet needs to cope with failures. We now consider the bigger picture – what impact does failures of several switches, or even a container have on overall traffic?

We consider the same failure model as was used in §2.7.2 – a container or up to 3 switches can fail simultaneously. We evaluate failure resilience of Duet by measuring the maximum link utilization under these two scenarios: failure of a randomly selected container, or 3 randomly selected switches.

A random switch failure affects link traffic load in two ways. It causes the traffic of the VIPs assigned to the failed switch to be shifted to the backstop SMuxes, and other through traffic to be shifted to the alternative path. A container failure affects the traffic in more complicated ways: it not only causes all the switches inside to be disconnected, but also makes all the traffic with sources and destinations (DIPs) inside to disappear.

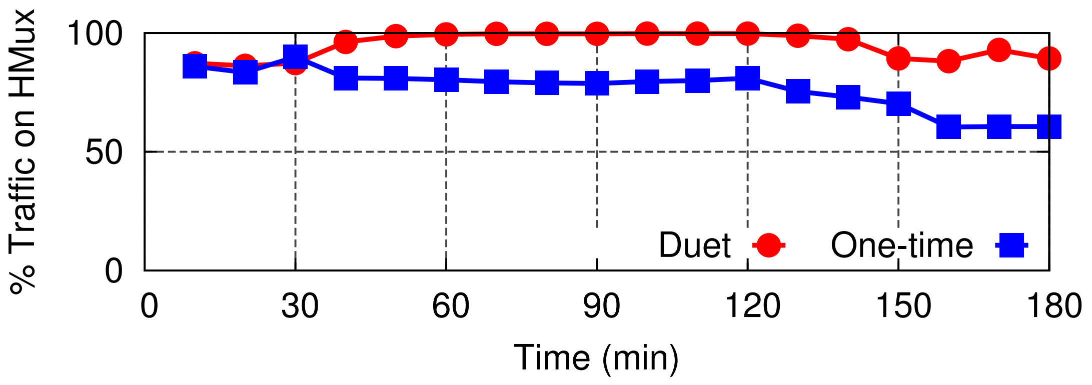
Figure 2.19 shows the measured maximum link utilization during the two failure scenarios in the 10 experiments. We see that as expected, link failures can result in transient congestion. However, the utilization increase of any link in the network is no more than 16%, and hence is comfortably absorbed by the 20% bandwidth reservation made in the VIP assignment algorithm. Interestingly, the single container failure (with 44 switches inside) often results in less congestion than 3-switch failure. This can be explained by two reasons: (1) any traffic with source and sinks (DIPs) inside the container has disappeared, and (2) all the rest traffic which have sources or sinks outside the container are not shifted to other paths as their paths do not go through any switch inside the container.

2.7.6 VIP Migration

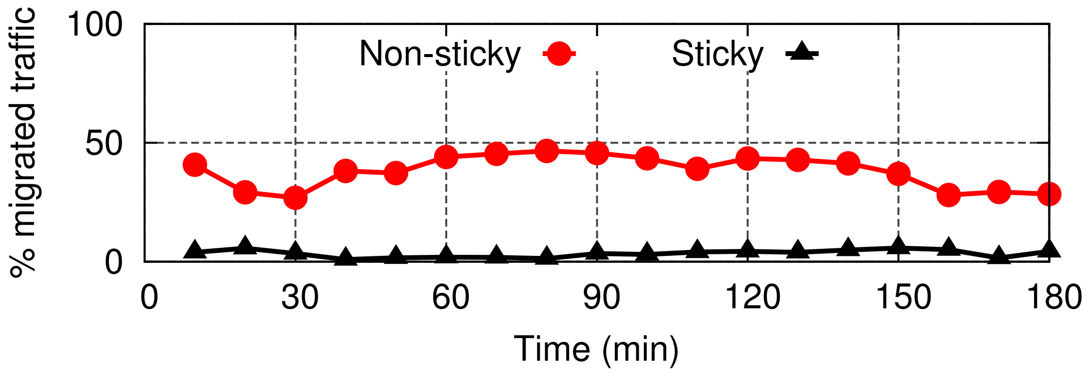
In this section, we evaluate the effectiveness of DUET's VIP migration algorithm, *Sticky* (§2.3.2). We set the threshold to be $\delta = 0.05$, *i.e.,* a VIP will migrate to a new assignment only if doing so reduces the MRU by 5%.

We compare *Sticky* with *Non-sticky*, which calculates the new assignment from scratch based on current traffic matrix (§2.3.1), but migrates all the VIPs at the same time through SMuxes to avoid the memory deadlock problem. We evaluate these two schemes by re-running the 3-hour traffic trace, where we reassign and migrate the VIPs for *Sticky* and *Non-sticky* every 10 minutes. The total VIP traffic varies between 6.2 to 7.1 Tbps in this trace.

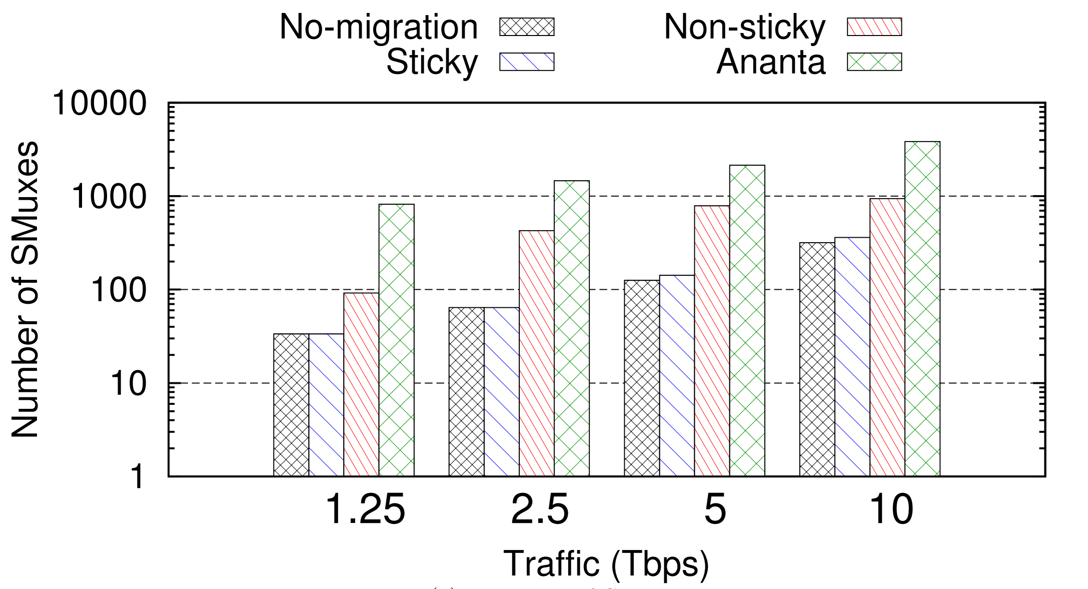
Effectiveness: We first compare the portion of total traffic that are handled by the HMuxes under the two assignment schemes – the larger the portion, the more effective the assignment algorithm. Here, we also compare *Sticky* and *Non-sticky* against *One-time* algorithm, which assigns the VIPs at time 0 sec, and never change it. Figure 2.20(a) shows the results over the duration of the trace. First, as expected, while the portion of traffic handled by HMuxes started out the same, the initial assignment which is used by *One-time* throughout the trace, gradually loses its effectiveness, and results in only 60-89% (average 75.2%) of the total being handled by HMuxes. In contrast, *Sticky* and *Non-sticky* handle 86-99.9% (average 95.3%) of



(a) Traffic load-balanced by HMux



(b) Traffic shuffled during migration



(c) Number of SMux

Fig. 2.20.: Effectiveness of different migration algorithms.

the traffic in HMuxes, from continuously adapting to the traffic dynamics. Second, even though *Sticky* only migrates VIPs that reduce the MRU by at least 5%, it is as effective as *Non-sticky* in maximizing the traffic assigned to HMuxes. In particular, it handles 86-99.7% traffic (average 95.1%) in HMuxes, which is almost identical to the 87-99.9% traffic (average 95.67%) handled by HMuxes under *Non-sticky*.

Traffic shuffled: Next, we compare the fraction of the total VIP traffic migrated under *Sticky* and *Non-sticky*– the less traffic are migrated, the fewer SMuxes need to be reserved as stepping stone. Figure 2.20(b) shows that migration using *Non-sticky* results in reshuffling almost 25-46% (average 37.4%) of the total VIP traffic each time throughout the trace duration, compared to only 0.7-4.4% (average 3.5%) under *Sticky*. Such a drastic reduction in the traffic shuffled under *Sticky* is attributed to its simple filtering scheme: a VIP is only migrated if it improves the MRU by 5%.

Number of SMuxes: Figure 2.20(c) shows the number of SMuxes needed by *Sticky* and *Non-sticky*. Additionally, we also calculate the SMuxes needed without migration (marked as No-migration) as well as number of SMuxes needed in Ananta considering the SMux capacity to 3.6Gbps. The number of SMuxes needed in *Sticky* and *Non-sticky* is calculated as maximum of SMuxes needed for VIP traffic, failure and transition traffic. It can be seen that, *Non-sticky* always requires more SMuxes compared to No-migration and *Sticky*, showing that *Sticky* does not increase the number of SMuxes to handle the traffic during migration.

2.8 Discussion

Why are there empty entries in switch tables? Duet uses empty entries in the host table, ECMP table, and tunneling table in switches to implement HMux. Several reasons contribute to the abundance of such free resources in our production datacenter. The host table of ToR switches has only a few dozen entries for the hosts within each rack, and that of the rest of the switches is mostly empty. The ECMP table of switches is mostly empty because of the hierarchical DC network topology, where each switch has a small number of outgoing links among which all outgoing traffic is split via ECMP. The tunneling table is mostly free since few online services use encapsulation other than load balancing itself. We acknowledge that other DCs may have a different setup, but we believe that our design will be applicable in common cases.

VIP assignment: While the greedy VIP assignment algorithm described in §2.3 works well in our scenarios, we believe that it can be improved. The VIP assignment problem resembles bin packing problem, which has many sophisticated solutions. We plan to study them in future. Also, while we consider VIPs in order of traffic, other orderings are possible (*e.g.,* consider VIPs with latency sensitive traffic first).

Failover and Migration: DUET relies on SMuxes to simplify failover and migration. As hinted in §2.2.3, it may be possible to handle failover and migration by replicating VIP entries in multiple HMuxes. We continue to investigate this approach, although our initial exploration shows that the resulting design is far more complex than our current design.

2.9 Related Work

To the best of our knowledge, DUET is a novel approach to building a performant, low-cost, organically scalable load balancer. We are not aware of any load balancing architecture that fuses switch-based load balancer with the software load balancers. However, there has been much work on load balancers, and we briefly review it here.

Load balancer: Traditional hardware load balancers [14, 15] are expensive and typically only provide 1+1 availability. DUET is much more cost effective, and provides enhanced availability by using SMuxes as a backstop. Importantly, compared to traditional load balancers, DUET gives us control over very important vantage point in our cloud infrastructure.

We have already discussed Ananta [2] software load balancer extensively. Other software-based load balancers [16–18] are also available, but they lack the scalability and availability of Ananta, as shown in [2]. Embrane [19] promises scalability, but suffers from the same fundamental limitations of the software load balancer.

OpenFlow based load balancer: Two recent proposals focus on using OpenFlow switches for load balancing. In [20], authors present a preliminary design for a load balancing architecture using OpenFlow switches. They focus on minimizing the number of wildcard rules. The paper, perhaps because it is a preliminary design, ignores many key issues such as handling switch failures. Plug-n-Serve [21] is another preliminary design that uses OpenFlow switches to load balance web servers deployed in unstructured, enterprise networks. Duet is very different from these approaches. Duet uses a combined hardware and software approach. Duet does not rely on OpenFlow support. Duet is designed for data center networks, and pays careful attention to handling numerous practical issues including various types of failures and VIP migration to adapt to network dynamics.

Partitioning OpenFlow rules: Researchers have also proposed using OpenFlow switches for a variety of other purposes. For example, DIFANE [22] uses some switches in the data center to cache rules, and act as authoritative switches. While a load balancing architecture can be built on top of DIFANE, the focus of the paper is very different from Duet. In vCRIB [23] authors propose to offload some of the traffic management rules from host agent to ToR switches, as well as to other host agents. Their goal is to ensure resource-aware and traffic-aware placement of rules. While vCRIB also faces problems such as managing network dynamics (*e.g.,* VM migration), their main focus is quite different than Duet.

SDN architecture and middleboxes: Similar to Duet, researchers have leveraged SDN architecture in the context of middleboxes to achieve policy enforcement and verification [24, 25], which is again a different goal than Duet.

Improving single server performance: Researchers have substantially improved packet processing capabilities on commodity servers [26, 27], which could potentially improve SMux performance. But, these improvements are unlikely to bridge the differences in packet processing capabilities between HMux and SMux for the load balancer workload.

Lastly, several algorithms for calculating flow hashes (*e.g.,* resilient hashing [8], cuckoo-hashing [26]) offer a wide variety of trade-offs. We do not review them here, although DUET can leverage any advances in this field.

2.10 Summary

DUET is a new distributed hybrid load balancer designed to provide high capacity, low latency, high availability, and high flexibility at low cost. The DUET design was motivated by two key observations: (1) software load balancers offer high availability and high flexibility but suffer high latency and low capacity per load balancer, and (2) commodity switches have ample spare resources and now also support programmability needed to implement load balancing functionality. The DUET architecture seamlessly integrates the switch-based load balancer design with a small deployment of software load balancer. We evaluate DUET using a prototype implementation and extensive simulations using traces from our production DC. Our evaluation shows that DUET provides 10x more capacity than a software load balancer, at a fraction of its cost, while reducing the latency by over 10x, and can quickly adapt to network dynamics including failures.

3. RUBIK: UNLOCKING THE POWER OF LOCALITY AND END-POINT FLEXIBILITY IN CLOUD SCALE LOAD BALANCING

3.1 Introduction

Despite the high capacity and low cost benefits in Duet, Duet and other load balancer designs incur high bandwidth usage of the DC network because of the intrinsic nature of traffic redirection. First, even if the traffic source and the DIPs that handle the traffic are under the same ToR, the traffic first has to be routed to the Muxes, which may be faraway and elongate the path traveled by the traffic. Second, in both Ananta and Duet, the Muxes select DIPs for a VIP by hashing the five-tuple of IP headers, and hence are oblivious to DIP locations. As a result, even if the Mux and some DIPs are located nearby the source, the traffic can be routed to faraway DIPs in the DC, again traversing longer paths. Lastly, these designs do not leverage the server location flexibility in placing the DIPs closer to the sources to shorten the path. The second problem with the Duet LB design is that the traffic detouring through core links breaks the full-bisection bandwidth guarantees originally provided by full-provisioned networks such as Clos and FatTree.

Our evaluation of traffic paths in a production DC network shows that such traffic detour significantly inflates the bandwidth usage of the DC network. This high bandwidth usage not only requires the DC operator to provision high network bandwidth which is costly, but also makes the network prone to transient congestion which affects latency-sensitive services.

In this chapter, we first quantify the network bandwidth overhead in Duet and other prior load balancer designs. We propose Rubik, a new LB that significantly reduces the high bandwidth usage by LB. Like Duet, Rubik uses a hybrid LB design consisting of the HMuxes and SMuxes, and aims to maximize the VIP traffic handled by HMuxes to reduce the LB costs. While doing that, RUBIK reduces the bandwidth usage using two synergistic design principles. First, RUBIK exploits the locality, *i.e.,* it tries to load balance VIP traffic generated within individual ToRs across the DIPs residing in the same ToRs. This reduces the total traffic entering the core network. Second, RUBIK exploits end-point flexibility, *i.e.,* it tries to place the DIPs for a VIP in the same ToRs as the sources generating the VIP traffic.

3.2 Background

In this section, we briefly explain LB workloads and quantify their impact on the network bandwidth usage.

3.2.1 VIP traffic

In the Azure DC, 18-59% (average 44%) of the total traffic is VIP traffic which requires load balancing [2]. This is because services within the same DC use VIPs to communicate with each other to use the benefits provided by the VIP indirection. As a result, all incoming Internet traffic to these services (close to 30% of the total VIP traffic in our DC) as well as a large amount of inter-service traffic (accounting for 70% of the total VIP traffic) go through the LB. For a DC with 40k servers, LB is expected to handle 44 Tbps of traffic at full network utilization [2]. Such indirection of large traffic volume requires a scalable, high performance (low latency, high capacity) and highly available LB.

3.2.2 Workload Characteristics

We make the following observations about the VIP traffic being load balanced in our production DC, by analyzing a 24-hour traffic trace, for 30K VIPs.

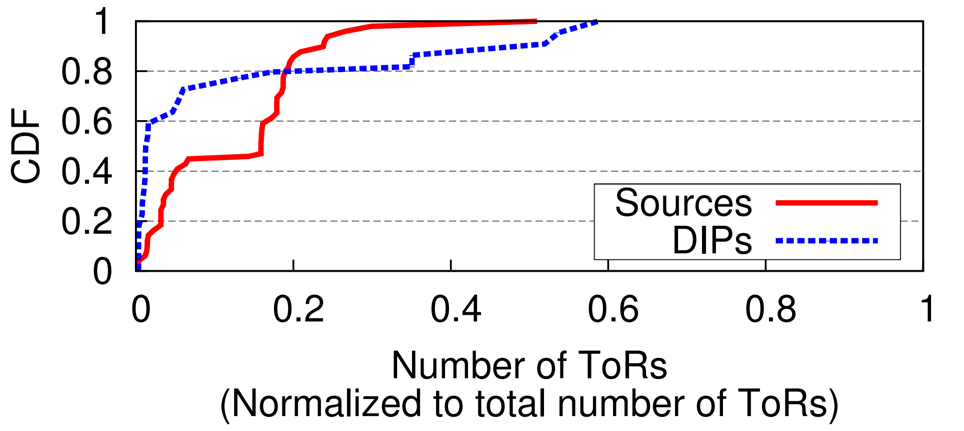


Fig. 3.1.: Distribution of the number of ToRs where the sources and DIPs are located.

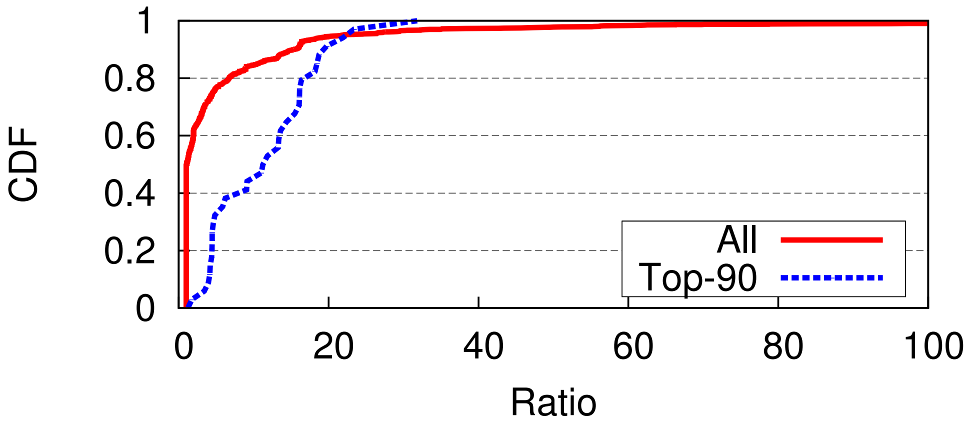


Fig. 3.2.: Ratio of 99^{th} percentile to average traffic volume for each VIP across all sources.

The traffic sources and DIPs for individual VIPs are scattered over many ToRs. Fig. 3.1 shows the number of ToRs where the traffic sources and DIPs for the top 10% VIPs which generate 90% of the total VIP traffic are located. We see that traffic sources are widely scattered – the number of ToRs generating traffic for each VIP varies between 0-44.5% of the total ToRs. Also, the number of ToRs where the DIPs for a VIP are located varies between 0-58% of the total ToRs in the DC.

The traffic volume of sources per VIP are highly skewed. We measure the traffic from all the ToRs for each VIP. Figure 3.2 shows the CDF of the ratio of the 99^{th} percentile to the median per-ToR traffic volume for each VIP. We see that the source traffic volume for each VIP is highly skewed – the ratio varies between 1-35 (median 18) for the top VIPs generating 90% of the total traffic. The large skew happens for multiple reasons, including different numbers of servers, skew in the popularity of the objects that are served, and locality [28, 29].

VIP dependencies: Many large-scale web services are composed of multi-tier services, each with its own set of VIPs. When the top-level service receives a request, it spawns multiple requests to the services at the second tier, which in turn send requests to services at lower tiers. As a result, the VIP traffic exhibit hierarchical dependencies – the DIPs serving the VIPs at tier i become the traffic sources for the VIPs at tier $(i + 1)$. We observe that 31.1% VIPs receive traffic from other VIPs. These VIPs employ 25.1% of the total DIPs and contribute to 27.6% of the total VIP traffic. The remaining 72.4% VIP traffic comes from the Internet, other DCs, and other servers in the same DC that are not assigned to any VIPs.

The dependency among the VIPs can be represented in a DAG. The depth of the DAG observed is similar to the depths reported by Facebook and Amazon [30].

3.3 Motivation

We next assess the impact of the VIP traffic characteristics (§3.2.2) on the DC network bandwidth usage under the Duet LB. We simulate how Duet handles the VIP traffic using a 24-hour traffic trace from our production DC on a network topology that closely resembles our production DC. The topology, workload, and results are detailed in §3.10. Duet maximizes the total traffic handled by HMuxes, on average 97% in the 24-hour period.

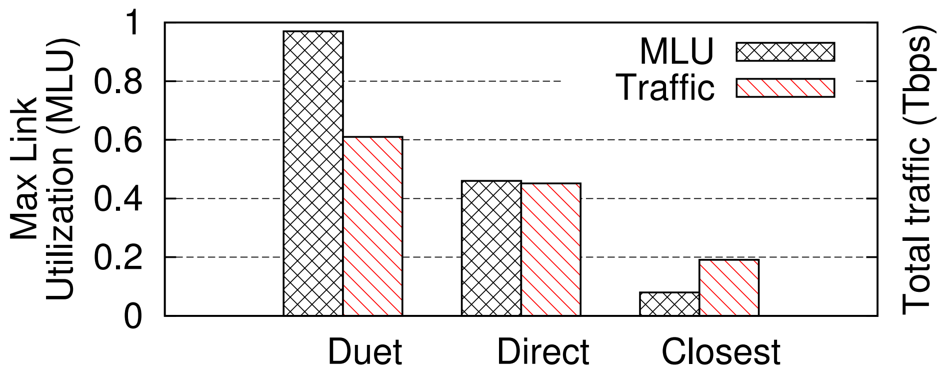


Fig. 3.3.: MLU and total traffic under various LB schemes. Total traffic is measured across all the DC network links.

Table 3.1.: Path length for different LB designs.

Duet	Direct	Closest
5.47	3.94	1.78

High link utilization Figure 3.3 shows the MLU and total traffic in the DC network[1]. While Duet is able to handle 97% of the total VIP traffic by leveraging HMuxes, it also inflates the MLU to 0.98 (or 98%). This high MLU can be explained by two design decisions of Duet.

First, Duet assigns a VIP only to a single HMux. But the traffic sources and DIPs for individual VIPs are spread in a large number of ToRs (Figure 3.1). The diverse location of traffic sources and DIPs per VIP suggests *no matter where the single Mux for a VIP is positioned in the network, it will be far away from most of the traffic sources and DIPs for that VIP*, and hence most VIP traffic will traverse through the network to reach the HMuxes and then the DIPs, which inflates the path length between the sources and DIPs.

Table 3.1 shows that the average number of hops between the sources and the DIPs across all individual VIPs is 5.47 in Duet. Notice that the traffic between two hosts

[1]Absolute values for "total traffic" are omitted for confidentiality.

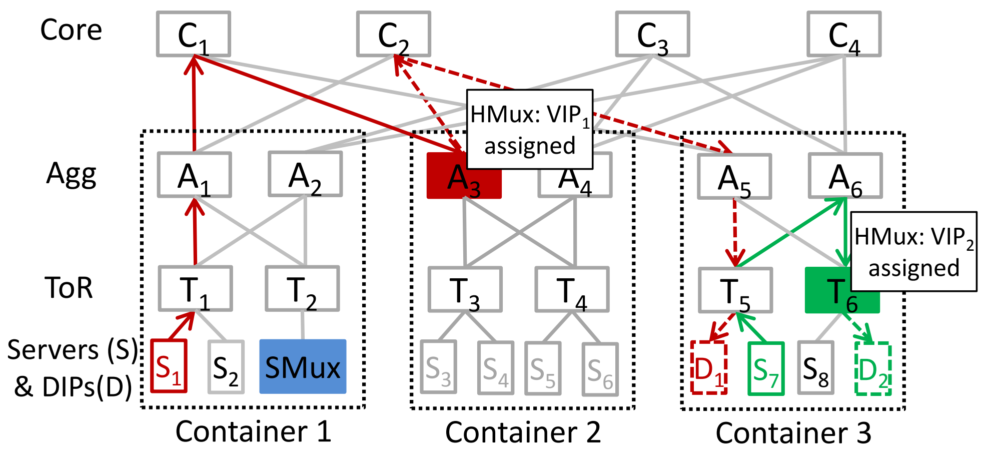


Fig. 3.4.: Duet architecture. Links marked with solid and dotted lines carry VIP and DIP traffic, respectively.

that does not go through the LB would have a maximum of 4 hops (ToR-Agg, Agg-Core, Core-Agg, Agg-ToR). Thus the average path length of 5.47 in Duet indicates that most traffic goes through the core links and further experiences some detour in the DC network. Figure 3.4 shows an example where the VIP-1 traffic originated at S_1 has to travel 6 hops to reach DIP D_1 – 3 hops to reach the HMux at switch A_3, and 3 more hops to reach D_1.

To dissect the impact of the redirection, we measure the MLU and total traffic in the DC network in a hypothetical case where the HMuxes are located on a direct path between the sources and DIPs, labeled as "Direct". Figure 3.3 shows that in this case the MLU is reduced to 0.46 (from 0.98 in Duet), and the bandwidth used is lowered by 1.36x, compared to Duet. Also, the average path length in "Direct" is lowered to 3.94 (1.38x improvement). This means the redirection design in Duet inflates the MLU by 2.13x and bandwidth used by 1.36x.

The second cause for the high link utilization is location-oblivious DIP selection in Duet. The HMux splits the VIP's traffic by hashing on the 5-tuples in the IP header, and chooses the DIP based on the hash. Thus, even if there is a DIP located under the same ToR as the HMux and has the capacity to handle all the local traffic for the VIP, the HMux will spread the local traffic among all DIPs, many of which can be far away in the DC.

To measure the impact of location-oblivious DIP selection, we measure the MLU and bandwidth used in a hypothetical case, where the traffic from the individual sources is routed to the closest DIP and assuming the HMuxes lie on the path. This mechanism is labeled as "Closest". Figure 3.3 shows that the MLU is reduced to just 0.08, and the bandwidth used reduces by 3.19x compared to Duet. Also, the average path length is lowered to just 1.78 hops.

Effective full bisection bandwidth reduced at core Many DC networks have adopted topologies like FatTree and Clos [31] to achieve full-bisection bandwidth. Such networks guarantee that there is enough aggregate capacity between Core and Agg switches as between Agg and ToR switches, and hence the core links will never become a bottleneck for any traffic between the hosts.

However, traffic indirection can break this assumption, if the HMuxes reside in Agg or ToR switches. This happens to Duet, as Duet considers all the switches while assigning VIP-to-DIP mappings. This is illustrated in Figure 3.4. When VIP_1 is assigned to an Agg switch (A_3), the traffic from source S_1 travels the core links twice en-route to DIP D_1 – first to get to HMux A_3, and then to D_1. In contrast, direct host-to-host traffic only has to traverse core links at most once. As a result, the effective bandwidth in the core links is reduced – in Figure 3.4, the available bandwidth to container-2 (servers S_3-S_6) is reduced due to the LB traffic among other containers.

Our evaluation in §3.10.3 shows the traffic overhead in Duet, *i.e.,* the ratio of the additional traffic due to redirection to the total traffic without redirection is 44% in core links and 16% in containers. This means the remaining bisection bandwidth of the Agg-Core links is lower than the remaining bisection bandwidth in the ToR-Agg links. This breaks the full-bisection guarantee provided by the FatTree or Clos, which jeopardizes other applications that co-exist in the DC and assume full-bisection bandwidth is available (*e.g.,* [32, 33]).

3.4 Rubik Overview

In the previous section, we saw that the traffic indirection in Duet incurs substantial overhead in the DC network bandwidth usage. In this thesis, we propose a new LB design, Rubik, that significantly reduces the bandwidth usage in the DC network while providing low cost, high performance and high availability benefits.

Rubik is based on two key ideas motivated by the observations in the last section. First, it exploits *locality*, *i.e.,* it tries to load balance traffic generated in individual ToRs across the DIPs present in the same ToRs. In this way, a substantial fraction of the load balanced traffic will not enter the links beyond ToRs which reduces the DC network bandwidth usage and MLU.

The second key idea of Rubik is to exploit *DIP placement flexibility* to place DIPs closer to the sources. In Rubik online services specify the number of DIPs for individual VIPs, and Rubik decides the location of the servers to be assigned to individual VIPs. This idea is synergistic with the first idea, as it facilitates exploiting locality in load balancing within ToRs.

Realizing the two ideas is challenging, because (1) there are a limited number of servers in each ToR where DIPs can be assigned, (2) switches have limited memory for storing VIP-to-DIP mappings, (3) a VIP may have traffic sources in more ToRs than the total number of DIPs for that VIP. In such a case, a DIP cannot be assigned in every ToR that has traffic sources, (4) dependencies between the VIPs make it even harder, as the sources to some of the VIPs are not known until DIPs for other VIPs are placed.

Rubik addresses the above capacity limitations (switch memory and DIPs in a ToR) using two complimentary ideas. First, Rubik uses a new LB architecture that splits the VIP-to-DIP mapping for a VIP across multiple HMuxes to *enable* efficient use of switch memory while containing local traffic. Second, it employs a novel algorithm that *calculates* the most efficient use of switch memory for containing the most local traffic.

3.5 Rubik Architecture

Rubik uses a new LB design that splits the VIP-to-DIP mapping for each VIP into *multiple local and a single residual VIP-to-DIP mappings.* This idea is inspired by the observation that the traffic for individual VIPs is skewed (§3.2.2) – some ToRs generate more traffic than other ToRs for a given VIP. In Rubik, we assign local mappings to the ToRs generating large fractions of the traffic and also assign enough DIPs to handle those traffic. The local mapping for a VIP load balances traffic for that VIP across the DIPs present under the same ToR (called local DIPs). We then assign a single residual mapping for that VIP to handle the traffic from all the remaining ToRs, where no local mapping is assigned.

Effectively, the VIP-to-DIP mapping for a VIP is split across the local and residual mappings such that a single DIP appears in only one mapping. Assigning a DIP only once makes the most efficient use of the limited tunneling table space of HMuxes so the total VIP traffic handled by the HMuxes can be maximized. The assignment module (§3.6) then calculates the actual assignment that maximizes the VIP traffic handled locally.

We now explain the Rubik architecture in detail.

Local mapping If some of the sources and DIPs for a VIP already reside in the same ToR, Rubik exploits this locality by load balancing the source traffic across those local DIPs. To ensure that the traffic does not flow outside the ToR in detouring through the HMux, Rubik stores a subset of the VIP-to-DIP mapping, *i.e.,* containing only the local DIPs, at the ToR itself (*e.g.,* HMux T_2 in Fig. 3.5). We denote such a mapping containing the subset of local DIPs as a *local mapping.*

Residual mapping For an individual VIP, we assign a single residual mapping to handle the remaining traffic not handled by the local mappings (called residual traffic). We pool all the remaining DIPs for a VIP together in a single DIP-set, called the *residual mapping* for that VIP (*e.g.,* HMux on C_1 and A_5 in Fig. 3.5). The residual mapping for each VIP announces the VIP using BGP so that other

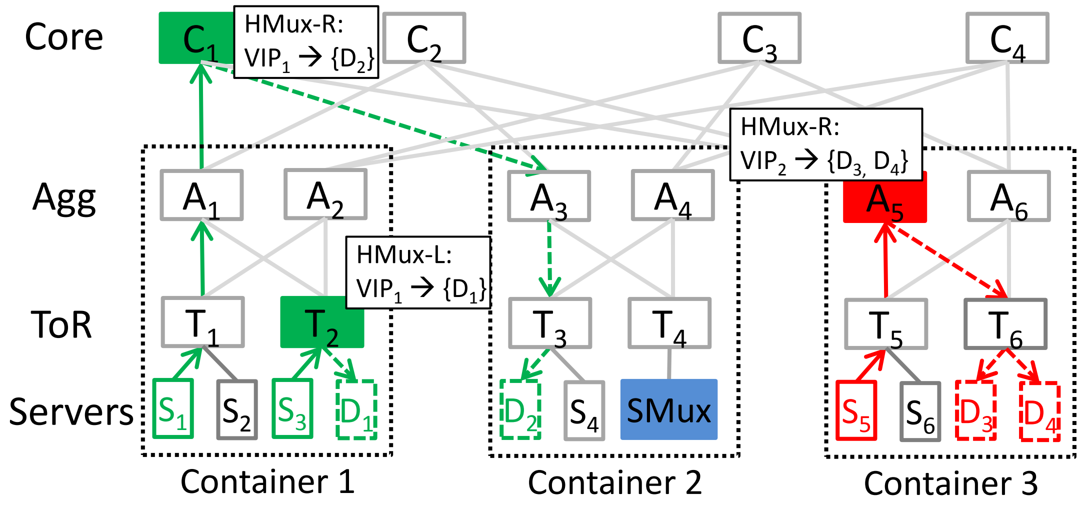


Fig. 3.5.: Rubik Architecture. DIPs for VIP_1 are split in local (HMux-L) and residual mappings (HMux-R).

routers (or switches) route the VIP traffic to the HMux where its residual mapping is assigned.

In principle, we can replicate the residual mapping at all the ToRs containing any remaining traffic sources. Such replication can reduce the number of hops between the sources and HMux, but it can also consume a significant amount of the limited tunneling table space. Therefore, we only assign the residual mapping for a VIP to a single HMux, and the optimal choice of HMux to store the residual mapping of a VIP depends on the location of the remaining traffic sources and residual DIPs.

VIP routing The above DIP-set splitting design has one potential problem. If the HMuxes storing either the local mappings or the residual mapping of a VIP all announce the VIP via BGP to the network, some of the residual source traffic may be routed towards the HMuxes storing local mappings if they are closer than the HMux storing the residual mapping. This would significantly complicate the DIP placement, and DIP-set splitting and placement problem. We avoid this complication by making the HMuxes storing local mappings *not* announce the VIP via BGP. In this way, only local source traffic within a ToR sees the local mapping and is split to the local DIPs.

SMuxes Because of the limited switch memory, the numbers of VIPs and DIPs supported by HMuxes remain limited. Current HMuxes can support up to 16K VIPs [4], and our DC has 30K+ VIPs. Also, it remains challenging to provide high availability during HMux failures. We address both problems by deploying a small number of SMuxes as a backstop, to handle the VIP traffic that could not be handled using HMuxes. We also announce all the VIPs from all SMuxes. We use Longest prefix matching (LPM) to: (1) preferentially route the VIP traffic to the HMuxes for the VIPs assigned to both HMuxes and SMuxes, (2) route the traffic to the remaining VIPs not assigned to HMuxes to the SMuxes.

The use of SMuxes in this way also provides high availability during residual mapping failure. §3.7 gives details on how Rubik recovers from a variety of failures.

Summary The benefits of this architecture can only be realized by carefully calculating the DIP placement, and local and residual mappings for individual VIPs subject to a variety of constraints, which we describe next.

3.6 Joint VIP and DIP Assignment

Rubik's objective is to maximize the traffic handled by the HMuxes, while maximizing the traffic handled locally within ToRs. The assignment algorithm determines for each VIP, (1) the location of its DIPs; (2) the number of DIPs in each ToR in the local VIP-to-DIP mapping; and (3) the number of DIPs in the residual mapping, and the HMux assigned to store the mapping.

Rubik needs to calculate this assignment such that the capacity of all resources (switch tables, links, and servers per ToR) is not be exceeded. Also, Rubik needs to ensure that it assigns DIPs in the failure domains (*i.e.,* ToRs) specified by the online services. The placement calculated at a given time may lose effectiveness over time as the VIP traffic changes, and VIPs and DIPs are added and removed. To adapt to such cloud dynamics, Rubik reruns the placement algorithm from time to

```
Input: $V, M, N_v, f_v, S, L, b_{t,v}, C_{t,v}$
Output: $x^D_{t,v}, x^M_{t,v}$
topological_sort(V in DAG)
for $l$ = 1, depth of DAG do
    local_mapping_and_dip_placement(VIPs in DAG_level($l$))
end
residual_mapping_placement()
```

Algorithm 1: Rubik Assignment Algorithm

time. While calculating a new assignment, Rubik has to ensure that the number of machines migrated from the old assignment is under the limit.

The assignment problem is a variant of the bin-packing problem (NP-hard [7]), where the resources are the bins, and the VIPs are the objects. It is further compounded because the VIP traffic exhibits hierarchical dependencies (§3.2.2).

To reduce the complexity, Rubik decomposes the joint assignment problem into two independent modules, (1) *DIP and local mapping placement*, (2) *residual mapping placement*, as shown in Algorithm 1. The first module places the DIPs and local mappings for all the VIPs to maximize the total traffic load-balanced locally on individual ToRs. We calculate DIP and local mapping simultaneously, because the problem of DIP and local mapping placement are intertwined, as the traffic for a VIP is contained within a ToR only if the ToR has (1) enough DIPs to handle the traffic, and (2) enough memory to store the corresponding VIP-to-DIP mapping.

Since the VIP traffic exhibits hierarchical dependencies (§3.2.2), we create a DAG that captures the traffic flow and hence the dependency between the VIPs, and then perform a topological sort on the DAG to divide the VIPs into different levels. We then place the DIPs and local mappings for the VIPs level-by-level (lines 3:6 in Algo. 1). As we place DIPs for VIPs in one level, the sources in the next level become known.

The second module places the residual mappings of all the VIPs to maximize the total traffic handled by HMuxes. The residual mapping placement subproblem

Table 3.2.: Notations used in the algorithm.

Notation	**Explanation**
Input	
S, L, V	Sets of switches, links, and VIPs
M_t	# servers under t-th ToR
T_s	Table capacity of s-th switch
L_e	Link traffic capacity of link e
N_v, f_v	#DIPs and failure-domain for v-th VIP
$b_{t,v}$	Traffic sent to v-th VIP from t-th ToR
$C_{t,v}$	Traffic capacity of server in t-th ToR when assigned to v-th VIP
Variables	
$x^D_{t,v}$	Number of servers (DIPs) in t-th ToR assigned to v-th VIP
$x^M_{t,v}$	Number of table entries in t-th ToR assigned to v-th VIP

remains NP-hard. But, the residual VIP traffic is typically only a small portion of the total traffic and hence we can apply heuristics to solve it without significantly affecting the quality of the overall solution (line 7).

3.6.1 DIP and Local Mapping Placement

The first module places the DIPs and the local mappings of all the VIPs for which the sources are known such that the total VIP traffic load balanced within the ToRs is maximized. We formulate the joint DIP and local mapping placement problem as ILP using notations shown in Table 3.2 as follows.

Input: The input includes (1) the network topology and resource information (capacity of switch tables, links, and servers in the ToRs), (2) for every VIP in current level, the number of DIPs and number of failure domain and traffic, and (3) max. number of DIPs to migrate (δ).

Output/Variables: The output includes the local VIP-to-DIP mappings on individual ToRs, and placement of all the DIPs (including residual DIPs), for all VIPs.

Let $x^D_{t,v}$ denote the number of machines in the t-th ToR assigned as the DIPs for the v-th VIP, and $x^M_{t,v}$ denote the number of machines out of these $x^D_{t,v}$ machines that are used in the local mapping for the VIP, *i.e.,* they will appear in the local VIP-to-DIP mapping of the t-th ToR.

Objective:

maximize Locality $L = \sum_{v \in V} \sum_{t \in T} y^M_{t,v} \cdot b_{t,v}$

where $y^D_{t,v}$ is set if there are any DIPs in the t-th ToR assigned to the v-th VIP, and $y^M_{t,v}$ is set if the t-th ToR switch (HMux) contains local VIP-to-DIP mapping for the v-th VIP. This way, $y^M_{t,v} \cdot b_{t,v}$ denotes if traffic for v-th VIP in t-th ToR is handled locally, and we maximize traffic handled locally across all VIPs and ToRs.

$$y^M_{t,v} = \begin{cases} 1 & x^M_{t,v} \geq 1 \\ 0 & \text{Otherwise} \end{cases} \qquad y^D_{t,v} = \begin{cases} 1 & x^D_{t,v} \geq 1 \\ 0 & \text{Otherwise} \end{cases}$$

Constraints:

(1,2) Switch table size and number of servers not exceeded on every ToR

$$\forall t \in T, \sum_{v \in V} x^M_{t,v} \leq T_t, \quad \sum_{v \in V} x^D_{t,v} \leq M_t$$

(3,4) Specified number of DIPs assigned for every VIP; failure domain constraints

$$\forall v \in V, \sum_{t \in T} x^D_{t,v} = N_v, \quad \sum_{t \in T} y^D_{t,v} \geq f_v$$

(5a, 5b) DIPs are not overloaded (no hot-spots)

$$\forall t \in T, \forall v \in V, y^M_{t,v} \cdot b_{t,v} \leq x^M_{t,v} \cdot C_{t,v}$$

$$\forall v \in V, \sum_{t \in T}(1 - y^M_{t,v}) \cdot b_{t,v} \leq \sum_{t \in T}(x^D_{t,v} - x^M_{t,v}) \cdot C_{t,v}$$

Constraint (5a) ensures the DIPs mapped in the local mapping are not overloaded. Constraint (5b) ensures the DIPs in the residual mapping are not overloaded.

(6) Limiting the number of DIP moves

$$\sum_{v \in V, t \in T} |x^M_{t,v} - x^{M,old}_{t,v}| \leq \delta$$

where $x^{M,old}_{t,v}$ denotes the number of DIPs in the ToR in the previous assignment, and δ is the threshold on the maximum number of DIPs to be moved. We convert constraint (6) into the linear form as:

$$\sum_{v \in V, t \in T} z_{t,v} \leq \delta$$

$$\forall t \in T, \forall v \in V, z_{t,v} \geq x^M_{t,v} - x^{M,old}_{t,v}, z_{t,v} \geq x^{M,old}_{t,v} - x^M_{t,v}$$

(7) ToRs have more DIPs than in local mappings

$$\forall v \in V, t \in T, x^D_{t,v} \geq x^M_{t,v}$$

(8a,8b) Writing $y^M_{t,v}$, $y^D_{t,v}$ in linear form

$$\forall t \in T, \forall v \in V, 0 \leq y^M_{t,v}, y^D_{t,v} \leq 1, y^M_{t,v} \leq x^M_{t,v}, y^D_{t,v} \leq x^D_{t,v}$$

3.6.2 Residual Mapping Placement

The second module places the residual mappings for the VIPs among the switches while maximizing the total VIP traffic load balanced by the residual mapping HMuxes (traffic not handled by local mappings), subject to switch memory and link capacity constraints.

This assignment problem is the same as that in Duet, and we solve it using the same heuristic algorithm as in Duet. Briefly, to assign the VIPs, we first sort the

VIPs in decreasing traffic volume, and attempt to assign them one by one. We define the notion of maximum resource utilization (MRU). MRU represents the maximum utilization across all resources – switches and links. To assign a given VIP, we consider all switches as candidates. We calculate the MRU for each assignment, and pick the one that results in the smallest MRU, breaking ties at random. If the smallest MRU exceeds 100%, *i.e.,* no assignment can accommodate the traffic of the VIP, the algorithm terminates. The remaining VIPs are not assigned to any switch – their traffic will be handled by the SMuxes.

3.7 Failure Recovery

A key requirement of the LB design is to maintain high availability during failure: (1) the traffic to any VIP should not be dropped, (2) existing connections should not be broken. As in Duet, Rubik relies on SMuxes to load balance the traffic during various failures. In addition to storing VIP-to-DIP mapping for all the VIPs, we use the ample memory on individual SMuxes to provide connection affinity by maintaining per-connection state.

Residual mapping HMux failure: Failure of the HMux storing the residual mapping of a VIP only affects the traffic going to that HMux; the traffic handled by other local and residual mappings is unaffected. The routing entries for the VIPs assigned to the failed HMux are removed from all other switches via BGP withdraw messages. After routing convergence, traffic to these VIPs is routed to the SMuxes, which announce all VIPs. Since each SMux stores the same residual DIPs and uses the same hash function as the residual mapping HMux to select a DIP, existing connections are not broken.

Local mapping failure: When a ToR switch fails, all the sources and DIPs for a VIP under it are also disconnected. As a result, the traffic the local mapping was handling also disappears. Further, the rest of the traffic for that VIP continues to be routed to the residual mapping or other local mappings, and are not affected.

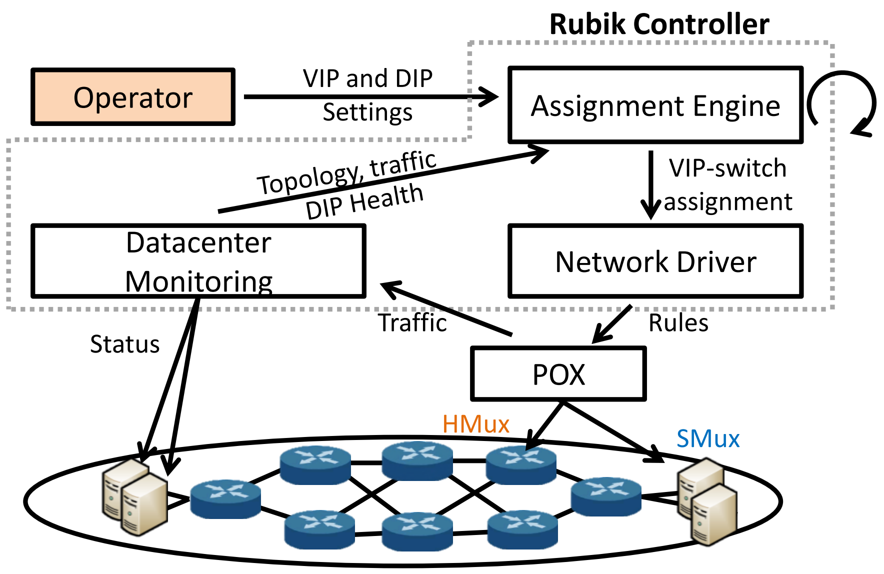


Fig. 3.6.: Rubik implementation.

SMux failure: On an SMux failure, traffic going to that SMux is rerouted to the remaining SMuxes using ECMP. The connections are not broken as all the SMuxes use the same hash function.

DIP failure: Existing connections to the failed DIP would necessarily be terminated. For VIPs whose mapping are assigned to SMuxes, connection to the remaining DIPs are maintained as SMuxes use consistent hashing in DIP selection [2]. For VIPs assigned to HMuxes, the connections are maintained using smart hashing [8].

3.8 Implementation

We briefly describe the implementation of the three building blocks of Rubik, (1) Rubik controller, (2) network driver, (3) HMux and SMux, as shown in Figure 3.6.

Rubik controller: The controller orchestrates all control activities in Rubik. It consists of three key modules: (1) DC monitor, (2) Assignment engine, (3) Network driver. The DC monitor periodically captures the traffic and DIP health information from the DC network and sends it to the assignment engine. The assignment engine calculates the DIP placement, local and residual VIP-to-DIP mappings for all the

VIPs, and pushes these new assignment to the network driver. We use CPLEX [34] to solve the LP (§3.6.1).

Network driver: This module is responsible for maintaining VIP and DIP traffic routing in the LB. Specifically, when the VIP-to-DIP assignment changes, the network driver announces or withdraws routes for the changed VIPs according to BGP.

HMux and SMux: We implement HMuxes and SMuxes using Open vSwitches that split the VIP traffic among its DIPs using ECMP based on the source addresses [35]. We implement smart hashing [8] using OpenFlow rules. The replies from the DIPs directly go to the sources using DSR [2].

Lastly, we use POX to push the rules and poll the traffic statistics. We developed a separate module to monitor the DIP health. The code for all the modules consists of 3.4K LOC in C++ and Python.

3.9 Testbed

Setup: We evaluate Rubik prototype using Open vSwitches and Mininet. Our testbed (Fig. 3.7) consists of 20 switches (HMuxes) in 4 containers connected in a FatTree topology. Each ToR contains an SMux (marked "M") and 2 hosts that can be set as DIPs (marked "S").

Services: We evaluate the performance of Rubik using two services that require load balancing: (1) HTTP web service, (2) Bulk data transfer service. The web service serves static web pages of size 1KB and generates a large number of short-lived TCP flows. The Bulk data transfer service receives a large amount of data using a small number of long-lived TCP flows. All the servers and clients for these services reside in the same DC.

Experiments: Our testbed evaluation shows: (1) Rubik lowers congestion in the network; (2) Rubik achieves high availability during a variety of failures – local mapping, residual mapping, and DIP failure.

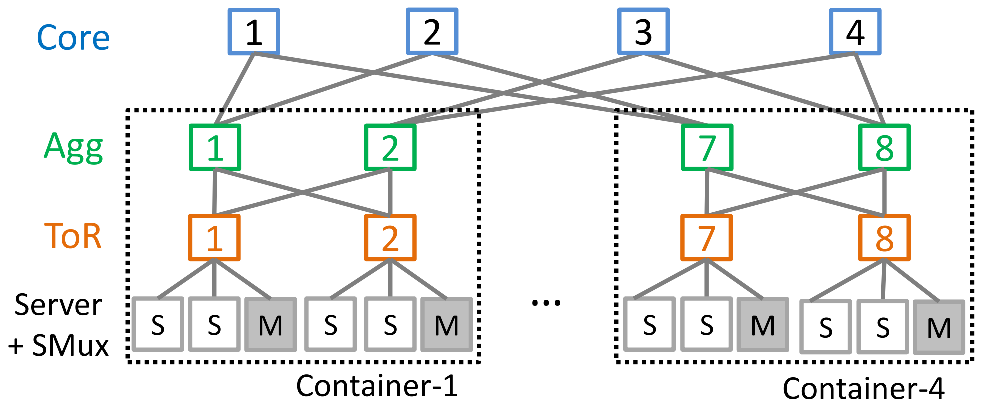


Fig. 3.7.: Our testbed. FatTree with 4 containers connected to 4 Core switches.

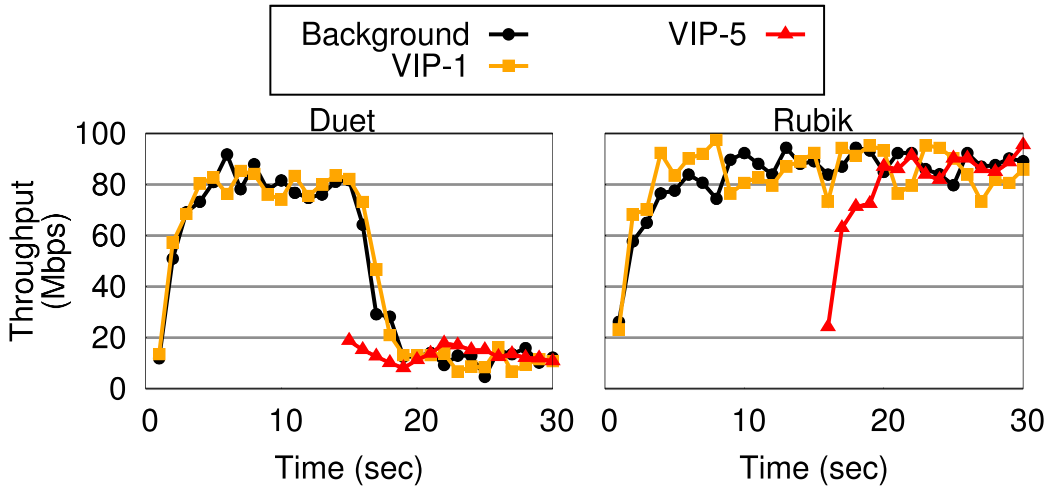


Fig. 3.8.: Rubik reduces congestion.

3.9.1 Reduction in Congestion

First we show that Rubik reduces congestion in the DC network by using local mappings. In this experiment, initially 4 VIPs (each with 1 source and 1 DIP) are assigned to 4 different HMuxes. Additionally, there is background traffic between 2 hosts. Figure 3.8 shows the per-second throughput measured across 2 flows. "VIP-1" denotes the throughput for one of the 4 VIPs added initially. "Background" denotes the throughput for the background flow (not going through the LB). Initially, there is no congestion in the network and as a result all flows experience high throughput. At

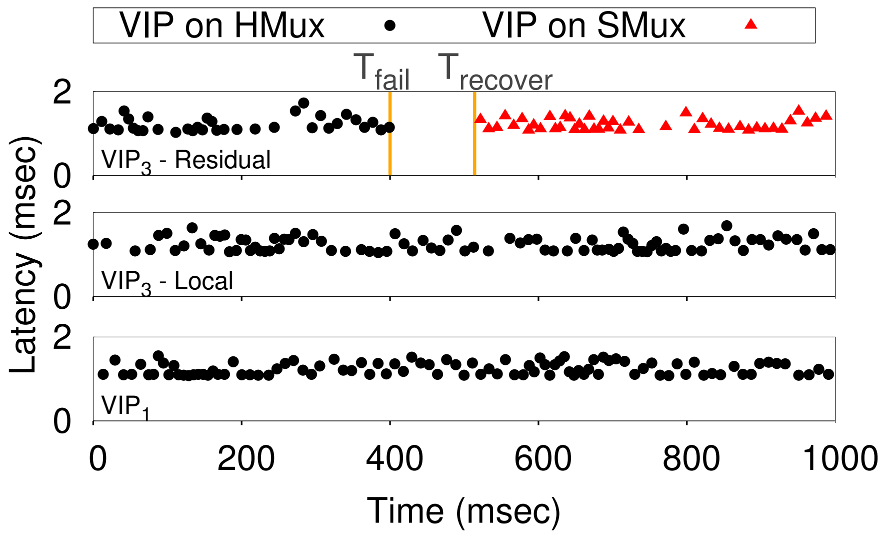


Fig. 3.9.: VIP availability when residual mapping fails.

time 15 sec, we add a new VIP (VIP-5) that has 2 DIPs and 2 sources sending equal volume of traffic, and assign it using Duet. However, assigning the new VIP causes congestion as the new flows compete with the old flows. As a result, the throughput for all the flows drop by almost 5-6x.

We repeat the same experiment with Rubik. At time 15 sec, we assign the VIP-5 using Rubik. Rubik assigns local mappings to handle the VIP-5 traffic. As a result, adding VIP-5 does not cause congestion (no drop in throughput), as shown in Figure 3.8. This experiment shows that by exploiting locality, Rubik reduces the congestion and improves the throughput by 5-6x.

3.9.2 Failure Mitigation

Next we show how Rubik maintains high availability during various failures.

Residual mapping failure: Fig. 3.9 shows the availability of the VIP, measured using ping latency, when its residual mapping fails. In this experiment, we have 3 VIPs (VIP-1, 2, 3) assigned to the data-transfer service. VIP-1 and VIP-2 have one

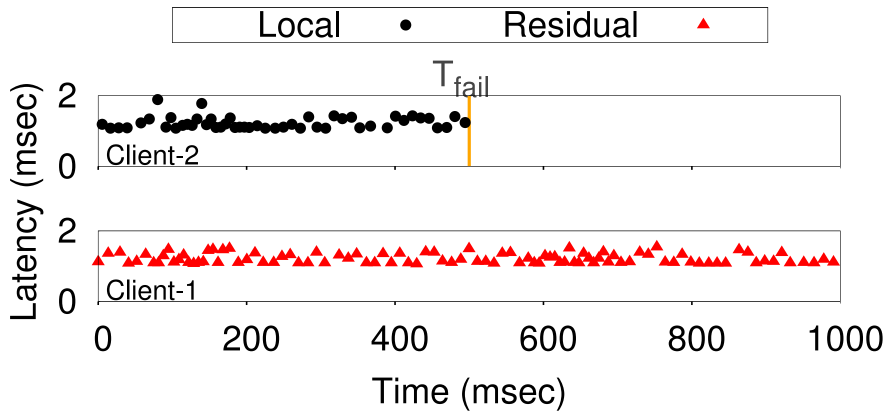


Fig. 3.10.: VIP availability during local mapping failure.

source and one DIP each in different ToRs, and their traffic is handled by residual mappings (no local mapping). VIP-3 has two sources and two DIPs. One source and one DIP are in the same ToR – the local mapping on that ToR handles their traffic. The remaining source and DIP are in two different ToRs, and their traffic is handled by the residual mapping.

At 400 msec, we fail the HMux storing the residual mapping for VIP-3. We make four observations: (1) On HMux failure, VIP-3 traffic handled by it is lost for 114 msec. (2) After 114 msec, VIP-3 is 100% available, *i.e.*, all of the pings are successful again. During this time, the routing converges, and the traffic that used to go to the HMux is rerouted to the SMuxes. (3) The traffic for VIP-3 handled by the local mapping (shown as VIP-3-Local) is not affected – no ping message is dropped. (4) Other VIPs (only VIP-1 is shown) are not affected – their ping messages are not dropped.

This shows that Rubik provides high availability during residual mapping failure.

Local mapping failure: Figure 3.10 shows the impact of local mapping failure on the availability of the VIPs. We use the same setup as before, and fail the HMux where VIP-3's local mapping was assigned. We measure the ping message latency from 2 sources for VIP-3 (denoted as Client-1, 2). The traffic from Client-2 is handled locally,

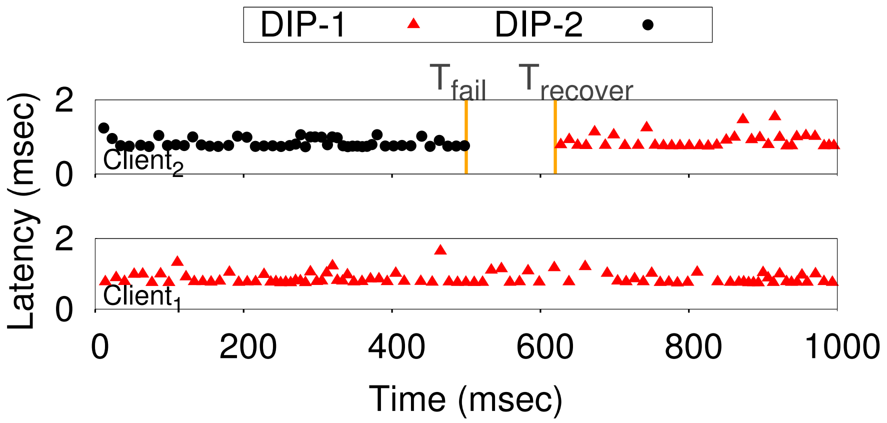


Fig. 3.11.: VIP availability during DIP failure.

whereas Client-1 traffic is handled by the residual mapping. When local mapping fails (at 500 msec), all the sources and DIPs under it disappear. Therefore, ping messages for Client-2 are lost as Client-2 itself is down. Figure 3.10 shows that the traffic from Client-1, which is handled by the residual mapping, is not affected.

DIP failure: Lastly, we evaluate the impact of DIP failure on service availability. In this experiment, we use a single VIP with 2 sources (Client-1, 2) and 2 DIPs (DIP-1, 2), located in different ToRs. Therefore, both DIPs are assigned to the residual mapping. Initially, the traffic from Client-1 is served by DIP-1 and that of Client-2 is served by DIP-2. We fail DIP-2 at 500 msec.

Figure 3.11 shows the latency for the ping messages from Client-1 and Client-2. When DIP-2 fails, the ping messages for Client-2 are lost for about 120 msec. After 120 msec, Client-2 traffic is served by DIP-1. This is because when DIP-2 fails, the residual mapping is adjusted using smart-hashing, *i.e.,* the traffic going to the failed DIP is split across the remaining DIPs. As a result, the traffic going to DIP-2 is now served by DIP-1. It can also be seen that Client-1 traffic is not affected – there is no drop in the ping messages. This shows that a DIP failure does not affect the traffic going to other DIPs, and traffic going to the failed DIP is spread across remaining DIPs.

3.10 Simulation

In this section, we use large-scale simulations of Rubik and Duet to show: (1) Rubik handles a large percentage of traffic in HMuxes as in Duet but incurs significantly lower maximum link utilization (MLU); (2) Rubik reduces the traffic in the core by 3.68x and in the container by 3.47x; (3) Rubik contains 63% of VIP traffic within ToRs; (4) Rubik does not create hotspots.

Network: Our simulated network closely resembles that of a production DC, with a FatTree topology connecting 500K VMs under 1600 ToRs in 40 containers. Each container has 40 ToRs and 4 Agg switches, and the 40 containers are connected with 40 Core switches. The link and switch memory capacity were set with values observed in the production DC.

Workload: We run the experiments using traffic trace collected from the production DC over a 24-hour duration. The trace consists of the number of bytes sent between all sources and all VIPs. Figure 3.12 shows the total traffic per hour fluctuates over the 24-hour period[2].

Comparison: We compare the performance of Duet, Rubik-LO and Rubik. Duet exploits neither locality nor DIP placement. Rubik-LO is a version of Rubik that only exploits locality without moving the DIPs; it assumes DIP placement is fixed and given, and only calculates the local and residual mappings. Rubik exploits both locality and flexibility in moving the DIPs. Rubik performs stage-by-stage VIP-to-DIP mapping assignment following the VIP dependency.

3.10.1 MLU Reduction

We first compare the trade-off between the MLU and fraction of the traffic handled by the HMuxes under the three schemes. Note that all three schemes try to maximize the total traffic handled by HMuxes. The traffic not handled by HMuxes is handled by SMuxes.

[2]Absolute values are omitted for confidentiality.

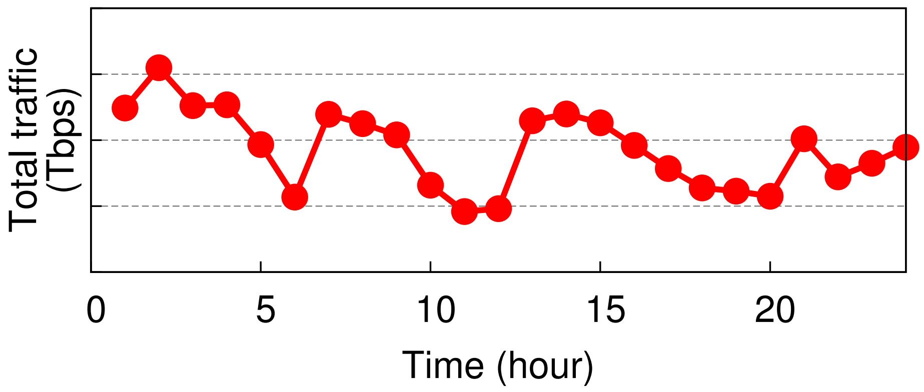


Fig. 3.12.: Total traffic variation over 24 hours.

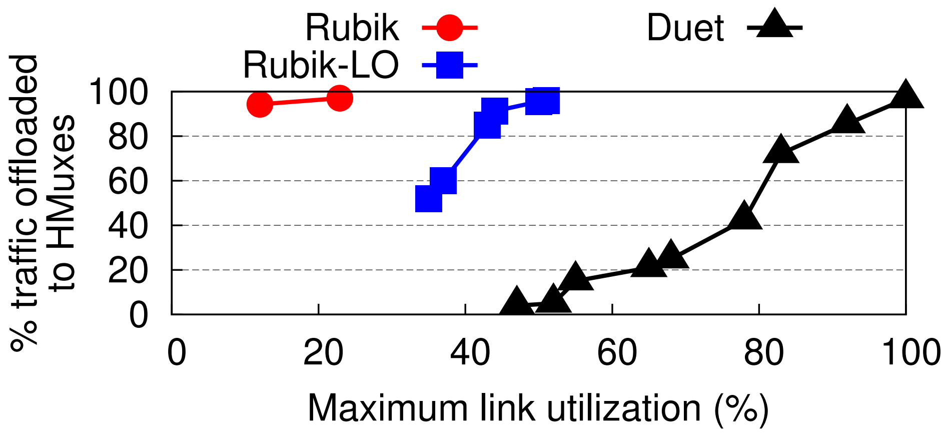


Fig. 3.13.: Traffic handled by HMuxes vs. MLU.

Figure 3.13 shows the fraction of traffic handled by HMuxes under the three schemes. The MLU shown is the total MLU which resulted from load balancing all VIP traffic, handled by HMuxes and by SMuxes. We see that Duet can handle 97% traffic using HMuxes, but incurs a high MLU of 98%. But when MLU is restricted to 47%, Duet can only handle 4% traffic using HMuxes.

In contrast, Rubik-LO handles 97% VIP traffic using HMuxes at MLU of 51%. It handles 52% of VIP traffic using HMuxes at MLU of 35%. This improvement over Duet comes purely from exploiting locality.

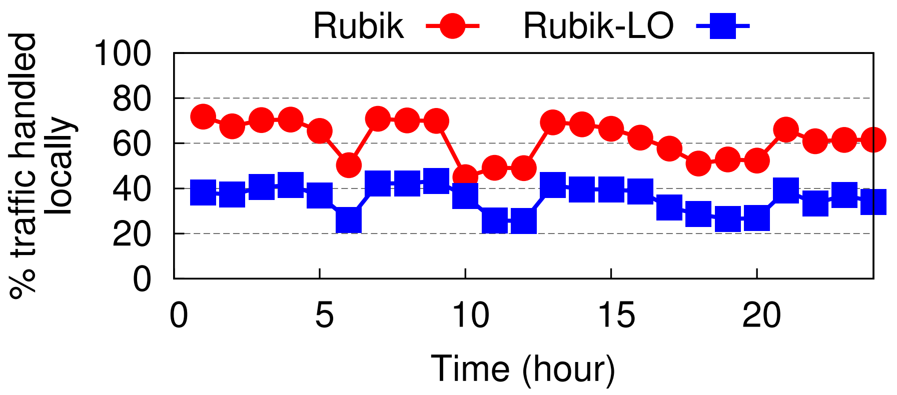


Fig. 3.14.: Traffic handled using local mappings.

Lastly, Rubik significantly outperforms both Duet and Rubik-LO. It handles 97% traffic with a low MLU of 22.9%, a 4.3x reduction from Duet. Also, at a MLU of 12%, Rubik handles 94% traffic using HMuxes.

3.10.2 Traffic Localized

Rubik significantly reduces the MLU by containing significant amount of traffic within individual ToRs. Figure 3.14 shows the fraction of the total traffic contained within ToRs in Rubik and Rubik-LO over the 24-hour period, where these mechanisms calculate new assignment every hour. In Rubik, we limit the machine moves to 1% based on the trade-off detailed in §3.10.5.

We see that Rubik-LO localizes 25.5-43.4% (average 34.8%) of the total traffic within ToRs, and Rubik localizes 46-71.8% (average 63%) of the total traffic within ToRs. Additionally, for the VIPs generating 90% of the total VIP traffic, we find that, the local mappings handle traffic from 37.8-48.6% (average 41.8%) sources, and 50.2-57.7% (average 53%) of the total DIPs are assigned to their local mappings.

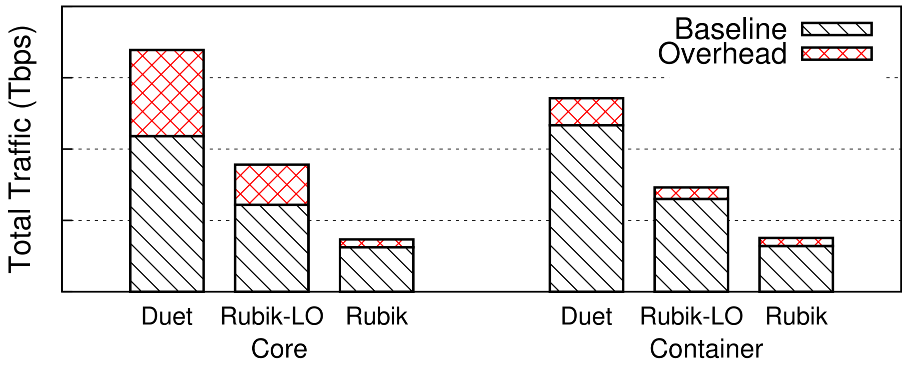


Fig. 3.15.: Total traffic in core and container.

3.10.3 Traffic Reduction

Figure 3.15 shows the total bandwidth usage across all the links caused by the VIP traffic under the three mechanisms. We separately show the total traffic on the core links (between Core and Agg switches) and containers links (between ToR and Agg switches). The total traffic shown is the average over 24 hours. Furthermore, we break down the total traffic into baseline and overhead due to redirection. The baseline traffic shows the amount of traffic generated if the HMuxes were on the direct path between source and DIPs, which would cause no redirection. The remaining traffic is the extra traffic due to the redirection to route traffic to and from HMuxes.

Rubik and Rubik-LO significantly reduce the total traffic in the core network and containers. Compared to Duet, on average Rubik-LO reduces the total traffic by 1.94x and 1.88x, respectively. Rubik reduces the total traffic by 3.68x and 3.47x, respectively.

Secondly, Rubik-LO and Rubik reduce the traffic overhead due to traffic redirection by 2.1x and 10.9x compared to Duet. It should be noted that both Rubik and Rubik-LO cannot eliminate the traffic overhead, because they cannot localize 100% of the VIP traffic. As a result, the traffic not localized is handled by the HMuxes storing residual mappings, which causes traffic detour.

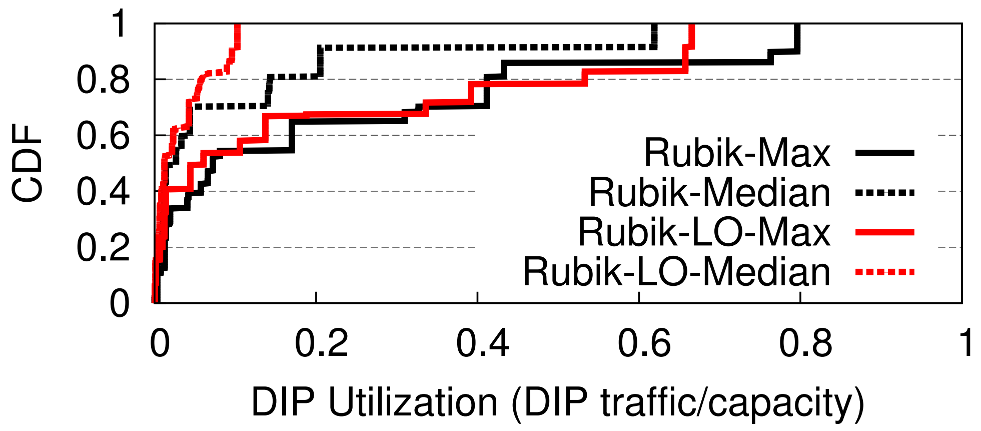


Fig. 3.16.: DIP utilization distribution across VIPs.

3.10.4 DIP Load Balance

To exploit locality, Rubik partitions the DIPs for a VIP into local and residual DIP-sets, which can potentially overload some of the DIPs (hotspots). We calculate the average and peak DIP utilization (DIP traffic/capacity) across all DIPs for every VIP. Figure 3.16 shows the CDF across all VIPs in Rubik and Rubik-LO. It shows that both schemes ensure that the peak utilization for all the DIPs is well under 80%, which is the constraint given to the assignment algorithm. Furthermore, for 80% VIPs, the peak utilization is under 40%. This shows Rubik does not create hotspots.

3.10.5 Impact of Limiting Machine Moves

Lastly, we evaluate the impact of limiting machine moves in Rubik's assignment LP formulation (§3.6.1) on the fraction of traffic localized and MLU. Figure 3.17 shows the two metrics as we reduce the percentage machine moves allowed. Without any restriction, Rubik assignment results in moving 13.7% of the DIPs. When the percentage machine moves is 1%, the fraction of traffic localized decreases by 8.7% whereas the MLU increases by 6.6%, and the execution time to find the solution increases by 2.3x compared to unrestricted machine moves. This shows that most

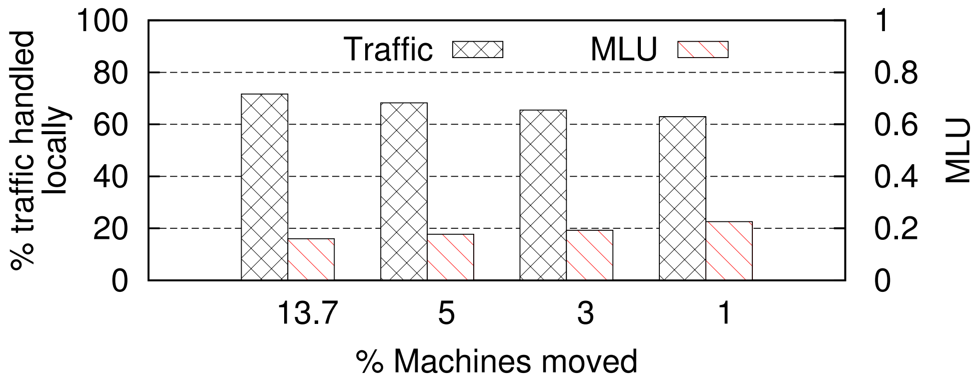


Fig. 3.17.: Impact of machine moves.

of the benefits of Rubik are maintained after restricting the machine moves to just 1%. We therefore used this threshold in all the previous simulations and testbed experiments.

3.11 Related work

To our best knowledge, Rubik is the first LB design that exploits locality and end-point flexibility. Below we review work related to DC LB design which has received much attention in recent years.

LB: Traditional hardware load balancers [14, 15] are expensive and typically only provide 1+1 availability. We have already discussed Duet [4] and Ananta [2] load balancers extensively. Other software-based load balancers [16–19] have also been proposed, but they lack the scalability and availability of Ananta [2]. In contrast to these previous designs, Rubik substantially reduces the DC network bandwidth usage due to traffic indirection while providing low cost, high performance benefits.

OpenFlow based LB: Several recent proposals focus on using OpenFlow switches for load balancing. In [20], the authors present a preliminary LB design using OpenFlow switches. They focus on minimizing the number of wildcard rules. In [36], the authors propose a hybrid hardware-software design and propose algorithms to calculate the weights for splitting the VIP traffic. Plug-n-Serve [37] is another preliminary design that uses OpenFlow switches to load balance web servers deployed in unstructured, enterprise networks. In contrast, Rubik is designed for DC networks and efficiently load balances the traffic by exploiting locality and end-point flexibility.

SDN architecture and middleboxes: Researchers have leveraged the SDN designs in the context of middleboxes for policy enforcement and verification [24, 25], which has a different goal from Rubik. Researchers have also proposed using OpenFlow switches for a variety of other purposes. *e.g.,* DIFANE [22] and vCRIB [23] use switches to cache rules and act as authoritative switches. Again their main focus is quite different from Rubik.

3.12 Summary

Rubik is a new load balancer design that drastically reduces the bandwidth usage while providing low cost, high performance and reliability benefits. Rubik achieves this by exploiting two design principles: (1) locality: it load balances traffic generated in individual ToRs across DIPs present in the same ToRs, (2) end-point flexibility: it places the DIPs closer to the traffic sources. We evaluate Rubik using a prototype implementation and extensive simulations using traces from our production DC. Our evaluation shows together these two principles reduce the bandwidth usage by the load balanced traffic by over 3x compared to the prior art Duet.

4. YODA: HIGHLY AVAILABLE LAYER-7 LOAD BALANCER

In the previous chapters, we saw how Duet and Rubik provide high scalability, low cost, low network bandwidth overhead and more importantly high availability to the L4 load balancers. As we show in this chapter, the principles that enabled high availability in the L4 load balancers cannot be extended to the L7 load balancers. We present Yoda that provides high availability to such L7 load balancers.

4.1 Background and Motivation

We start with a brief background on L7 load balancing in the cloud, describe the current L7 LB solution HAProxy, and point out its limitations.

4.1.1 L7 load balancer

L4 load balancing is a basic load balancing mechanism [2, 20, 36]. Cloud-scale L4 LBs such as Ananta [2] and Duet [4] select the server by calculating the hash over the L4 fields (the 5-tuple consisting of IP type, source and destination IP addresses and port numbers), and use IP-in-IP encapsulation to forward the traffic to the selected server.

As online services grow complex, they raise the need for partitioning website logic and data across different sets of servers or even different DCs. For example, an online service can have one set of servers to handle PHP content and a different set of servers to handle CSS content.

L7 load balancing enables such fine-grained partitioning of traffic handled by the online service. In contrast to an L4 LB, an L7 LB provides content-based switching,

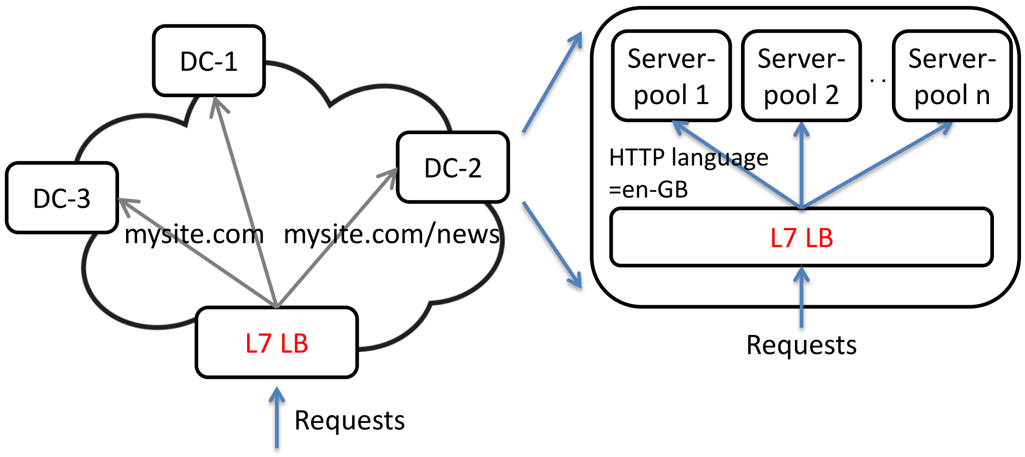


Fig. 4.1.: Typical L7 LB deployment.

where the LB inspects L7 header (HTTP) content in the incoming requests to select the server where the requests are forwarded.

Figure 4.1 shows the typical usage of an L7 LB by online services as well as CDNs. First, at the edge, the L7 LB is used to select the DC when not all the DCs store all the data and applications (*e.g.,* the requests for `mysite.com/news` is served only through DC2). Additionally, within a single DC, an L7 LB is used to load balance the traffic across different server pools (*e.g.,* the requests with language en-GB is served only through serverpool-1 in Figure 4.1).

Key requirements: As an L7 LB touches every packet received by online services, its performance and robustness directly affects the performance and user experience of the online services. For this reason, the LB faces stringent robustness and performance requirements. (1) **Availability:** The LB needs to be online, and maintain connections during failures, (2) **Scalability:** The LB needs to adapt to the dynamic traffic load; (3) **Low latency:** The LB should incur minimal extra delay while load balancing the traffic. (4) **Flexible and expressive interface** so online service operators can easily express policies.

Table 4.1.: Impact of proxy failure on different websites.

website	**nytimes**	**reddit**	**stanford**
impact	page timed-out	page timed-out	page timed-out
website	**vimeo**	**soundcloud**	**email service**
impact	service stopped	service stopped	email failed

4.1.2 Existing L7 LB

Today, major cloud service provides offer highly available and scalable L4 LB services but not an L7 LB service. As a result, tenants are forced to build their own L7 LB or to use third-party L7 solutions [16, 38] in cloud.

Current solutions in public cloud such as HAProxy [16] enable L7 load balancing using a proxy-like mechanism. First, each proxy LB instance establishes a TCP connection with the client and receives the HTTP content. Next, it inspects the HTTP content and selects a server based on the user policies. Once the server is selected, it establishes a TCP connection with the server and simply copies the data between these two connections. To handle higher traffic load, multiple proxy instances are used, and traffic is split among the multiple proxy instances using DNS or the L4 LB service.

4.1.3 Limitations of Existing L7 LB

Problem 1 – Low reliability: The current proxy-based L7 LB mechanism faces a key limitation: **Each L7 LB instance becomes a single point of failure**. Since the proxy operates at L4 and establishes TCP connections with both the client and server, when the LB instance fails, the TCP connection state of the connections with the server and client is lost, the packets from the server and client are dropped. Eventually the two connections are terminated after HTTP timeout, which can be several seconds or even minutes. Prior work has also emphasized that low reliability of middleboxes (*e.g.,* LBs) poses a significant challenge to middlebox vendors and network operators. For example, FTMB [39] notes that *many middlebox vendors have argued against resetting the existing connections when failures happen because of the potential for user-visible disruption to applications.*

Failures are common: [40] shows the hardware middlebox (common in private cloud) failures have low reliability – cumulatively hardware middleboxes (including LBs, Firewalls) account for 43% of all the network device failures. When the middleboxes fail, all the existing connections on those middleboxes fail, and the traffic is lost. [40] shows that 100's of GB of traffic is lost during LB failures.

Similarly, when the L7 LBs are implemented in the software, *e.g.,* HAProxy (common in public cloud), the connections break when the server crashes. The server crashes occur for a variety of reasons including software bugs, hardware failures, maintenance operations, power failures *etc.*, which are common.

To understand the impact such failures can have on user experience, we emulate a proxy failure that breaks a single established connection, and measure its impact on 10 popular websites in terms of page-load time and session reset. Table 4.1 summarizes the results[1]. We see that breaking a single established connection due to the proxy failure either elongates the page-load time by 5 min (default Mozilla Firefox browser HTTP timeout) for popular websites such as nytimes.com or reddit.com, or breaks ongoing sessions for websites like vimeo.com or soundcloud.com. Such high latency and session resets can significantly affect the online service user experience and revenue especially for mobile and/or latency sensitive applications [12, 13, 41].

The low reliability of current L7 LBs leave the websites vulnerable when the LBs crash.

Problem 2 – Management overhead: Using current proxy-based L7 LB solutions (HAProxy), online services have to manage their own LB instances, *i.e.,* tenants have to somehow ensure scalability and availability on their own. Online services

[1]Due to page limit, we show the results for 6 websites.

cannot elastically add/remove LB instances as removing LB instances may break the connections.

4.2 Yoda Key Ideas

In the previous section, we saw that the current proxy-based L7 LBs suffers a fundamental limitation: client sessions are broken in case of LB instance failure. In this thesis, we propose a new multi-tenant cloud L7 LB service called Yoda. Like HAProxy, Yoda comprises of multiple instances running on VMs, does not require administrative access to cloud infrastructure (*e.g.,* switches), and hence can be easily deployed by a third- party. But unlike HAProxy, Yoda provides high availability and scalability, based on two key observations: (1) Each LB instance should not use its actual IP for connecting with the client or server, since upon failure its IP becomes unreachable[2]. (2) Each LB instance should not store the flow state for the connections to the client/server locally, which will be lost upon failure.

These observations motivate the following design principles in Yoda:

Using VIP for client/server connections: In Yoda, each Yoda instance always uses the VIP while making connections to clients and servers. In other words, the clients and the servers always see the other endpoint of their connections as the VIP. This "front-and-back" indirection masks the Yoda instance failures from the clients and servers, and allows for transparent failover of client flows from being handled by one Yoda instance to another. Yoda uses existing L4 LB service in the cloud to redirect VIP traffic to/from the Yoda instances. Specifically, Yoda requires L4 LB to split incoming requests across Yoda instances, and SNAT outgoing responses with the VIP. These requirements are easily met by existing L4 LBs [2, 4]. This decoupling of the L4 and L7 LBs provides modularity and enables the design of both LBs to evolve independently.

[2]Assigning its IP to another Yoda instance potentially in a different part of the DC network can incur a significant delay due to route convergence.

Storing flow state in persistent datastore: Running YODA instances behind VIP is not enough. When one YODA instance fails, the subsequent packets for a flow handled by the failed instance will be rerouted (by the underlying L4 LB) to another YODA instance. For the flow to be maintained, the subsequent packets between the client and the server should be correctly translated, *i.e.,* the sequence numbers, ports, IP addresses should be consistent with before. To do this, the flow state consisting of the client-VIP and VIP-server connections must survive the YODA instance failure and be accessible and reusable by other YODA instances. To achieve this, YODA decouples the flow state from each instance and stores it in a persistent in-memory datastore, called TCPStore, which can be retrieved by any other L7 YODA instances.

The above ideas for achieving availability effectively allows for transparent client flow migration between YODA instances, and thus naturally provides high scalability, where YODA instances can be dynamically added or removed to match the traffic load without affecting existing flows.

Tunneling at L3 for efficiency: A third idea of YODA design is to minimize TCP stack processing and storage operations by tunneling packets. We observe the YODA instance just needs to maintain a TCP connection with the client until it receives the HTTP header, then opens a TCP connection to the selected server and forwards the HTTP request. Since the header has only a few packets, TCP congestion control has not kicked in this phase. After the server receives the HTTP request, the LB instance can simply tunnel the traffic between the server and the client at L3, by properly adjusting the TCP/IP header. Thus in either phase, the LB instance can avoid performing TCP congestion control and leave congestion control to the client and server.

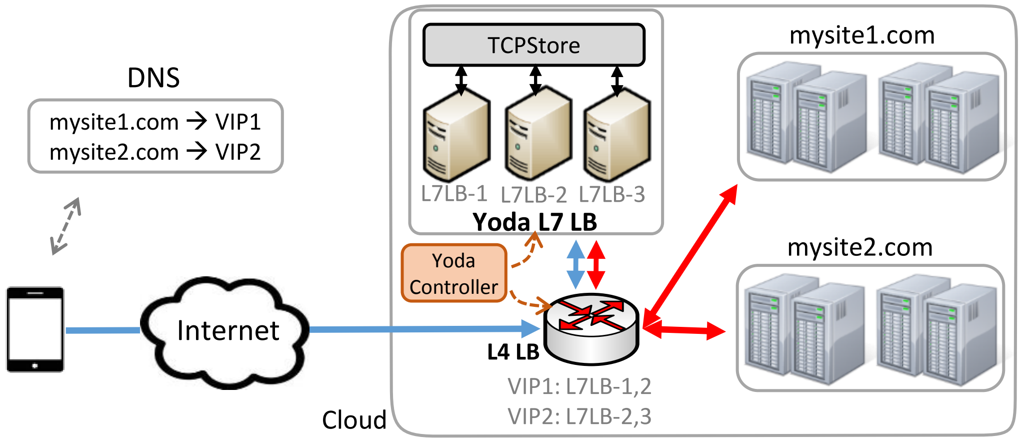


Fig. 4.2.: YODA architecture. Red lines denote L7LB-servers traffic. Blue lines denote L7LB-clients traffic.

4.3 Yoda Design

We now describe the YODA design. To simplify the description, we consider HTTP 1.0, where there is one request/response on one TCP connection. The changes required to support HTTP 1.1 are described in §4.4.2.

4.3.1 Basic operations

Receiving VIP traffic: Figure 4.2 shows the YODA architecture. Each online service (mysite1, 2) is assigned a VIP in its DNS mapping, which clients use to send requests. The requests are first received by the L4 LB in cloud. The L4 LB uses the mapping between the VIP and YODA instances (calculated and set by the YODA controller) to split the incoming VIP traffic among the YODA instances (*e.g.,* VIP1 traffic is split between L7LB1, 2).

YODA **instance**: In a nut shell, an YODA instance performs three functions: (1) *receiving* the HTTP header from the client to select the server; (2) *forwarding* the HTTP request to the server; (3) *forwarding* the data between the client and the server. YODA instances operate in two phases for every flow: (1) connection phase, where

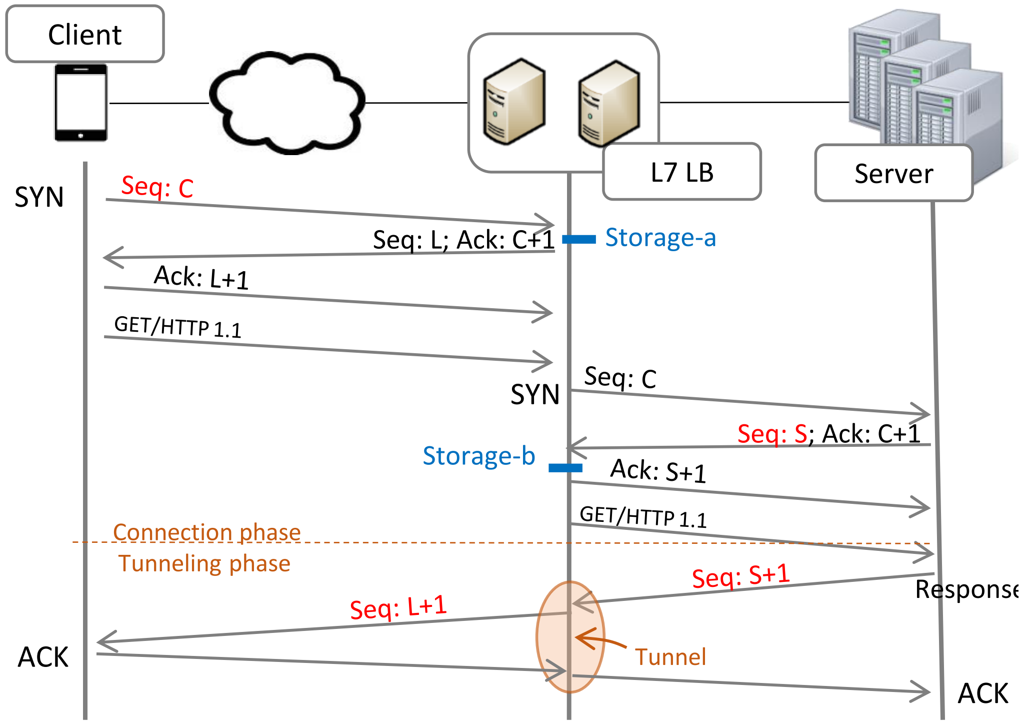


Fig. 4.3.: Connection establishment in YODA. The sequence numbers used while establishing the connections are shown over the arrows. The times when YODA instances store flow state are marked in Blue.

it connects with the client and server, (2) tunneling phase, where it forwards traffic between the client and server connections.

Establishing connection with the client: Figure 4.3 shows how YODA establishes the connection with the client. Upon receiving a SYN packet indicating a new TCP connection, a YODA instance performs two tasks: (1) It stores the TCP header from the client before responding with the SYN-ACK (shown as storage-a in Figure 4.3), so that other YODA instances can retrieve the TCP fields and the sequence numbers on failure of this YODA instance. (2) It sends a SYN-ACK to the client. For a given connection with the client, all YODA instances generate the same SYN-ACK – the sequence number used for the SYN-ACK is generated by hashing the source IP-port tuple. This avoids the need to store SYN-ACK state in TCPStore.

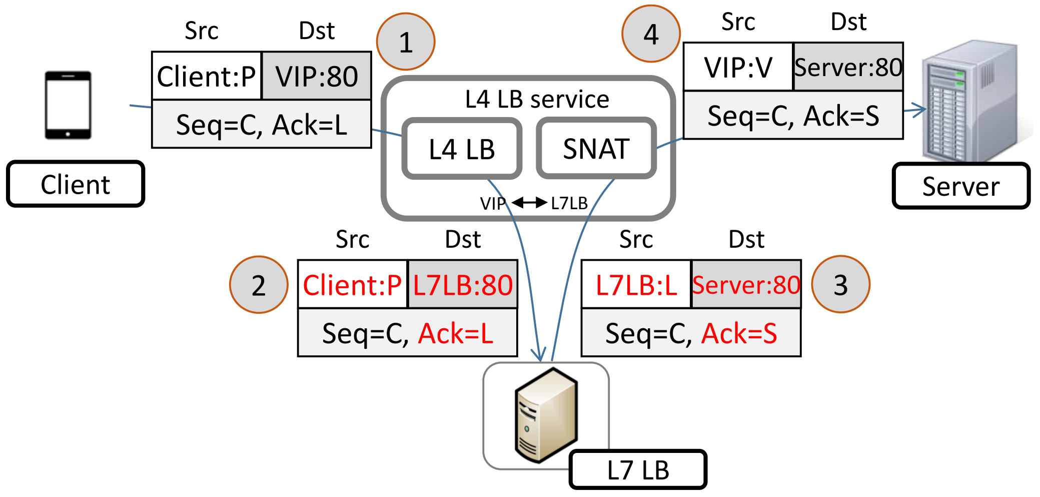


Fig. 4.4.: Front-and-back indirection: YODA leverages the L4 LB to establish TCP connections with both the client and the server: both the client and server see the VIP as the other endpoint of their connections.

Note that the YODA instance cannot select the server yet as it has not received the HTTP header.

When the HTTP header is received, the YODA instance selects the server based on the user policies. We observe there is no need to send any ACK to the client to receive the HTTP header packets as they typically fit in the TCP initial window (but ACK is sent and added to TCPStore if needed), and there is no need to add anything else to TCPStore in this phase.

Establishing connection with the server: YODA instance establishes a connection with the server using the VIP (using the SNAT functionality of the L4 LB), *i.e.*, the source IP in the SYN packet sent to the server is set to the VIP. When it receives the SYN-ACK from the server, it stores the flow state, *i.e.*, IP addresses/ports and the sequence numbers in the TCPStore before sending an ACK (event storage-b in Figure 4.3). This ensures the flow state can be recovered by another YODA instance on current instance failure. YODA instances use the same starting sequence number received from the client to establish connection with the server. This reduces packet processing for the subsequent packets.

Tunneling subsequent packets: Since subsequent packets do not need to be inspected, the YODA instance tunnels all subsequent packets between the client and the server on the two connections at L3 as shown in Fig 4.3. However, the sequence number received in the packet from the server will not match the sequence number used by the YODA instance in its connection with the client, and would require translation to maintain the connection. To do this, the instance keeps the following state locally and also in TCPStore: (1) the starting sequence numbers from the client and the servers C and S, (2) the assigned server.

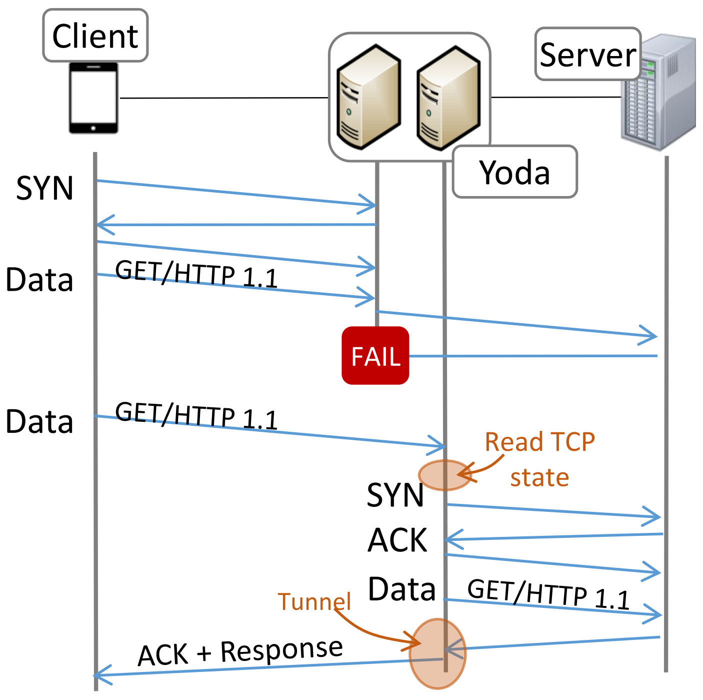
Figure 4.4 shows how the source/destination addresses and the sequence numbers in the packets from the client and server are translated, during both the connection phase and the tunneling phase, so that both the client and server will only see VIP as the other endpoint of their connections.

Terminating connection: The flow state stored at the YODA instance and TCPStore is removed when the instance receives FIN-ACK.

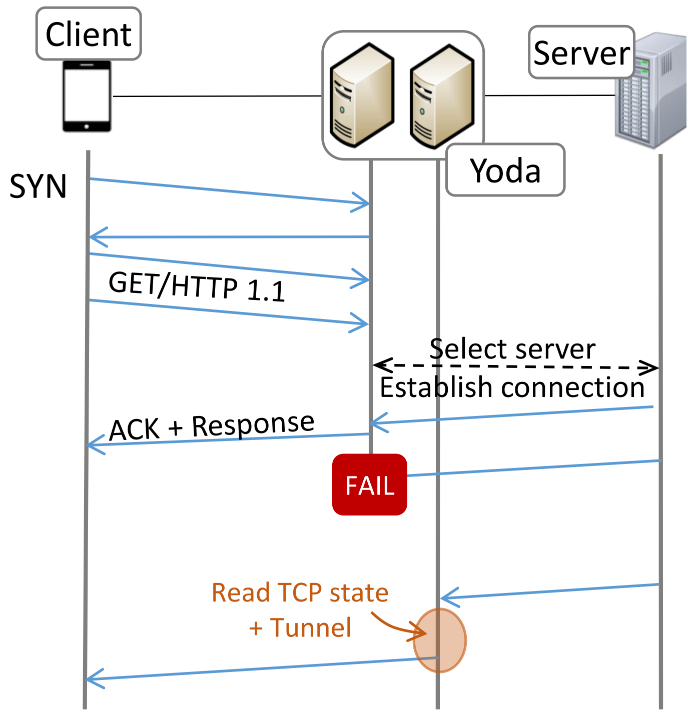
4.3.2 Handling YODA instance failure

The guiding principle to maintain flows during YODA instance failure is that *each instance stores all the packets it ACKes* (*e.g.,* SYN from client and SYN-ACK from server) in the connection phase in TCPStore, as shown in Figure 4.3, so that no state is lost on failures. After failure, other YODA instances can retrieve the flow state from the retransmitted packets and from TCPStore. We now detail on how we use this principle to maintain connections.

Recall the servers and clients do not see the individual YODA instance failure, as YODA instances use the VIP on the connections with the servers and clients. When a YODA instance fails, the YODA controller detects the failure and removes the failed instance from all the mappings at L4 LB. As a result, the underlying L4 LB forwards subsequent packets from the server and the client to one of the remaining instances. We consider failure during the connection phase and tunneling phase separately.



(a) Connection phase



(b) Tunneling phase

Fig. 4.5.: Yoda Failure recovery.

In the connection phase, if the Yoda instance fails before sending SYN-ACK, then when the client resends the SYN packet, the packet is forwarded to one of the remaining Yoda instances and treated as the first packet for a new flow (we observe the SYN timeout to be 3 sec in Ubuntu and packets are rerouted by the L4 LB in less than 600 msec, as shown in §4.6.2).

The more interesting case is when the instance fails any time after inserting the client SYN header into the TCPStore and before forwarding the ACK (from server) back to the client as shown in Figure 4.5(a). As the client does not receive ACK for the HTTP packet(s), it retransmits those packets, which will be forwarded by the L4 LB to one of the remaining Yoda instances. The instance looks up TCPStore to retrieve the flow state (SYN header and sequence number), and detects that it is the first data packet. It then starts a new connection with the server, and continues with the Data/ACK exchange with the client. In this way, the client gets the response without knowing the prior Yoda instance failed.

Lastly, Figure 4.5(b) shows the packet flow during Yoda failure in the tunneling phase. When the Yoda instance fails and another instance receives the re-transmitted packet, it retrieves the flow state from TCPStore, and uses it to adjust the packet header and forwards the packets correctly.

4.3.3 Yoda TCPStore architecture

Yoda decouples the flow state from LB instances and stores them in a separate storage, TCPStore. The key requirements for the storage are: (1) low latency, as any latency added by the storage operations inflates the end-to-end latency, (2) persistence, as the flow state is critical to maintaining the client flows during failure.

Yoda TCPStore is built on top of Memcached, a scale-out key-value store [5] running on commodity VMs. Memcached provides three APIs: `set(key,value)`, `get(key)`, `delete(key)` for accessing key-value pairsflow state. The key drawback of Memcached is that it stores each key-value pair on a single Memecached instance, and does not provide persistence when a Memcached instance fails. We address this limitation by storing each key-value pair on multiple Memcached instances by modifying the Memcached client library which runs on every YODA instance and handles all Memcached operations. §4.5 gives more details on the implementation of TCPStore.

The YODA TCPStore minimizes the latency of the storage operations through a series of optimizations including (1) decentralized server selection, (2) concurrently issuing operations to all replica servers, and (3) using long-lived TCP connections between the Memcached clients and servers. We omit the details here due to page limit.

4.3.4 VIP assignment

So far we have discussed how YODA achieves high availability and scalability – two primary goals of YODA. We now detail the remaining design component on: (1) how YODA selects the server based on the rules, (2) how to assign L7 LB rules for all online services (*i.e.,* VIPs) to the YODA instances to optimize cost, latency, and failure resilience.

Server selection: YODA provides an interface that lets online service operators express their policies as OpenFlow-like rules (§4.4.1) that consist of match, action and priority fields. The YODA instances match the HTTP headers in incoming new connections against these rules in order to select the server for each client flow.

Since designing a new classification algorithm is not the focus of YODA, it uses existing algorithm from HAProxy with an extension to support priority. HAProxy maintains a single table with all the rules chained, and scans all the rules linearly to select the backend server for every incoming new connection. In YODA, we add a new priority field to the rules, and arrange the rules in the decreasing order of the priority. Priority enables YODA to support rich set of policies easily as detailed in §4.4.1.

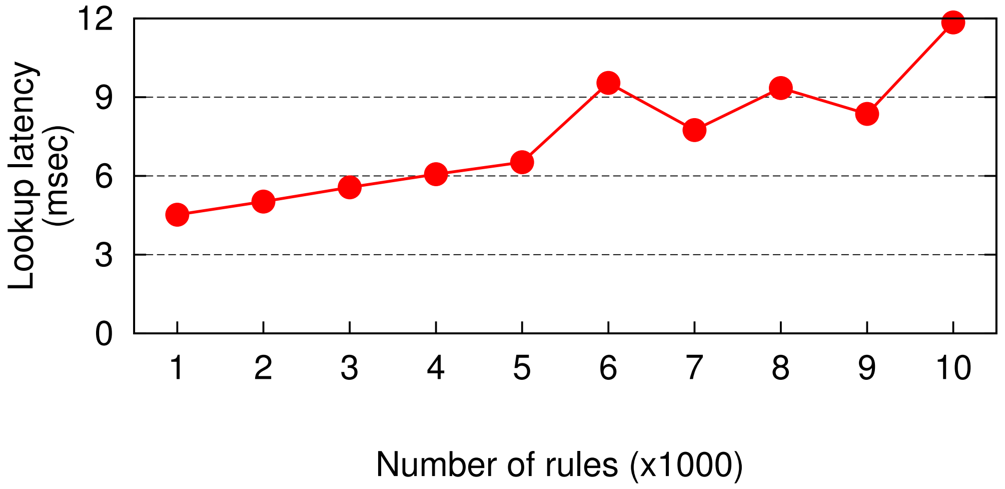


Fig. 4.6.: Look-up latency in HAProxy.

Additionally, YODA maintains a hash table that stores the mapping between the connection (identified using TCP/IP 5-tuple) and the assigned backend server. YODA uses this mapping to forward the subsequent packets on the individual connections.

Next, we detail on how YODA assigns the rules to the YODA instance.

Rule assignment: One simple approach is to assign all the VIPs, and thus all their rules to all YODA instances. We call this scheme all-to-all. Such an approach provides high robustness since any remaining instance can handle traffic for any VIP when some YODA instances fail.

However, it can incur high latency from scanning too many rules for every new flow. We first assess the impact of the number of rules on latency in HAProxy. Figure 4.6 shows that the (P90) latency increases about linearly with the number of rules. The latency for scanning 10K rules is roughly 3x than for 1K rules, which is a significant increase especially for latency sensitive applications.

To address these limitations, YODA uses a many-to-many model for VIP assignment, where a VIP (and its rules) is assigned only to a subset of YODA instances. This model significantly reduces the number of rules on each YODA instance which reduces server selection latency.

Table 4.2.: Notations used in the algorithm.

Notation	**Explanation**
Input	
Y, V	Sets of Yoda instances and VIPs
t_v, r_v	Total traffic and rules for v-th VIP
o_v	Over-subscription for v-th VIP
T_y, R_y	Total traffic and rule capacity of y-th Yoda instance
n_v	# Yoda instances assigned to v-th VIP
f_v	# Yoda failures for v-th VIP to be tolerated
Output Variable (binary)	
$x_{v,y}$	set if v-th VIP is assigned to y-th Yoda instance

The Yoda controller: (1) calculates the VIP assignment (*i.e.,* which VIPs are assigned to which Yoda instances), as the underlying L4 LB splits and forwards traffic at the VIP granularity, all the rules for the individual VIPs are to be installed at the Yoda instances where the VIP is assigned. (2) installs the L7 LB rules at the corresponding Yoda instances, (3) installs the VIP-to-Yoda-instance mapping in the L4 LB, so that VIP traffic is forwarded to the Yoda instances that store the L7 rules for those VIPs.

We formulate the VIP assignment problem in Yoda as an ILP to minimize the Yoda instances for a given latency constraint and level of robustness, as shown in Figure 4.7.

Input: The input to the ILP includes the (1) set of VIPs (V) and Yoda instances (Y), (2) total traffic and rules to each individual VIP (t$_v$, r$_v$), (3) traffic and rule capacity for individual Yoda instances (T$_y$, R$_y$), (4) Number of replicas for individual VIPs (n_v), (5) over-subscription ratio o_v for each VIP (o_v and n_v are specified by online services).

We calculate $f_v = n_v \cdot o_v$, *i.e.,* the number of Yoda instances assigned to v-th VIP whose failure will not overload other Yoda instances.

ILP Variable: $x_{v,y}$

Objective: Minimize $\sum_{y \in Y} y_y$, $y_y = 1$ if $\sum_{v \in V} x_{v,y} \geq 1$, else 0

Constraints:

Traffic capacity: $\forall y \in Y, \sum_{v \in V} x_{v,y} \cdot \frac{t_v}{n_v - f_v} \leq T_y$ (1)

Rule capacity: $\forall y \in Y, \sum_{v \in V} x_{v,y} \cdot r_v \leq R_y$ (2)

Number of replica: $\forall v \in V, \sum_{y \in Y} x_{v,y} = n_v$ (3)

Transient traffic: $\forall y \in Y, \sum_{v \in V} z_{v,y} \cdot t_v \leq T_y$ (4)

where $z_{v,y} = max(\frac{x_{v,y}}{n_v}, \frac{x^{old}_{v,y}}{n^{old}_v})$ (5)

Connections migrated: $\sum_{v \in V, y \in Y} m_{v,y} \cdot C^{old}_{v,y} \leq \delta$ (6)

where $m_{v,y} = max(x^{prev}_{v,y} - x_{v,y}, 0)$ (7)

Expressing y_y: $\forall y \in Y, 1 \geq y_y \geq \frac{\sum_{v \in V} x_{v,y}}{N}$ (8)

Expressing $z_{v,y}$: $z_{v,y} \geq \frac{x_{v,y}}{n_v}$ and $z_{v,y} \geq \frac{x^{old}_{v,y}}{n^{old}_v}$ (9)

Expressing $m_{v,y}$: $m_{v,y} \geq 0; m_{v,y} \geq x^{prev}_{v,y} - x_{v,y}$ (10)

Fig. 4.7.: ILP formulation of VIP assignment.

Output: YODA instances assigned to each individual VIP – $x_{v,y}$ is set if v-th VIP is assigned to y-th YODA instance.

Objective: Minimize the total number of YODA instances, which lowers the YODA operating cost.

Constraints:

- **Traffic capacity:** Each YODA instance has enough capacity to handle traffic even after f_v failures (Eq. 1 in Figure 4.7) for all the VIPs assigned to it. $\frac{t_v}{n_v - f_v}$ denotes the traffic for v-th VIP after f_v failures.
- **Rule capacity:** Each server has enough memory to store the rules for VIPs assigned to it (Eq. 2).

- **Number of YODA instances:** Each VIP is assigned to n_v YODA instances (Eq. 3).
- **VIP migration:** Transient traffic and requests do not overwhelm YODA instances and TCPStore. We detail this constraint in the next section.

4.3.5 Updating VIP assignment

The VIP assignment calculated at one point of time will not be always optimal, due to traffic dynamics, YODA instance failure and recovery, VIP addition/deletion, and changes in the user policies (different number of rules). YODA adapts to such dynamics by adding/removing YODA instances and re-computing VIP assignment from time-to-time. After computing a new VIP-to-YODA instance assignment, the YODA controller simply changes the VIP-to-YODA-instance mapping at the L4 LB, and the L4 LB will start forwarding packets based on the new mapping.

This VIP update mechanism raises two new algorithmic challenges due to transient flow migration that can potentially overload YODA instances and/or TCPStore.

Limiting YODA instance load: First, the VIP-to-YODA-instance mapping has to be changed on multiple L4 LB instances, which is not atomic [2]. As a result, during transition a YODA instance may receive some fraction of the traffic based on the new mapping (from the L4 LB instances that got updated), and some on the old mapping (from the L4 LB instances that are yet to be updated). This can potentially overload YODA instances. We address this challenge by re-purposing the capacity on the YODA instances reserved for failures to absorb the transient traffic. We add a new constraint where the transient traffic – sum of the max traffic for individual VIPs under the old or the new assignment is under the capacity (Eq. 4,5 in Figure 4.7).

Limiting TCPStore load: Secondly, VIP assignment change can potentially cause many connections to migrate if the YODA instances that were assigned to one VIP are removed in the next assignment, which can overload the TCPStore. We address this challenge by adding a constraint δ on the number of connections allowed

Table 4.3.: YODA Interface.

	Name	Priority	Match	action
1	r-jpg2	3	url=*.jpg	split={D_2=0.5, D_3=0.5}
2	r-css1	3	url=*.css	split={D_1=1}
3	r-css1	2	url=*.css	split={D_3=0.5, D_4=0.5}
4	r-cookie	0	cookie=*	table={cookie-table}

to migrate, determined by the throughput of the TCPStore (Eq. 6,7 in Figure 4.7). $C^{old}_{v,y}$ denotes the number of connections for the v-th VIP handled by y-th YODA instance, and $m_{v,y}$ denotes whether v-th VIP is removed from y-th YODA instance.

We describe how we collect the inputs to the assignment component in §4.5, and evaluate the assignment computation time in §4.7.

4.4 L7 LB Features

We now detail how YODA implements L7 LB features.

4.4.1 Interface

The primary goal of YODA interface design is to enable online service operators to easily express their policies on how to split the traffic. From studying use cases, we find that an OpenFlow-like interface provided by HAProxy is simple yet powerful. HAProxy lets operators declare the *rules* consisting of the equivalent of *match* and *action* fields. In YODA, we reuse the HAProxy interface with the addition of *priority*. Priority helps to reduce the number of rules when expressing load balancing policies (see primary-backup policy below). Below we highlight how this interface can be used to set some of the most common policies.

Weighted-split: In the simplest case, operators can specify the weights on how to split the traffic (rule-1 in Table. 4.3).

Primary-backup: In many cases, operators deploy services in a primary-backup model, and want to prioritize traffic to a primary server until it fails/overloads. This can be easily achieved using priority functionality in YODA. Operators can specify the same match condition with two actions with different priorities – a higher priority action specifies weights for primary server(s), and a lower priority rule specifies weights for the remaining servers (rule-2,3).

Sticky-sessions: Operators may want traffic from the same user/session to go to the same server handling that session, by matching on HTTP cookies (rule-4).

Least loaded server: In another common case, operators simply want to forward requests to the least loaded server. This can be done by setting the weights to (-1) for all servers.

4.4.2 Practical issues

We now describe how YODA handles important practical L7 LB issues including configuration changes.

Adapting to user policy change: Online services can change user policies (rules) dynamically during upgrades, new web design, or to add/remove backend servers. YODA's periodic VIP assignment calculation accounts for the number of rules. YODA then simply updates the new rules on the YODA instances where the VIP is assigned. This change does not break existing connections, as YODA instances only apply new policies to new connections. Packets on existing connections continue to be forwarded to their prior assigned server even during soft server removal so that the connections do not break. If the operator removes the server immediately, then it is treated as failure detailed next.

Backend server failure: The YODA monitoring component periodically pings the servers. When a server fails, its connections with YODA instances are terminated[3].

[3]Although not implemented currently, YODA instances can initiate connections with the new server and relay the same request to the new server, transparently to the client.

VIP addition and removal: When new services, *i.e.,* VIPs, are added, YODA first runs the VIP assignment algorithm described in §4.3.4 to calculate the set of YODA instances assigned to the VIP. Next, we add the rules corresponding to the new VIP to the YODA instances where it is assigned. Lastly, we assign the mapping between the VIP and the YODA instances to the L4 LB. The sequence of VIP removal is in reverse of VIP addition.

HTTP 1.1 and HTTP 2.0: In HTTP/1.1, a single TCP connection can be reused for multiple requests, which may match different rules and hence need to be forwarded to different backend servers. YODA supports this by having YODA instances inspect the packets from the clients for the HTTP content. If the server selected for the new request is different than the current one, it closes the old connection and initiates a connection with the new server, and also changes the mapping in TCPStore.

Pipelining requests in HTTP/1.1 poses an additional challenge that YODA instances have to ensure that the responses are sent *in-order*, *i.e.,* in the same order as the received requests, even during YODA failures. YODA achieves this by storing the order in which the requests were received in TCPStore, so that after failure, another instance can gather and forward the responses.

In addition to pipelining requests in HTTP/1.1, HTTP/2.0 and SPDY propose out-of-order delivery of responses to avoid head-of-line blocking [42]. YODA can easily support out-of-order responses from a server by correctly translating the TCP sequence numbers.

SSL support [43]: YODA supports SSL by sending the security certificates (set by the operators) to the clients. On failure during certificate transfer, another YODA instance resends the entire certificate (TCP buffer at the client will remove duplicate packets). The rest of the function remains the same as detailed in §4.2, except the YODA instance uses the security certificate for SSL termination to decrypt the request to get HTTP content and select the server.

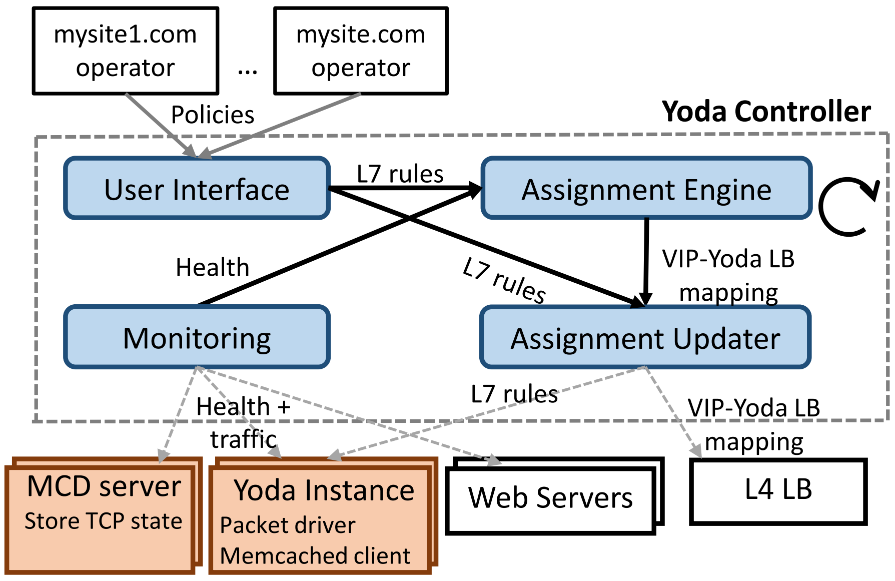


Fig. 4.8.: Components in YODA implementation (shown in colored boxes).

Sending the same request to multiple servers: Many websites send the same client request to multiple backend servers to reduce latency and handle server failures, and send the first server response back to the client. Although it is not currently supported, YODA can easily support this by establishing connections with multiple servers, and tunneling the first response it receives from the servers. It makes a mark to drop packets of later responses from other servers.

4.5 Implementation

We have implemented YODA on Linux completely at the user-level, *i.e.*, it requires no changes to the existing kernel, and therefore it can be deployed in the public cloud as a service by a third party. We now describe the key components of YODA: (1) YODA instances, (2) TCPStore, and (3) YODA controller, as shown in Figure 4.8. All components are written in C or python with the total 5k+ lines of code.

YODA **instances:** Each YODA instance runs in a VM and intercepts the packets forwarded from the L4 LB using `nfqueue` and `iptables` [44]. Each instance runs packet driver and Memcached client. The packet driver performs key functions of the YODA instance. It (1) establishes connections with clients/servers by sending and receiving raw packets, (2) tunnels the packets on existing connections by changing the TCP fields using `nfqueue`, (3) stores and retrieves the flow state in/from TCPStore whenever necessary. The packet driver runs in user-space.

Inside a YODA instance, the packet driver creates K queues, one for each of the K cores of the VM, and starts a multi-threaded module to listen on each queue. The packets are forwarded to `nfqueue` using `iptables` such that the packets on the same TCP connection are forwarded to the same queue. This mechanism ensures that the traffic is evenly spread across all the CPU cores, and packets on the same connection are handled in the same order they arrive. We did not use kernel-space SNAT/DNAT options as they do not provide fine-grained control over the TCP fields.

Lastly, each YODA instance keeps track of the traffic for individual VIPs that YODA controller reads periodically.

TCPStore: We run unmodified Memcached on multiple VMs that store the flow state. We modified the Memcached client library to store the same key-value pair on multiple (K) Memcached servers for persistence. For any TCPStore operation, the Memcached client first determines the K servers among the total N servers using K different hash functions, and consistent hashing.

When a Memcached server fails, we do not replicate its key-value pairs mainly because flows finish quicker than the replication latency.

Controller: The controller is at the heart of YODA architecture. It has four components: (1) User interface: It converts the user policies expressed using the YODA interface into the rules and sends them to the YODA instances. (2) Assignment engine: It calculates the assignment between the VIP and YODA instances using the ILP detailed in §4.3.4 implemented using CPlex [34], (3) Assignment updater: It takes the VIP assignment from the assignment engine and changes the mapping at the L4 LB. (4) Monitor: It gathers health information by pinging the YODA instances, Memcached servers, and backend servers every 600ms, and hence detects failure with at most 600ms delay. It also gets traffic statistics from the YODA instances. All components communicate with each other using RESTful APIs.

4.6 Evaluation

Our testbed experiments evaluate the key ideas and objectives of YODA design and show that: (1) Decoupling flow state from the L7 LB (or middleboxes in general) and maintaining them in a persistent storage is feasible – it incurs insignificant latency to the flows being balanced and the cost of running the persistent storage is low; (2) YODA provides high availability and scalability – it does not break flows during LB failures, addition, removal, and user policy updates.

Setup: Our testbed consists of 60 VMs in Windows Azure: 10 act as YODA instances, 10 as Memcached servers, 30 act as the backend servers, and 10 run a version of Ananta L4 LB that forwards packets to remaining YODA instances upon a YODA instance failure. The clients (generating requests) are located on a university campus.

The backend servers are split across 4 online services – each online service emulates a university website storing faculty and student webpages and embedded objects, which are collected from an actual university website. In total we collected 10K+ objects with sizes 1K-442KB (median 46KB). Each web-request fetches an HTML page and its embedded objects[4].

Each YODA and TCPStore instance runs on a separate VM with 8-core CPU and 14GB RAM, and the backend servers run on dual-core VMs with 3.5GB RAM, running the Apache/2.2.3 HTTP server. Each client generates the request workload using either a Python client that emulates web-browser or the `Apache benchmark` tool. All YODA instances, TCPStore servers, backend servers, and clients run Ubuntu 12.04.

[4]We modified the webpages to change the URL for the embedded objects accordingly.

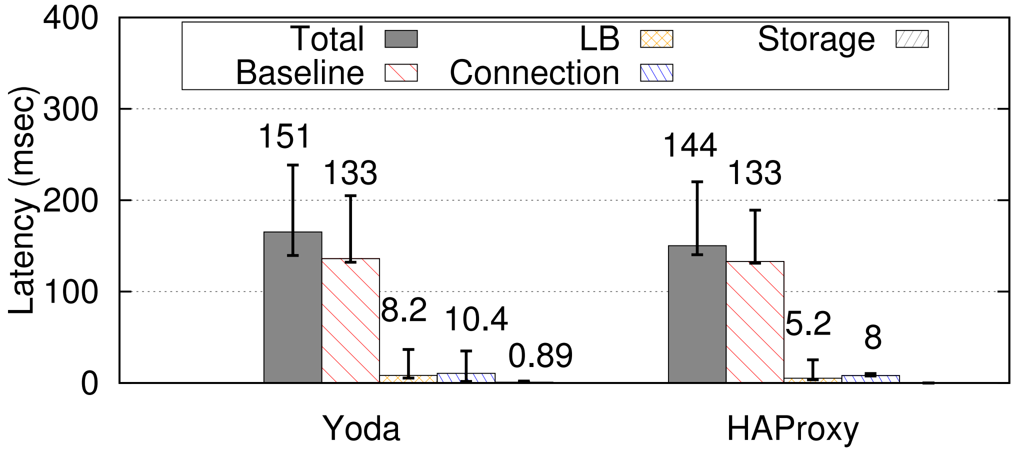


Fig. 4.9.: YODA latency breakdown.

4.6.1 Feasibility of decoupling flow state

We first evaluate the feasibility of decoupling flow state from the YODA instances by measuring its latency overhead to the flows, CPU overhead to the YODA instances, and the scalability of TCPStore.

Latency overhead: To measure the latency overhead, we configure clients to send requests at 50K req/sec for small objects (responses are of size 10KB). Using smaller objects stresses the mechanisms for decoupling flow state since they incur higher load on connection establishment and more frequent operations to TCPStore than larger objects.

To put the latency overhead in perspective, we break down the end-to-end latency (request completion time) into: (1) Baseline: the end-to-end latency when not using any load balancer, which includes the latency incurred by the Internet and backend server processing, (2) Connection: the time an LB instance takes to establish the TCP connection with the backend. (3) Storage: the latency incurred while inserting the flow information into TCPStore, unique to YODA. (4) LB: The remaining latency incurred by the LB instances while processing the packets.

Figure 4.9 shows the breakdown of median latency for YODA and HAProxy, of all flows in the experiment. We make the following observations. (1) The end-to-

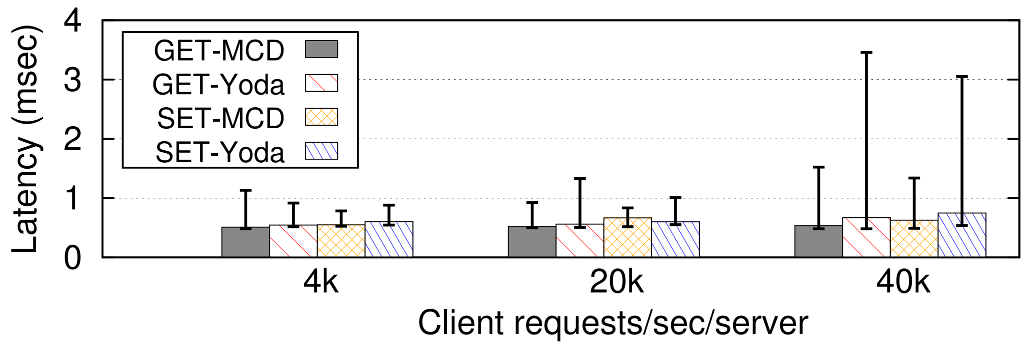


Fig. 4.10.: YODA TCPStore latency.

end latency in YODA and HAProxy are 151 msec and 144 msec, respectively, out of which the baseline latency is 133 msec. (2) The extra latency to store flow states in TCPStore in YODA is only 0.89 msec. (3) YODA takes slightly longer to establish connections – 10.4 msec compared to 8 msec in HAProxy, and to process packets – 8.2 msec compared to 5.23 msec in HAProxy. This is because currently YODA is written in Python and also as it copies packets between user and kernel space while HAProxy performs TCP splicing in the kernel; an in-kernel implementation of YODA will achieve the same latency as HAProxy.

CPU overhead: Next, we measure the CPU utilization on the LB instances. In YODA, the CPU saturates (100% utilization) for 12K client req/sec for small requests[5], and hits 80% for 90K packets/sec request rate for larger requests (flow size 2MB). Under HAProxy, the CPU utilization is 46% and 34% for the two cases. The close to 2x higher CPU utilization in YODA is again due to copying packets between the kernel and user space; we isolated the Memcached client calls and found them to consume negligible CPU load. We fully expect an in-kernel implementation of YODA to reduce the CPU utilization to be similar to that of HAProxy.

TCPStore performance: We first measure the latency overhead of supporting persistence in TCPStore. We measure the latency of `get, set and delete` operations while issuing 40K, 200K, and 400K ops/sec across 10 Memcached servers for

[5] [16] noted higher throughput due to high-end physical server.

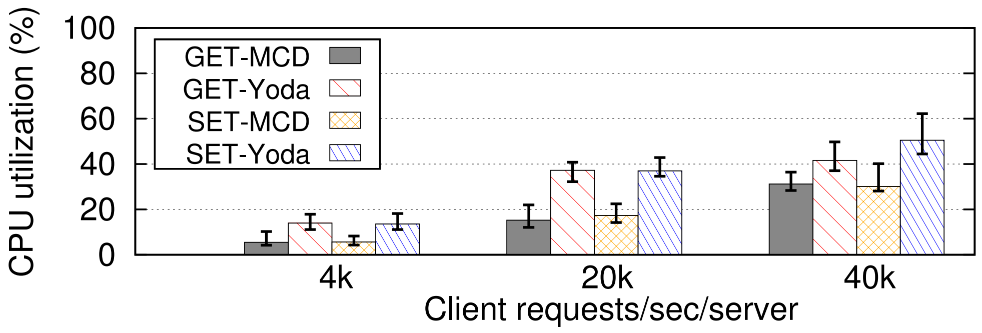


Fig. 4.11.: YODA TCPStore CPU utilization.

60 seconds. Figure 4.10 compares the latency for the default Memcached which does not store replicas and the Memcached changed in YODA to provide 2 replicas for persistence. Note the x-axis denotes the number of client requests per Memcached server. Recall for a single client request, a YODA instance issues two `set` operations to store the flow state (§4.2). We make two observations: (1) even at 40K client req/sec/server, the median operation latency under the default Memcached is only 0.75 msec, which is rather insignificant compared to the median end-to-end latency of 151 msec. (2) The overhead of adding persistence to the default Memcached is very small: for 40K req/sec/sever, the overhead for the three operations is less than 24% (0.18 msec). The low latency overhead benefited from sending the operation to two replica servers in parallel.

Figure 4.11 shows the CPU utilization of Memcached (default and in YODA). As expected, the persistence feature of TCPStore, which issues each operation to two servers, doubles the average CPU utilization.

Our evaluation shows that a single Memcached server can handle 80K client req/sec (at 90% CPU utilization), while each YODA instance can handle 12K client req/sec. This suggests decoupling flow state to support high availability in a scale-out L7 LB design is practical: we just need 1 TCPStore instance deployed for every 6.6 YODA instances.

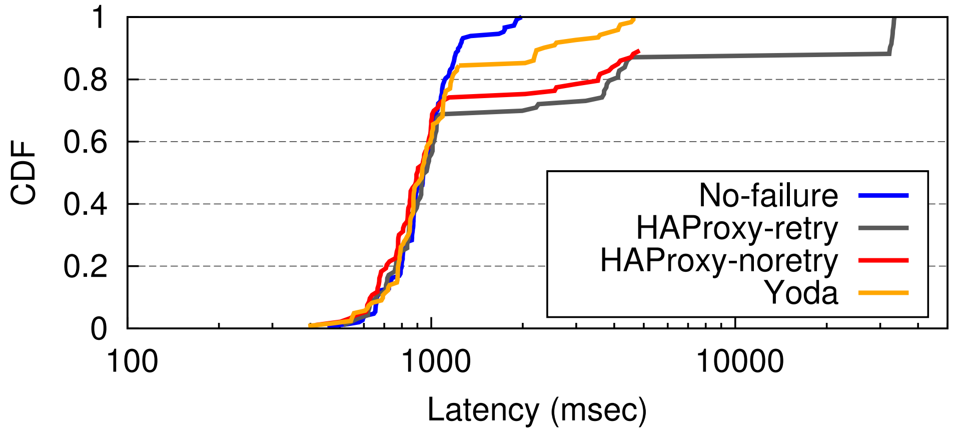
4.6.2 Failure recovery

One of the most important benefits of YODA is it maintains client flows during YODA instance failures. In this experiment, we start with 10 YODA instances and fail 2 of them simultaneously. Without failure, a webpage and all its objects are downloaded in 840 msec (median case). The clients emulate browser behavior with HTTP timeout set to 30 seconds, which is the least among the popular web browsers we tested (*e.g.,* Android Chrome has 60 sec, and C# HttpWebRequest library has 100 sec [45, 46]).

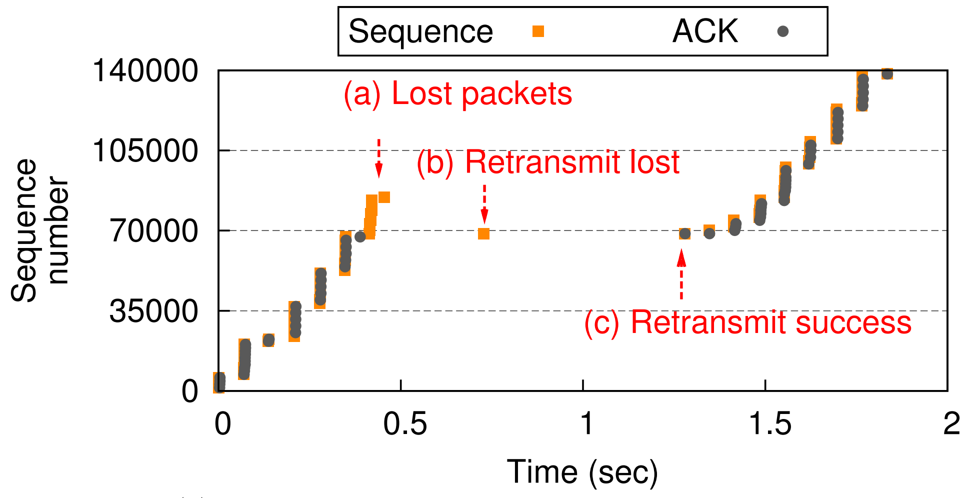
The clients send requests using 20 processes each. Each process waits for the completion/timeout of the previous request before issuing a new one. We repeat this experiment in four scenarios: (1) HAProxy without browser retry (HAProxy-noretry), (2) HAProxy with browser retry=1 (HAProxy-retry), (3) YODA without browser retry (denoted as YODA-noretry), (4) YODA with browser retry=1 (not shown as there was never any retry made).

Figure 4.12 shows the CDF of the end-to-end latency for requests. We make three observations: (1) HAProxy-noretry broke 24% of the flows on failure, whereas YODA and HAProxy-retry did not break any flow, (2) YODA increased the end-to-end latency by 0.6 to 3 seconds, and (3) HAProxy-retry increased the latency beyond 30 sec.

To understand how YODA maintained the flow during failure, we plot in Figure 4.12(b) the `tcpdump` output collected at the backend server for a flow that experienced the YODA instance failure. Upon the failure, (1) the packets from the server and client going through the failed instance are dropped during failover (point a in Figure 4.12(b)). (2) The server retransmits the packet first at 300 msec (point b), which the L4 LB sends to the failed LB, as the mapping at the L4 LB is yet to be updated. (4) The server retransmits the packet at 600 msec (point c). At this time, since the YODA monitor has detected the instance failure and updated the mapping at the L4 LB, the packet is sent to one of the online YODA instances. (5) That in-



(a) Latency



(b) Sequence number for a connection under YODA.

Fig. 4.12.: YODA Failure recovery.

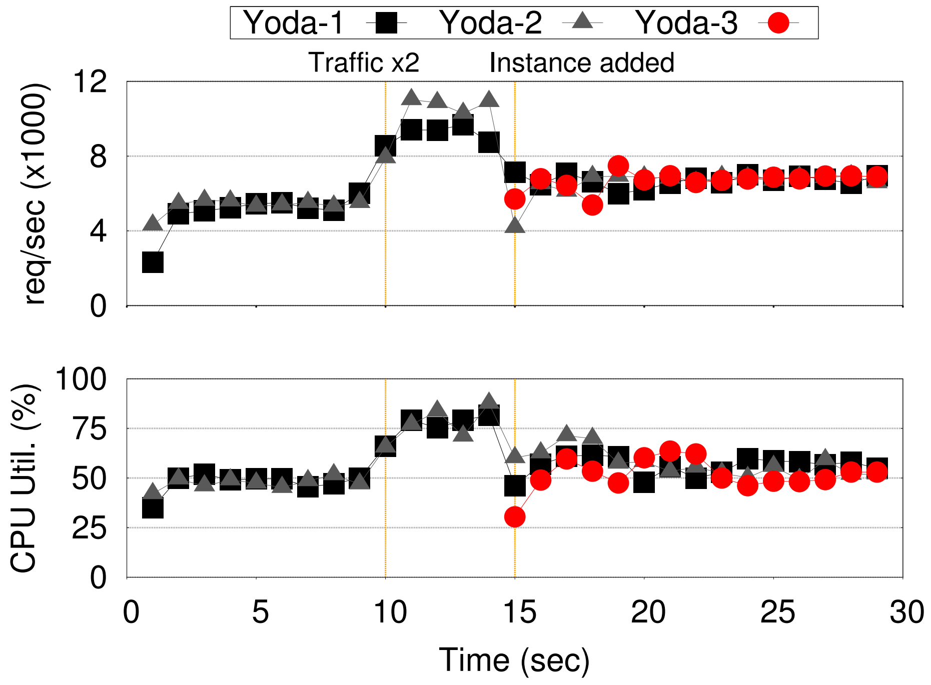


Fig. 4.13.: YODA scalability.

stance retrieves the flow state from TCPStore and forwards the packet to the client, which sustains the connection with the new YODA instance, and subsequent packets are forwarded normally. Note that the client did not timeout, and no HTTP request was resent.

In contrast, when an HAProxy instance fails, all the packets from the servers and ACKs from the clients going through the failed instance are dropped, breaking the connections with the server and the client. After the HTTP timeout of 30 sec, if the client resends the HTTP request (HAProxy-retry), the request is forwarded to one of the live HAProxy instances as the L4 LB is updated; those objects are successfully retrieved but with a delay of 30 seconds.

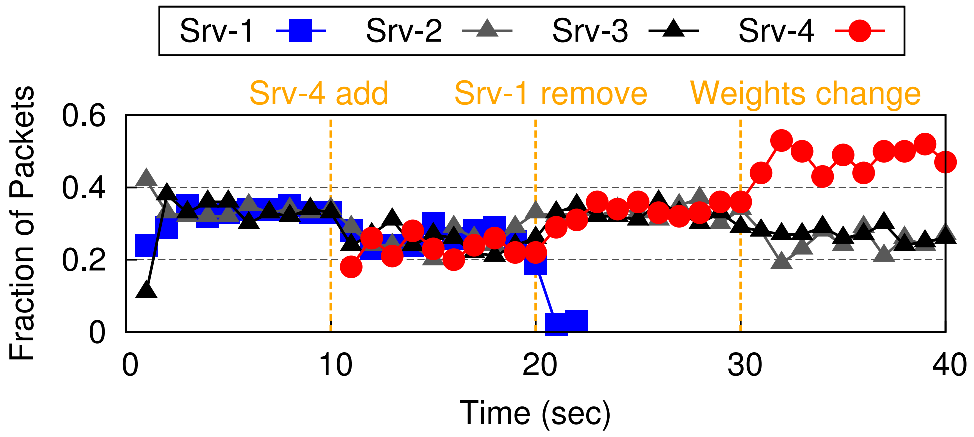


Fig. 4.14.: YODA user policy update.

4.6.3 Scalability

In this experiment, we show that YODA scales as the traffic coming to the load balancer increases without breaking existing connections. Figure 4.13 shows the request rate (req/sec) and CPU utilization during a 30-second long experiment, where we send requests using the `Apache benchmark` tool to fetch individual objects. In this experiment, we initially have 6 YODA instances. At time = 10 sec, we increase the traffic for each YODA instance from 5K req/sec to 10K req/sec, which increased the CPU utilization on the YODA instances from 40% to 80%. As a result, the YODA controller adds 3 more instances to reduce the CPU utilization which reduced the traffic on each YODA instance to 6.7K req/sec, and CPU utilization to 60%. Importantly, all client flows were maintained throughout the experiment.

Lastly, we also measure the latency throughout the experiment. Surprisingly, we do not see any significant variation in latency during the dynamics. The latency during the 10-15 second interval under higher traffic load is not inflated because the queues on the YODA instances will not build up until the CPU saturates. The same was observed with the software Mux in Ananta [2].

4.6.4 Safe policy update

In this experiment, we show that Yoda can safely update user policies without breaking existing flows. We also show that the weights set to split the traffic in user policies are correctly implemented by the distributed Yoda instances.

In this experiment, initially the user policy specifies equal weights among the 3 backend servers running on identical VMs. The operator wants to replace one of the VMs with a VM that has 2x CPU cores. To do this, the operator uses make-before-break, where at 10 sec, the operator adds the new VM (Srv-4), and at 20 sec, the operator removes one of the existing VMs (Srv-1). Lastly, at 30 sec, the operator changes the weights among the servers (Srv-2,3,4) to 1:1:2.

Figure 4.14 shows the dynamics during the 40-second interval. We see that the fraction of traffic and CPU utilization (not shown) change according to the policy changes. Between 0-10 sec, server -1 to -3 receive equal traffic. As the new VM is added at 10 sec, the traffic is now split equally across the 4 servers. At 20 sec, as Srv-1 is removed, the traffic is split equally among the 3 servers. Finally, at 30 sec, the traffic is split following the 1:1:2 ratio. Again, none of the client flows were broken as Yoda adapts to the user policy changes (§4.4.2).

4.7 Simulation

In this section, we evaluate Yoda's VIP assignment algorithm (§4.3.4) using a traffic trace production cloud. We show that: (1) Yoda reduces the L7 LB cost by 3.7x on average when all online services share the LB. (2) Yoda update algorithm is effective in limiting the transient traffic overload and flows migrated;

Setup: The traffic trace collected from our production cloud consists of all flows received by the Internet-facing services in a 24-hour period (during a weekday). The trace consists of 100+ VIPs and 50K+ L7 rules. We calculate the assignment between the VIP and the Yoda-instances every 10 mins. The VIP assignment algorithm finishes in 1.5-21.5 sec (median 3.92 sec) using the CPLEX ILP solver.

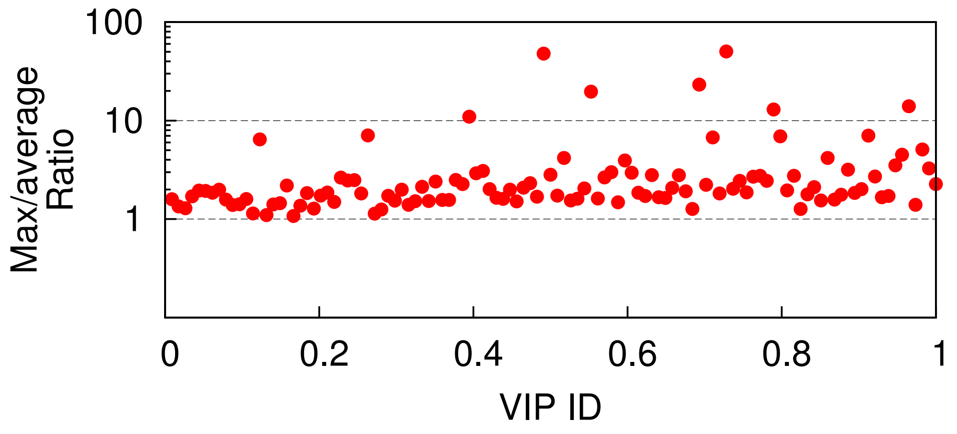


Fig. 4.15.: Max-to-average traffic ratio for all VIPs. VIPs are sorted in decreasing order of the traffic.

4.7.1 Cost reduction

We estimate the cost reduction based on the ratio of the max. to average traffic for the individual VIPs, as the max-to-average ratio indicates the cost savings possible by using the elasticity of Yoda-as-a-Service. This is because in using HAProxy, each individual online service would need to provision LB based on its peak traffic demand, as dynamically removing/adding HAProxy instances could break the connections, whereas in using Yoda-as-a-service, an online service can easily scale up/down its share of LB instances without breaking connections, and just has to pay for its share of Yoda instance usage, and over time, its average LB instance usage is proportional to its average traffic load. Note we are ignoring discretization effect in both cases, which makes the saving conservative.

We calculate the ratio of the max-to-average traffic for individual VIPs in every 10-min interval throughout 24 hours. Figure 4.15 shows the max-to-average ratio for all the VIPs. The VIPs are sorted in decreasing order based on their traffic volume. By using Yoda, these online services can save L7 LB cost by 1.07x to 50.3x. (average = 3.7x across all VIPs).

Next, we show that YODA can provide these cost benefits by frequently calculating (every 10 mins) the VIP to YODA instance mapping (§4.3.4), and updating it through congestion-free update (§4.3.5).

4.7.2 Impact of updates

We evaluate the effectiveness of YODA's assignment algorithm in terms of (1) the number of YODA instances required, (2) fraction of the YODA instances overloaded during transition, (3) number of flows migrated. The smaller these values, the more effective the assignment algorithm.

We set the target latency due to YODA to 5 msec, which translates into 2K rules on each YODA instance (Ry=2K) based on Figure 6. We set the limit on the number of flows to be migrated to 10%. Lastly, we set $n_v = 4 \cdot \frac{t_v}{T_y}$, *i.e.,* each VIP gets 4x more replicas by using YODA as a shared service than using YODA individually.

We compared two versions of the algorithm: (1) YODA-no-limit where there is no limit on the transient traffic or number of flows migrated, (2) YODA-limit, which provides congestion-free transition, and may require more instances.

Numbers of rules: Figure 4.16(b) shows the median number of rules across the LB instances that YODA-no-limit and YODA-limit generate normalized to that by the all-to-all scheme. We see that YODA instances store 0.5-3.7% (median 1%) of the rules compared to Ananta, which reduces the number of rules in YODA by 100x compared to Ananta. But this also comes at a cost of increasing number of instances.

Number of instances: Figure 4.16(c) shows the number of YODA instances for YODA-limit and YODA-no-limit. As a reference, we also show the result for a base-line all-to-all assignment, which requires least number of instances *i.e.,* the total traffic divided by traffic capacity of each instance.

We make two observations: (1) YODA-no-limit (and -limit) requires 4.6-73% (average = 27%) more instances than the all-to-all. This is the overhead of reducing the number of rules. (2) Importantly, YODA-limit only requires up to (-8) – 11.7% more instances compared to Yoda-no-limit (median 1.3%). In some cases Yoda-limit required lesser instances because optimality gap in CPLEX was set to 10%. This shows that Yoda assignment algorithm is effective in *packing* VIPs under the constraints and the overhead of providing congestion-free transition is very small.

Transient overload: Next, Figure 4.16(d) shows over the 24-hour period, Yoda-no-limit results in 0-20.4% (median 5.3%) of the Yoda instances overloaded during transition, which is significantly reduced in Yoda-limit. The instances that were overloaded in Yoda-limit were already overloaded before starting the new round. This also emphasize the need to frequently update the VIP assignment.

These result confirm that Yoda-no-limit results in significant number of Yoda instances overloaded during transition compared to Yoda-limit.

Number of flows migrated: We observe that under Yoda-no-limit 2.7-95% (median 44.9%) flows migrated from one Yoda instance to other, but under Yoda-limit the number is significantly lower at 0-29.8% (median 8.3%) (not shown). We set the limit on the number of flow migration to 10%, but the LP gave infeasible assignment at two points. In such cases, we increased the limit by increments of 10%, and the LP gave a feasible assignment when the limit was 30%.Yoda-limit was

4.8 Questions and Answers

We discuss a few questions that are frequently asked.

Q: Is it necessary to separate L4 and L7 LB?

A: We propose a separate L7 LB service for two reasons: (1) A large fraction of the traffic in datacenters do not need to go through L7 LB but need L4 LB. (2) Building the L7 LB service on top of L4 LB follows the module design principle.

Q: Why are there thousands of rules per tenant?

A: The rules are specified on the HTTP, TCP and IP fields. There could be a large number of rules, as there are billions of URLs and cookies.

Q: Does Yoda instance access TCPStore for every packet?

A: No. TCPStore is accessed while establishing connection and after Yoda instance failures.

Q: Does an L4 instance failure affect Yoda's reliability?

A: No. L4 LB has built-in resilience to instance failures [2].

Q: Can the L4 LB failure resiliency mechanism used in L7?

A: No. L4 LB selects the server by hashing TCP/IP tuples, which are carried by every packet of a flow. In L7 LB, the HTTP header used for server selection is not embedded in every packet.

4.9 Related Work

To our knowledge, Yoda provides the first highly available cloud-scale L7 LB design. In the following, we review related work on other middlebox functions in the cloud and network state management and migration.

Middleboxes in cloud: Several startups provide middlebox functions in the cloud. (*e.g.,* Aryaka [47], Barracuda [38] Qualys [48]). APLOMP [49] helps online operators split middlebox functionality between the cloud and enterprise, which motivates the need for 3^{rd} party offered middlebox functionalities in the cloud. Additionally, FTMB [39] and [40] show the evidence of middlebox failures in cloud which facilitates the need for highly available middleboxes.

Layer-4 LB: There are several designs proposed for scalable layer-4 LB in cloud. Ananta [2] uses software instances, whereas Duet [4] and Rubik [50] use hardware switches and software instances for load balancing. The principles used by layer-4 LB designs in providing availability cannot be used in layer-7 LB. Specifically, the layer-4 LBs use IP 5-tuple to select the server, and all the packets on a given connection carry the same IP tuple. Therefore, even if the packets on the same connection reach different LB instances, the LB instances select the same backend server. On the other hand, the HTTP header is only in the first few packets, which necessitates the need to store the selected backend persistently.

State management: OpenNF [51] helps to scale middleboxes safely by moving the flows and state from one middlebox to another. However, its focus is on providing elasticity and not failure resiliency, and thus it does not explicitly decouple flow state and store them in a persistent storage for transparent flow recovery in case of failure. FTMB [39] has high cost of maintaining the state. SSM [52] maintains user sessions, which Yoda can leverage to maintain user sessions in addition to connections. Prior work has looked at seamless migration of the BGP sessions and router failures [53,54], which is different goal that Yoda.

Request redirection and CDN: DONAR [55] and Oasis [56] focus on request redirection across multiple DCs and enterprises through DNS. Centrifuge [57] focuses on lease management of accesses to data that are partitioned across in-memory servers. It does not recover state lost in failure and assumes applications will recreate the lost state.

TCP splicing, handoff and migration: Application layer proxies typically use TCP splicing [58, 59] to bind the sockets sockets connected to the server and client. But these designs break the flows when proxies fail. Work on TCP handoff [60, 61] focus on bypassing the LB instance (front-end) on the reverse path and do not address LB instance failure. TCP Migration also tries to maintain connections when servers fail or in the presence of IP mobility. But they require changes at the clients transport layer [62, 63] or socket API [64] or assigning the same IP to all servers [65, 66] which cannot work in the cloud.

Rule assignment and packet processing: Although rule management is not a main focus, Yoda can benefit from the recent work on rule management and fault-resilient updates [22, 23, 67–70]. Additionally, Yoda can leverage recent works to improve its packet processing performance within individual instances [26, 71].

4.10 Summary

YODA is a distributed L7-LB-as-a-service designed to meet the availability, scale and operational requirements of multi-tenant clouds. The high availability in YODA is attributed to three design choices: (1) decoupling the TCP state from individual instances and storing it in a persistent in-memory storage (called TCPStore). (2) A novel mechanism that enables one instance to re-use TCP state created on different instance. (3) Using virtual IP to establish connections with the clients and servers so that the connections are shielded from the failure and other dynamics within the YODA instances (fron-tand-back indirection). Our evaluation of YODA using a prototype implementation and simulations using a production traces shows that YODA can transparently restore flows with low latency overhead in decoupling the state (0.89 msec) while meeting high availability and scalability requirements. Additionally, compared to individual tenants using the load balancer service in isolation, YODA as a shared service can reduce the L7 load balancing cost by 3.7x.

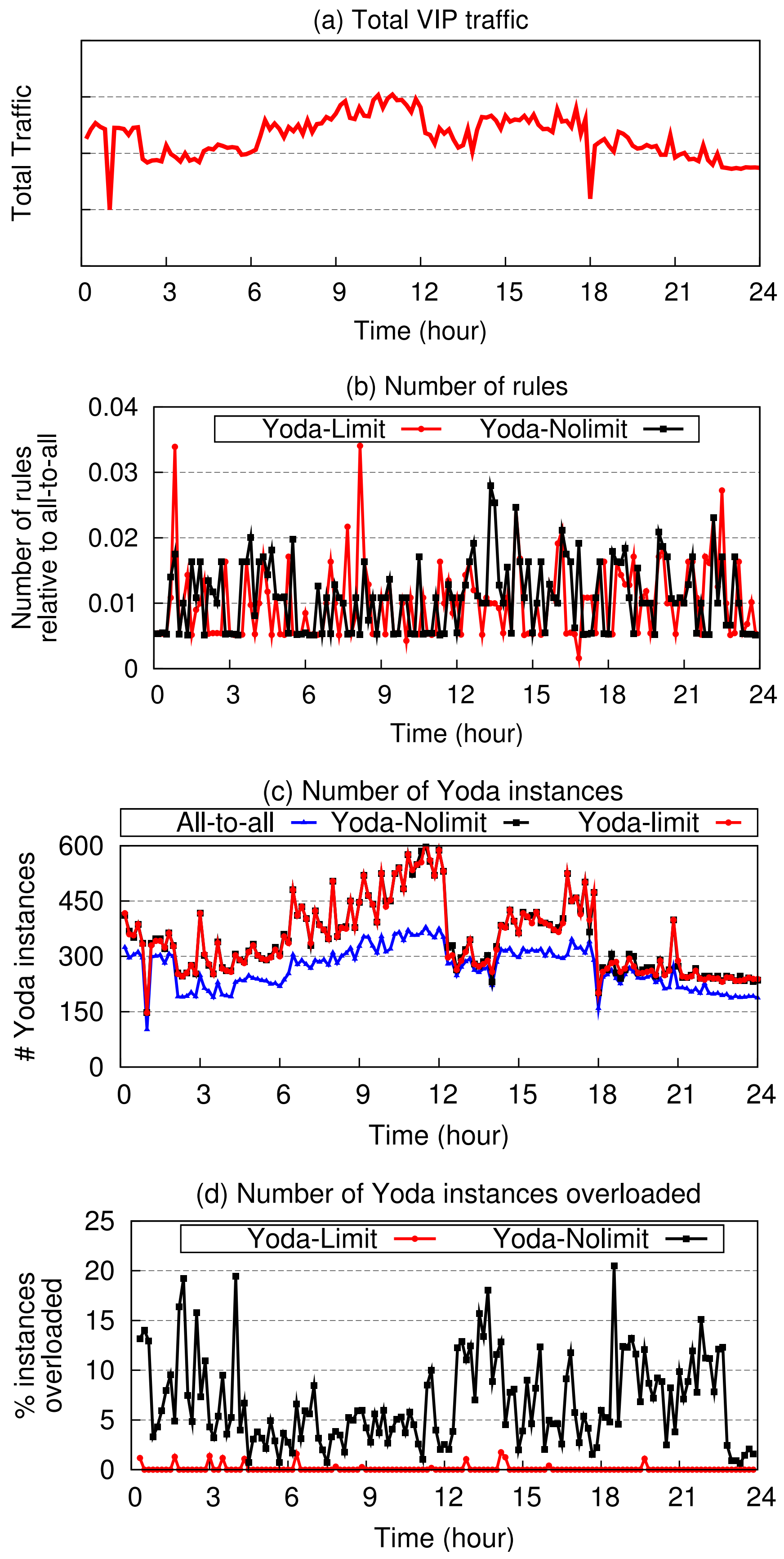


Fig. 4.16.: YODA update.

5. CONCLUSION

In this thesis, using load balancer as an example, we revisit existing cloud scale middlebox designs to identify their fundamental shortcomings. Then we propose a new class of distributed hybrid middleboxes designed to provide high capacity, low latency, high availability, and high flexibility at low cost. The DUET load balancer design was motivated by two key observations: (1) software load balancers offer high availability and high flexibility but suffer high latency and low capacity per load balancer, and (2) commodity switches have ample spare resources and now also support programmability needed to implement load balancing and other major layer-4 middlebox functionalities. The DUET architecture seamlessly integrates the switch-based load balancer design with a small deployment of software load balancer. We evaluate DUET using a prototype implementation and extensive simulations using traces from the production DC.

While DUET achieves significant improvements in terms of performance and scalability, it continues to suffer from the network bandwidth overhead problem. This thesis proposes RUBIK to drastically reduce the bandwidth usage while providing low cost, high performance and reliability benefits of Duet. RUBIK achieves this by exploiting two design principles: (1) *locality*: it load balances traffic generated in individual ToRs across DIPs present in the same ToRs, (2) *end-point flexibility*: it places the DIPs closer to the traffic sources. We evaluate RUBIK using a prototype implementation and extensive simulations using traces from the production DC.

Lastly, current layer-7 middleboxes suffer from poor availability, and the design choices that enabled high availability in the layer-4 middleboxes cannot be extended to the layer-7 middleboxes. Additionally, the lack of offerings from the major cloud providers force the tenants to design and maintain their middleboxes on their own. This thesis proposes YODA – a distributed L7-LB-as-a-service designed to meet the availability, scale and operational requirements of multi-tenant clouds. The high availability in YODA is attributed to three design choices: (1) decoupling the TCP state from individual instances and storing it in a persistent in-memory storage (called TCPStore). (2) A novel mechanism that enables one instance to re-use TCP state created on different instance. (3) Using virtual IP to establish connections with the clients and servers so that the connections are shielded from the failure and other dynamics within the YODA instances (front-and-back indirection).

Our evaluation shows that DUET provides 10x more capacity than a software load balancer, at a fraction of its cost, while reducing the latency by over 10x, and can quickly adapt to network dynamics including failures. Further, compared to DUET, RUBIK can reduce the bandwidth usage by the load balanced traffic by over 3x. The evaluation of YODA using a prototype implementation and simulations using a production traces shows that YODA can transparently restore flows with low latency inflation ($<$ 1msec) while meeting high availability and scalability requirements. Additionally, compared to individual tenants using the load balancer service in isolation, YODA as a shared service can reduce the L7 load balancing cost by 3.7x.

6. FUTURE WORK

We believe further improving middlebox infrastructure in cloud is an exciting and rich space of research. Some of the problems in middlebox infrastructure that we want to explore in the future are outlined below.

6.1 Middleboxes with RDMA

RDMA over Ethernet (specifically RoCEv2 [72]) is replacing the current TCP/IP based network transport at a fast pace [73, 74]. The key driving force for this transition is the low latency and high throughput benefits in RDMA as it bypasses OS and reduces the data copy overhead when transferring the data between servers. The existing literature has focused on integrating RDMA with storage systems [74–76]. These works overlook designing middleboxes for RDMA since storage systems do not heavily use middleboxes. However, since the web applications heavily use middleboxes, and as RDMA makes powerful strides in integrating with the web applications, there is a strong need for revisiting middlebox design for RDMA.

Designing high performance and highly available middleboxes with RDMA raises many new challenges in both layer-4 and layer-7 middleboxes. The key challenges stem from: (1) Existing TCP/IP based software middleboxes [2, 77] use network functionalities such as encapsulation/decapsulation and direct-server-return (DSR) to easily enable routing to/from middleboxes. Such network functionalities are very useful to provide high scalability and availability and to easily handle failures. However, these functionalities are not available in RDMA as packets bypass the OS stack, which makes providing high scalability and availability an onerous task; (2) The challenges are further amplified due to various dynamics such as failure, migration and congestion; (3) As an alternative, the packets could go through the OS to implement network functions. However, such design can inflate the latency, and could diminish or even eliminate the benefits promised by RDMA; (4) Providing high availability to layer-7 middleboxes is even harder. Designs such as Yoda that decoupled and stored TCP state are not feasible in RDMA settings due to their roughly 40x latency overhead in such settings. Alternatively, hardware middleboxes cannot implement all middlebox functionalities due to limited resources.

6.2 What made specialized hardware middleboxes obsolete

The key reason for the fast penetration of NFV is because traditional specialized hardware middleboxes are too costly, and have poor scalability and availability in cloud settings. However, specialized hardware middleboxes still offer better performance by using FPGAs and ASICs tailored specifically for the middlebox functions. I want to take a step back and have a closer look at the specialized hardware middleboxes to determine the key root causes for their downfall. In the next step, I want to design new genes of specialized hardware middleboxes that offer the same benefits such as better cost, scalability, availability and programmability as NFVs but deliver higher performance. Secondly, I also want to take a look at middleboxes in WAN. Many cloud operators have reported that the DDoS-type traffic should be blocked before hitting the servers (or NFVs) within data-center, as such traffic strangles the precious access bandwidth. I want to design middleboxes that handle such malicious traffic in WAN itself to protect other resources.

6.3 New Datacenter Hardware

Datacenter is a fertile playground to induct new hardware and technologies to continually improve the performance of online services. In one such case, new NICs are designed with FPGA to offload custom network functions ranging from simple tasks like computing check-sums to more complicated ones such as SSL offload. The

FPGA on the NIC can be leveraged to offload middlebox functions that can lower the latency overhead and relieve CPU load related to the middlebox functions.

New hardware such as FPGA powered NICs provide new opportunities and challenges in designing middleboxes. Prior middlebox designs assumed two classes of resources – switches (hardware) and servers (software). The FPGA powered NICs add another class to this list, as they have capacity and speed different than those of the switches and servers. It is unclear how to architect middleboxes around different classes of resources with varying performance. I view this problem as similar to designing distributed applications when the available storage resources (DRAM, SSDs and disks) have varying performance, and application performance depends on how applications leverage underlying storage resources.

Lastly, the datacenters are adopting optical and wireless networking which enables dynamic topologies, where the path between two end-points changes dynamically. Such topologies have implications on maintaining a coherent state of flows across middleboxes. As the path changes dynamically, it makes the traffic pass through more than one middlebox instance, which needs *fast state synchronization* across instances. I want to study implications of the dynamic topologies on middleboxes in the cloud.

6.4 The only constant is change

Apart from middleboxes, I also view network management as a rich space for research. Specifically, I want to investigate and build better schedulers for: (1) network updates that can induce a certain degree of sub-optimality in exchange for higher update speeds; (2) CoFlow abstraction to improve the job completion time of the analytic jobs by observing that current schedulers completely miss out on the spatial dimension (ports where flows of CoFlows arrive), which results in poor performance. Next, I want to carefully fuse the spatial dimension to improve CoFlow and job completion times at scale.

Lastly, today many of the datacenter workloads are dominated by search and big-data applications. Many different applications including deep neural network and bio-informatics are on the horizon and have very distinct requirements in terms of storage, low-latency processing, privacy and security. I believe that as these applications become mainstream, they hold the potential to shake the design principles used to build today's cloud infrastructure, and would require revisiting many of the decisions made in designing today's cloud infrastructure.

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