Backpressure Flow Control

Prateesh Goyal\textsuperscript{1}, Preey Shah\textsuperscript{2}, Naveen Kr. Sharma\textsuperscript{3}, Mohammad Alizadeh\textsuperscript{1}, Thomas E. Anderson\textsuperscript{3}
MIT CSAIL\textsuperscript{1}, IIT Bombay\textsuperscript{2}, University of Washington\textsuperscript{3}

Abstract

Effective congestion control in a multi-tenant data center is becoming increasingly challenging with rapidly increasing workload demand, ever faster links, small average transfer sizes, extremely bursty traffic, limited switch buffer capacity, and one-way protocols such as RDMA. Existing deployed algorithms, such as DCQCN, are still far from optimal in many plausible scenarios, particularly for tail latency. Many operators compensate by running their networks at low average utilization, dramatically increasing costs.

In this paper, we argue that we have reached the practical limits of end-to-end congestion control. Instead, we propose a new clean slate design based on hop-by-hop per-flow flow control. We show that our approach achieves near optimal tail latency behavior even under challenging conditions such as high average link utilization and in-cast cross traffic. By contrast with prior hop-by-hop schemes, our main innovation is to show that per-flow flow control can be achieved with limited metadata and packet buffering. Further, we show that our approach generalizes well to cross-data center communication.

1 INTRODUCTION

Multi-tenant data centers have become one of the largest and fastest growing segments of the computer industry. Data centers are increasingly dominating the market for all types of high end computing, including enterprise services, parallel computing, large scale data analysis, fault tolerant middleboxes, and global distributed applications [5, 14, 25]. These workloads place enormous pressure on the data center network to deliver, at low cost, ever faster throughput with low tail latency even for highly bursty traffic [11, 36].

To help address these demands, RDMA (Remote Direct Memory Access) has emerged as a popular protocol for very high-bandwidth end host communication. With RDMA, applications directly read and write protected regions of memory on other machines, with the protocol implemented in hardware without the need to interrupt or process requests in software on the remote node, and with an assumption that network packet losses are very low. As such, RDMA is simple enough to be implemented for 100 Gbps links and beyond [24].

An ongoing challenge for deploying RDMA, particularly for multi-tenant data centers, is network congestion control. A commonly used option is PFC (Priority Flow Control) [34], where switches send blanket pause frames one hop upstream to prevent buffer overflow and packet loss. These pause frames can cause head-of-line (HoL) blocking where unrelated traffic is stalled waiting for other flows to release downstream buffers. To address this, most data center operators add some form of end-to-end congestion control to avoid triggering pause frames. Two examples are DCQCN [38] and HPCC [22]; these use sophisticated control theory to rate-limit individual flows to better manage resources in the middle of the network.

Unfortunately, tail latency performance – often the key metric for applications using RDMA – degrades with DCQCN or HPCC under load, to a far greater degree than would be predicted by queueing theory. One could add better switch scheduling algorithms, such as fair queueing, but we show that even the combination of DCQCN and fair queueing remains far from optimal for tail latency for typical network workloads and configurations. This gap is likely to widen as network speeds increase.

Instead, we propose a different approach, to revisit the idea of per-hop per-flow flow control. The key challenge for data center networks, in our view, is to efficiently allocate buffer space at congested network switches. This becomes easier and simpler when control actions are taken per flow and per hop. Per-hop per-flow flow control is of course not a new idea, having been introduced, and discarded, with ATM networks over two decades ago [6, 19]. These earlier schemes required per-flow state at each switch even for quiescent flows, an amount of state that would not be practical in today’s data center networks. Our principal insight is to show that per-hop per-flow flow control can be approximated with a small amount of switch state and a modest amount of signalling overhead. We note that our approach is approximate, and therefore it is not completely loss-free.

We evaluate our approach, called BFC (Backpressure Flow Control), using ns3 [1] on synthetic traces drawn to be consistent with measured workloads from Google and Facebook data centers [28], for a small-scale multi-level Clos network topology. We show that BFC, unlike DCQCN and HPCC, achieves close to the optimal tail latency performance on these workloads across a wide range of flow transfer sizes, even when the network is reasonably heavily loaded and is disrupted by a small amount of incast traffic.

Our specific contributions are:

\begin{itemize}
  \item A detailed protocol for per-hop per-flow flow control, called BFC, that uses a small amount of metadata to achieve near-optimal tail-latency performance for typical data center workloads.
  \item An evaluation of BFC against a set of practical and idealized congestion control algorithms on a set of synthetic traces drawn from measured data center workloads.
\end{itemize}

Our code and simulation harness is public and can be made available online. Our code and simulation harness is public and can be made available online.
available to the reviewers on request. There are several important limitations to our work. (i) The source code for HPCC was made available only in the past week; as a result, we only present a partial comparison of BFC versus HPCC, along with a more complete comparison of BFC versus other approaches. (ii) Our approach is designed to be able to be combined with different switch scheduling algorithms, such as hierarchical round robin and priority scheduling. In this paper, however, we only compare to fair queueing. (iii) We have not completed a detailed hardware design for BFC, and so we are unable to give a precise estimate of the VLSI chip area needed for its implementation. BFC could be implemented in P4, but it requires scheduling control beyond the capability of current generation Openflow v2 switches. (iv) Our evaluation is limited to RDMA congestion control algorithms. We believe our hop-by-hop approach will compare favorably to TCP congestion control [3, 32], but we do not present a quantitative comparison at this time.

2 BACKGROUND

Remote Direct Memory Access (RDMA) has become an increasingly popular paradigm for data center network communication. For applications, RDMA offers extremely low overhead and high bandwidth, with message handling implemented in hardware without OS kernel mediation. And with RDMA over Converged Ethernet (RoCEv2) [17], RDMA can run on any switched Ethernet physical layer.

Because server network interface cards (NICs) have long been designed to provide high bandwidth for large TCP transfers, the primary benefit of RDMA has been for applications with fine-grained sharing between nodes. These applications are often latency sensitive [10], and when work is farmed out in parallel, tail-latency sensitive [11]—application performance often depends on the worst case communication time, rather than the average case.

An ongoing challenge for RDMA networks is how best to handle congestion. Because message handling is implemented in hardware, RDMA systems typically use simple error handling, such as Go-Back-N, that make performance highly vulnerable to packet loss. To address this, most RDMA-enabled data centers deploy a hop-by-hop flow control mechanism called PFC (Priority-based Flow Control). (For simplicity, we will focus on the case where there is congestion among the traffic at a particular priority level.) With PFC, if the packets from a particular input port start building up at a congested switch (past a configurable threshold), the switch sends a pause frame upstream stopping that input from sending more traffic until the switch has a chance to drain. This prevents the switch buffers from being overrun and needing to drop packets.

Unfortunately, PFC has a side effect: head-of-line (HoL) blocking [38]. For example, suppose a congested switch is receiving an incast of packets from several of its input ports; the incast will arrive more quickly than it can drain, building up buffers, and causing PFC pause frames to be sent upstream on those input ports. This will prevent packet loss, but only at the cost of pausing unrelated traffic that traverses the paused link but is destined for other uncongested output ports. These flows will be delayed until the incast packets can be drained. Worse, as packets queue up behind a PFC, additional upstream PFCs can be triggered, widening the scope of HoL-blocking. Even without incast, random fluctuations in load can cause short-term congestion and PFC pauses, leading to HoL-blocking of unrelated traffic. HoL-blocking is particularly problematic for tail-latency performance, as short flows can be queued behind much longer flows. Any congestion hurts tail latency, but HoL-blocking delays are often a large multiple of the delays a packet would experience in an idealized non-HoL blocking network at the same load level.

To reduce HoL-blocking, researchers and data center operators have proposed several end-to-end congestion avoidance systems that aim to reduce buffer occupancy and thus the need for PFC pause frames. Timely [26], DCQCN [38], and HPCC [22] monitor gather information from the network in response to packet transmissions to gate how quickly to send additional packets. Because earlier results have shown that DCQCN outperforms Timely [22, 39], we focus our comparison on the later two. DCQCN and HPCC can be effective when congestion is due to so-called elephant flows: large, many-round trip transfers that can be throttled up or down based on network conditions to avoid HoL-blocking for short, latency sensitive flows. However, measurements of data center network workloads suggest that most traffic completes within at most a few round trips; simply prioritizing short flows over long flows is insufficient when the network is overwhelmed with short to moderate flows.

We next argue that these trends are getting worse over time.

**Trend 1: Rapidly increasing link speed** Fig. 1 shows the switch capacity of top of the line data center switches manufactured by Broadcom [9,29]. Switch capacity and link speeds have increased by a factor of 10 over the past six years with no signs of stopping. As a result, much more data can be transmitted within a single round trip; for example, with a 100 Gbps network link and 12 us round trip delay, there can be 150 KB in flight before any feedback is received at the endpoint.

Larger bandwidth-delay makes it harder for end-to-end congestion control to manage network buffers. Fig. 2 shows the cumulative distribution function (CDF) of buffer occupancy with DCQCN across different link speeds, using the T2 topology (from §4) and the Google workload from Section 4 set to 75% load plus 5% incast, but without PFC enabled. DCQCN fails to limit buffer occupancy at higher link speeds, if utilization is kept constant.

**Trend 2: Buffer size is not scaling with switch capacity** Fig. 1 shows that the switch buffer size relative to its capacity has decreased by a factor of 2 (from 80 us to 40 us) over the past six years. With smaller buffers relative to link speed, buffers now fill up more quickly, making it more difficult
Our primary goal is to solve what we call the data center network CUT problem: for a high capacity (C) network where most flows, and most bytes in flows, fit inside the end-to-end bandwidth-delay product of the network, we want both (i) efficient network utilization (U), that is, low network operating cost, and (ii) low tail latency (T) for flow completion, that is, better application performance. Obviously, U and T are in conflict – if the workload can be modeled as Poisson or more heavy-tailed arrivals, then any increase in utilization will hurt flow completion times and tail latency to an even greater degree. Our goal is not to force the network operator to choose how best to trade off application performance for end-to-end congestion control to manage those buffers. Using 100 Gbps links and the same topology and workload as for Trend 3: Flow sizes remain small Fig. 4 shows the cumulative bytes contributed by different flow sizes for three different industry workloads. Although data processing jobs are typically considered to have large flows, almost all Facebook Hadoop traffic is likely to fit within a single round trip within the next few years, with websearch graph analysis to follow shortly thereafter.

There are several consequences. Even if we develop protocols to gather precise information about the state of the bottleneck switch and return it to the sender, that information could well be out of date by the time the sender could act on it. Even if congestion is stable, end-to-end adaptation is inherently iterative. The rate for any particular flow depends on secondary bottlenecks encountered by other flows as they simultaneously change their rate. Equilibrium sharing is reached with DCQCN only after many round trips. Flows could be intentionally slowed down to provide more consistency over time, but this would come at a cost in substantial added latency (and tail latency).

While some have argued that data center flows are (or should be) increasing in size [2], widespread use of RDMA is likely to be pushing in the opposite direction, as it supports fine-grained memory access as a programming model.

**Trend 4: Multi-data center communication** Fault tolerance concerns are leading many data center applications to be spread over clusters of nearby data centers. Protocols for homogeneous environments, similar link speed and latencies, need to be generalized to these emerging workloads. With high bandwidth connectivity between data centers and much longer round trip times, as well as the greater operating costs of long haul fiber encouraging higher utilization, these settings are likely to put even greater pressure on end to end congestion control.

**3 BFC DESIGN**

Our approach is to design a practical system for per-hop, per-flow flow control. We first describe the goals and constraints on our design, we then sketch a plausible straw proposal that surprisingly turns out to not work well at all, and we use that as motivation for our design, which we evaluate in the following section.

**3.1 A Practical Solution to Network CUT**

Our primary goal is to solve what we call the data center network CUT problem: for a high capacity (C) network where most flows, and most bytes in flows, fit inside the end-to-end bandwidth-delay product of the network, we want both (i) efficient network utilization (U), that is, low network operating cost, and (ii) low tail latency (T) for flow completion, that is, better application performance. Obviously, U and T are in conflict – if the workload can be modeled as Poisson or more heavy-tailed arrivals, then any increase in utilization will hurt flow completion times and tail latency to an even greater degree. Our goal is not to force the network operator to choose how best to trade off application performance...
against system cost, but simply to make efficient that tradeoff.

Our work is largely orthogonal to switch scheduling policy, but for concreteness, we will constrain our discussion to fair queueing. (One could equally combine our approach with hierarchical round robin, priority scheduling, and so forth. These would have different scheduling metrics.) We can then use an idealized fair queued switch with infinite buffering as a benchmark. Ignoring hardware constraints that would make such a switch impractical to build, it would do well at CUT: flows could be safely sent into the network at line rate, be queued at any bottleneck, and complete in a manner that (roughly speaking) would minimize both average and tail completion time by ensuring that every flow gets an equal share of the bottleneck link capacity.

We observe that for typical FIFO switches with limited buffering, and typical data center workloads, end-to-end congestion control cannot provide all three CUT properties. At best, we must choose two out of the three. We can support moderate utilization (U) and high capacity links (C), as long as we spread transmissions over multiple round trips so that network probes measure a stable property. These extra round trips result in higher latency. Or we can over-provision the network, so that congestion is a non-issue, allowing us to have high capacity links (C) with low tail latency (T). Or with a low capacity network, most flows will take many round trips, enabling end-to-end congestion control to be effective.

We want our approach to be practical, so we add three design constraints:

- In-order packet delivery: Packets from the same flow should be scheduled out the egress port in the same order that they arrived at the switch ingress, to simplify packet handling at the destination server.
- Zero configuration: Our solution should require minimal configuration. While it is possible to auto-tune parameters, a system with simpler and more intuitive configuration will be more likely to work robustly across rapidly evolving workloads and topologies.
- Hardware plausibility: We assume hardware design constraints typical of modern switches. For example, we assume that there are a limited number (32 in our simulations) of FIFO queues per output port, that are scheduled in deficit round-robin or priority order. We can independently pause and unpauses each queue without preventing or slowing down forwarding from other FIFO queues. However, if we pause a flow, we must pause all of the packets that share the same FIFO queue. In other words, head-of-line blocking is an inherent challenge caused by hardware design constraints; we cannot assume that the switch scheduler can pick arbitrarily from any of a large number of queued flows, in constant time. Similarly, any metadata should be small, relative to the space required for switch packet buffers.

### 3.2 Straw Proposal

We originally thought the following design would meet our goals. Following stochastic fair queueing [23], compute the egress port for every arriving packet. Use a hash function on the flow header to consistently assign the packets of each flow to the same randomly-chosen FIFO queue at the egress. We selectively pause a flow (at the upstream switch) if its queue depth becomes too large, e.g., one per-hop bandwidth-delay product. Above this point, we would (likely) have time to signal the upstream switch to resume sending before the queue completely drained. Of course, there could be contention at the upstream switch that would prevent the flow from refilling quickly enough, but let us leave that aside for the moment.

When we need to pause a flow, we need to pause all flows sharing the same FIFO queue at the upstream switch. To simplify matters, for the straw proposal, let’s assume that we use the same hash function at each switch; that way, all flows from the same ingress port to the same egress port, and assigned to the same FIFO queue, will be paused together. Unlike PFC, pauses in the straw proposal only affect individual flows; flows in other queues can continue to be forwarded downstream and scheduled out of the switch.

It seems like we should be done, and this paper could be short. However, the straw proposal does not work well, for three reasons. First, the birthday paradox causes worse tail latency at the congested switch than you might think. With even a modest number of flows, there is a significant chance that any specific flow will land in an already busy FIFO queue. For example, with 10 flows assigned to 32 FIFO queues, it is highly likely that two flows will be randomly assigned to the same queue, causing HoL-blocking and hurting tail latency. Second, if we need to pause a flow because a FIFO queue becomes too full, any flow at any upstream switch that hashes to the same queue will potentially be affected by the pause. In other words, the per-flow pauses will propagate upstream to affect other flows with the same hash value, even if they have different downstream routes.

These problems occur because flows experience hash collisions, even when we have enough physical resources to give each flow its own dedicated FIFO queue. We should expect this to be the common case. With Poisson flow arrivals and where every flow sends as fast as possible until it is done (as in RDMA), queueing theory [20] states that the expected number of queued flows at a link is $1/(1-U)$ where $U$ is the utilization of the link, and the variance in the queued flows is $1/(1-U)^2$. A link with 75% utilization (that is, moderately congested) will have an average of 4 queued flows, and a variance of 16, well under the typical number of physical queues, but enough to make the birthday paradox matter.

The third problem is buffer management. Recall that the buffering in switches is trending downward relative to switch capacity, to a small multiple of the end-to-end bandwidth delay product. The straw proposal attempts to keep one hop of bandwidth delay per flow at the bottleneck switch. In reality,
Mechanism
Low bandwidth signalling of paused flows
Dynamic assignment to physical queues
Efficient use of scarce physical queues
VFID = Hash
Bloom filter of Hash
Shared hash table indexed by the VFID
Shared associative state indexed by the VFID
Space efficient way of storing state for virtual flows
Track active flows at a switch
Track overflows in the Virtual Flow Hash Table

Table 1: Key design components of BFC.

| Function                  | Mechanism                                  | Purpose                                           |
|---------------------------|--------------------------------------------|--------------------------------------------------|
| Virtual Flow              | VFID = Hash$_{FID}$ (FlowID)               | Track active flows at a switch                    |
| Physical Queue            | Dynamic assignment to physical queues      | Efficient use of scarce physical queues           |
| Pause Frame               | Bloom filter of Hash$_{BF}$ (VFID)         | Low bandwidth signalling of paused flows          |
| Virtual Flow Hash Table   | Shared hash table indexed by the VFID      | Space efficient way of storing state for virtual flows |
| Overflow TCAM             | Shared associative state indexed by the VFID | Track overflows in the Virtual Flow Hash Table    |

however, this is overkill, as flows can only make use of one hop of bandwidth delay in aggregate. Suppose we have an incast of flows at a congested switch. With enough incoming flows, where each flow stores a bandwidth-delay at each hop, we could completely run out of buffer space and be forced into PFC-style blanket pauses. These in turn would cause cascading upstream HoL-blocking. Even without incast, moderate to severe congestion among a few hundred moderate-length flows could lead to periods of buffer space exhaustion, with no compensating benefit in system performance. To address these issues, we add two more design constraints:

- No needless HoL-blocking: If the number of queued flows at an egress is less than the number of FIFO queues, the design should have no HoL-blocking. This applies recursively, so if a flow is paused, but its upstream has fewer queued flows than the number of FIFO queues, it should also have no HoL-blocking.

- Stable buffer occupancy: We should aim to store the minimum number of packets needed to keep downstream links busy, consistent with scheduling objectives. In particular, buffer occupancy should not scale to a large multiple of the number of queued flows.

Providing a design that achieves these constraints, along with an evaluation of how this general approach affects performance, are the principal contributions of this work.

3.3 BFC Overview

Table 1 outlines the main components of our solution, called Backpressure Flow Control (BFC).

To fix the problems with the straw proposal, we need to dynamically assign flows to unallocated physical queues, respecting packet order so that subsequent packets for that flow will be placed into the same FIFO queue. Likewise, when we transmit the last packet of a flow, we need to reclaim the physical queue and return it to the unallocated set. Each flow has a unique 5-tuple of the source and destination address and port numbers; we can call this the virtual flow ID (VFID). Our implementation needs to do a lookup on the VFID on every arriving packet, and so we actually store the VFID as a hash of the 5-tuple. We also use this hash to signal per-flow pauses between switches, and so we configure all switches in the network to compute the VFID in the same way. We also need to be careful to handle collisions in the VFID space correctly. For simplicity, it is easiest to think of the VFID as the 5-tuple.

We have state per VFID: which physical queue it has been assigned, whether it has been paused, and how many packets are queued. Unlike ATM switches that implement credit-based per-flow flow control [19], BFC only has state for a flow while the VFID has a queued packet; the state is reclaimed when the last packet of the flow leaves the switch. During normal operation, even a heavily used switch will only have a few hundred active VFIDs, that is, flows with queued packets. A switch under incast can have more, and in the limit we could have as many active VFIDs as buffer slots. We discuss the VFID storage and lookup problem in more detail later in this section (§3.8).

Since there are a limited number of physical queues, it is possible that all physical queues have been allocated at a particular egress when a new flow arrives. In that case, HoL-blocking is unavoidable. In our implementation, we assign the flow to a random physical queue, but we could also be more sophisticated, for example, by assigning the flow to the shortest queue or by avoiding those physical queues that have already been paused. Note that once a flow has been assigned to a queue, all packets from that flow must be assigned to the same physical queue until all of the packets from that flow have left the switch.

In BFC, pauses are sent on each ingress link as a multistage bloom filter [12] representing all paused (and implicitly all unpaused) flows on a particular ingress link. The bloom filter is sent periodically, at a rate of one every half (per-hop) round trip delay. These pauses do not need to be synchronized across inputs. If a pause packet is dropped due to link corruption, the subsequent retransmission, half a round trip delay later, is idempotent. We remove pauses by keeping an internal counting bloom filter at the downstream switch, with a small integer per bit in the bloom filter. If two paused VFIDs map to the same bloom filter bit position, the count will be two, so that when the time comes to unpause one of the VFIDs, the bit transmitted upstream will still reflect the other VFID. A VFID is paused at the upstream if and only if the bloom filter matches on all four hash bit positions.

Note that BFC is not a per-hop credit-based flow control scheme [19]. Flows start in the unpaused state, and can send at full line rate until a pause is received; in practice, of course, the flow’s packets may be interleaved with other traffic traversing the same link, and so will generally send less quickly. When a resume is received (because the VFID no longer matches in the bloom filter), the upstream can again send at the full line rate until another pause is received.

Unlike PFC where a pause blocks all traffic on a link, in normal operation a pause only affects a single flow. Thus,
instead of delaying pauses as long as possible hoping to avoid them, we pause aggressively to preserve space in the congested switch for other flows to transit the switch unaffected. We pause whenever a packet arrives that, in expectation, will not exit the switch before we can unpausing and send the next packet – taking into account the delay to the next pause frame, its arrival at the upstream switch, and the transmission time for the subsequent packets from that flow to arrive at the downstream switch.

With deficit round robin, how long a queued packet will take to drain depends on a number of factors, including potentially the future arrival of packets for other flows. We estimate this as the number of active (unpaused) physical queues at the egress port times the queue depth of the physical queue (including packets from other VFIDs assigned to the same physical queue). As an estimate, it could be wrong. We compute whether a flow should remain paused each time we dequeue one of its packets at the downstream switch. We describe the pause/unpause computation more precisely below.

Note that when we pause a flow at the upstream switch, it blocks all flows sharing the same queue from proceeding (HoL-blocking). Likewise, at the downstream switch, a physical queue can be shared by a mixture of paused and unpaused VFIDs. Pausing one VFID does not prevent other flows from continuing to fill the downstream buffer, although generally those flows will also be paused when their packets are added if the queue is sufficiently large.

The packet scheduler uses deficit round robin to implement fair queuing, among the physical queues that are not paused. Of course, when multiple VFIDs are assigned to a physical queue, we do not guarantee fair queuing among flows.

### 3.4 Pause Threshold

The goal for pausing in BFC is to maintain the minimum amount of buffering without physical queues running out of packets due to pausing. We first describe the simpler case of how BFC computes the pause/resume threshold when each flow queued at the egress port has its own physical queue. We then describe the more general case in the next subsection.

Because of communication delay, any pause/resume for a virtual flow will take $HRTT$, the round trip delay for a pause message to reach the upstream and for its effect to return downstream. Moreover, since the pause/resume are only sent periodically (every $\tau$ s), the pause/resume can take an additional time $\tau$ to take effect. To account for this feedback delay, physical queues in BFC need some amount of buffering before pausing so that they do not run out of packets. More formally, a physical queue builds $(HRTT + \tau)$ of buffering (at the rate the physical queue is being dequeued) before pausing the virtual flow associated with the physical queue. $N_{active}$ is the number of active physical queues (that is, unpaused, with data to transmit). Then, the pause threshold $Th$ (for a particular physical queue) is $(HRTT + \tau) \cdot (\mu/N_{active})$, where $\mu$ is the capacity of the link. This assumes the scheduling policy is deficit round robin; a different calculation would be needed for other scheduling policies. Of course, this is just an estimate, as the actual delay depends on future packet arrivals and the depth of the other queues.

Because the pause takes $HRTT + \tau$ to take effect, one might think we should pause earlier than $Th$, to account for packets in flight or those that will be sent before the pause can take effect. If switch buffering was highly constrained, for example, we could monitor the rate at which data is arriving for a particular flow to estimate the precise moment that we should pause. Or, instead of a blanket unpausing, we could give the upstream switch a specific number of credits to use for filling the downstream buffer. Instead, we simply acknowledge that pauses take some time to take effect, and therefore the amount of buffering consumed can exceed $Th$, in the worst case, by $(HRTT + \tau) \cdot \mu$.

When a physical queue receives a pause from a downstream switch, it immediately pauses, but it does not necessarily transmit the pause to its upstream in turn. Instead, it uses its local value of $N_{active}$ to compute the desired buffer length it would need if it were not paused, and forwards a pause upstream when (or if) a packet is received that exceeds that threshold. This way, when the queue is eventually resumed, there is sufficient buffering in place to keep the link busy until its upstream can be resumed in turn. Note that we compute a flow’s pauses/resumes only on packet arrival/departure for that flow. We do not recalculate the set of paused flows every time $N_{active}$ changes.

### 3.5 Resuming Multiple Flows

We next discuss how pausing and resuming works when multiple flows are assigned to the same physical queue. It is important to keep in mind that while the different flows sharing a physical queue use the same egress port, they may come from different ingress ports.

A strawman approach would be to pause as before, whenever a flow’s packet arrives if the physical queue is too large. When resuming, we could unpausing all $n$ flows associated with the physical queue if the queue length drops below the threshold. The problem is that each of the $n$ newly unpaused flows could potentially send at line rate (if they originated on different ingress ports). This could easily exhaust the downstream switch buffers if applied across every congested egress queue, requiring a blanket pause to avoid packet loss.

Since this extra buffering is unneeded, we adopt the heuristic that we resume at most two flows per round trip (per physical queue). When a flow is to be resumed, we temporarily put it on a “toberesumed” list. Every bloom filter interval (half of $HRTT$), we take one flow off the “toberesumed” list and clear its information from the bloom filter. The choice to resume two flows per round trip helps in the scenario where a resumed virtual flow does not have packets at the upstream.
3.6 Communicating Pauses

We next turn to the mechanics of communicating and implementing pauses. Instead of a bloom filter, we could explicitly send each VFID to signal the pause/resume to the upstream switch. Since we only pause (and therefore resume) a flow at most once per packet arrival, with batching this would bound the link overhead to a reasonable level. We did not take this approach, however, as it would require communicating the pauses reliably, e.g., with a mechanism to detect and retransmit the pause packet if it was corrupted by the transmission link.

Instead, we communicate pauses periodically using a small, idempotent multistage bloom filter [12]. Multistage bloom filters are ideal for communicating sparse bit maps and are convenient for processing updates in parallel. In our evaluation we use 4 hash functions for the bloom filter, and a bloom filter size of 128 bytes. The contents of the bloom filter are specific to a given ingress. For a 100 Gbps link and a frequency of half of one hop round trip, this implies a signalling overhead of about 1%. We note that this overhead will decrease with future increases in link speed.

Of course, a bloom filter can have false positives; in our setting, this can mean needlessly pausing a virtual flow even though the downstream switch would permit it to send. With no HoL-blocking, there are at most 32 queued flows to be paused per ingress. With 4 hash functions, the likelihood of a false positive is then 1 in 5 million. The false positive rate goes up with incast workloads and high levels of congestion; however, in these settings, there is likely to be other flows with packets to transmit in its place.

When the upstream switch receives a bloom filter, it cannot easily invert it into a set of paused VFIDs. Instead, we take the set of physical queues and check the first packet against the bloom filter; if there is a match, we pause the physical queue. This implicitly pauses all other packets in the same physical queue. Equally, if the first packet is not paused, the second packet might be if it is for a different VFID. Thus, we recompute the match against the bloom filter after every packet send.

A detail is that the host network interface card (NIC) must be engineered to accept bloom filter pause frames, to avoid potentially overfilling the first level switch. We assume the NIC has sufficient hardware to maintain a physical queue per VFID, simplifying the logic.

3.7 Reducing HoL Blocking for Short Flows

Most traces of data center communication have found that a majority of flows are a single packet or less. If we assign the packets for these singleton packets to an already occupied physical queue, they will experience HoL-blocking and poor tail latency behavior. Instead, we observe that singleton packets can be directly scheduled. We maintain a special FIFO queue per egress port to hold the first packet from each flow. One can think of this as pre-scheduling this packet, in priority ahead of the other physical queues for that egress. We call this the “high priority queue.” We only put a packet into the high priority queue if its VFID is not paused and there are no other packets queued for that flow. We never pause the high priority queue. This means that a downstream switch may receive one additional packet for a given newly paused flow than it would otherwise. Additional packets that arrive for a flow are assigned to a regular physical queue, and scheduled normally.

In our prototype, we implement a simplification of this idea. The sender NIC marks the first packet of every flow, and this signals the switch to put that packet into the high-priority queue. Unmarked packets are assigned as usual.

3.8 Bookkeeping

When a packet arrives at a switch, we need to look up its VFID to locate its state: whether the flow has any packets queued, which physical queue is has been assigned, and whether it has been paused. Since it happens on every packet, this lookup needs to be fast. An array might work, but a switch is unlikely to have more than a few hundred active flows at any one time, while the VFID space is much larger to avoid collisions. Thus, a linear array would be sparse and inefficient.

Instead, we use a hash table with a bucket size of four; we avoid needing to store the VFID key by setting the number of buckets equal to the number of VFIDs. In addition to the VFID state, an entry in the hash table also stores the ingress and egress for the flow. This means that we can disambiguate VFID collisions that have different inputs and outputs; two 5-tuples that hash to the same VFID and that share the same input and output are treated (by that switch) as if they are the same flow. We choose the VFID space to be large enough that these collisions are rare. In our simulations, the hash table took 256 KB, or 2% of the simulated packet buffer space available in the switch.

A consequence of using a hash table with buckets is that the buckets may fill from time to time. We maintain a small (100 entry) associative cache for these overflow VFIDs. If the overflow cache overflows, we put the packets for these flows into a special per-egress overflow queue which is scheduled along with the other queues in the normal manner. This is not perfectly order preserving within a flow, as the hash table bucket may become free at some point in the future, (rarely) leading to packets in both the overflow queue and a regular physical queue.

3.9 Discussion

Deadlocks: Pushback mechanisms like PFC have been shown to be vulnerable to deadlocks in the presence of cyclic buffer dependencies among switches [16]. BFC is also vulnerable to such deadlocks. Deadlock can be avoided provided receiving hosts drain arriving packets (they cannot pause their upstream switch, even when overloaded), and routes are loop-free (such as the up-down routes typical of data centers or the deadlock-free routing techniques in Bolt [32]).

Guaranteed losslessness: Most data centers that support RDMA use PFC to ensure losslessness, because RDMA NIC
hardware implements only very simple loss recovery. BFC does not guarantee losslessness. In particular, a switch in BFC pauses a virtual flow only after receiving a packet for it from the upstream. This implies a new flow can send at least one packet to the bottleneck switch even if the switch is congested. In certain mass incast scenarios, this might be sufficient to trigger drops. In our prototype, we use FFC to address this situation, but in our evaluation it was never triggered.

4 EVALUATION

In this section, we evaluate the performance of BFC and compare it against several other schemes. We use the NS3 simulator [1] and our implementation of BFC is based upon the DCQCN simulator in [40].

4.1 Setup

Network Topology: For our experiments we consider two fat tree topologies. T1 has 128 leaf servers, 8 top of the rack (ToR) switches and 8 Spine switches (2:1 over subscription). Each Spine switch is connected to all the ToR switches, each ToR has 16 servers, and each server is connected to a single ToR. The smaller topology T2 has 64 leaf servers, 4 Tor (16 servers per ToR) and 8 Spine switches, with an over subscription of 2:1. For both topologies, all links are 100 Gbps with a propagation delay of 1 us. The maximum end-to-end base round trip time (RTT) is 8 us and the 1-Hop RTT is 2 us. The switch buffer size is set to 12 MB. Relative to the ToR switch capacity of 2.4 Tbps, the ratio of buffer size to switch capacity is similar to that of Broadcom’s Tomahawk3 from Fig. 1. We use Go-Back-N for retransmission and an MTU of 1KB. Our experiments use the shared buffer memory model, common in existing switches [9].

Comparison Schemes: DCQCN: DCQCN uses ECN bits and end-to-end control to manage buffer use at the congested switch. We use the author’s implementation in NS3. The PFC threshold is set to trigger when traffic from an input port occupies more than 11% of the free buffer. The ECN marking threshold triggers before PFC ($K_{min} = 100KB$ and $K_{max} = 400KB$). We use the same PFC thresholds for BFC, HPCC and other variants of DCQCN.

DCQCN+Win: We also simulate a variant of DCQCN described in [22], with a per-flow transmission cap of one end-to-end RTT at the full link rate. This cap limits the maximum number of inflight packets for a flow, reducing buffer occupancy without hurting performance.

DCQCN+Win+SFQ: We also combine DCQCN+Win with better switch scheduling. In this scheme, each egress port does stochastic fair queuing (SFQ) on incoming flows. Matching BFC, we use 32 physical queues per port [7].

HPCC: HPCC is a proposed end-to-end control algorithm that uses explicit link utilization information to reduce buffer occupancy and PFCs at the congested switch. We use the author’s implementation of HPCC in the NS-3 simulator, with parameters from the paper, $\eta = 0.95$ and $maxStage = 5$. Since the author’s implementation was only released 10 days before the deadline, we couldn’t run the SFQ variant of HPCC. We plan to include the variant for the camera ready.

Ideal-FQ: To understand how close BFC comes to optimal performance, we simulate ideal fair queuing with infinite buffer on switches. In our implementation of Ideal-FQ, egress ports do SFQ with a very large number of physical queues per port (1000 per port). The NICs use a window cap of 1 BDP. Note that Ideal-FQ is not realizable in practice; its role is to bound how well we might possibly do.

Performance metrics: We consider four performance metrics. 1) Tail latency, measured as the 99th percentile flow completion time (FCT) normalized to the best possible FCT for the same size flow, running at link rate (this is referred to as the FCT slowdown); 2) Utilization of the entire network; 3) The overall buffer occupancy at the switch; 4) Percentage of time links were paused due to PFC.

Workloads: We synthesized a trace to match the flow size distributions from the three industry workloads discussed in Fig. 3. 1) Aggregated workload from all applications in a Google data center; 2) Hadoop cluster at Facebook (FB_Hadoop); 3) Web search. The flow arrival pattern follows a lognormal distribution with $\sigma = 2$. We consider scenarios with and without incast. For consistency of presentation, we only report the FCT slowdown for the normal (non-incast) traffic.

4.2 Performance

Fig. 5 shows our principal result, the 99th percentile tail latency FCT slowdown across different flow sizes for the Google and FB_Hadoop workload on the T1 topology. In all the experiments, the average load is set to 65% of the network capacity. For Fig. 5a and 5b, the average load includes a 5% 100-to-1 incast traffic. The incast size is 20MB in aggregate. Fig. 6 shows the buffer occupancy and the percentage of time links were paused due to PFC in Fig. 5a.

Out of all the schemes, DCQCN is worst for tail latency across all flow sizes. Adding a window cap (DCQCN+Win) limits the inflight packets of a flow, reducing buffer size and PFC pauses, and thus, improving tail latency. The window cap alone is not enough, however, and the performance is still far from optimal (Ideal-FQ). The problem is that DCQCN is slow in responding to congestion. Since flows start at line rate, a flow can build up an entire end-to-end bandwidth-delay product (BDP) of buffering (100 KB) at the bottleneck before there is any possibility of reducing its rate. If the switch is scheduled in FIFO order (DCQCN+Win), this leads to long delays particularly at the tail. Excessive PFC pauses make this even worse. The problem is aggravated during incast events, as this increases the number of active flows at the switch. The bottle-
physical queues. However, the FCT slowdowns (especially for short flows) are still 8-10 × worse than optimal, because of the impact of (i) collisions in assigning flows to physical queues and (ii) buffer exhaustion and PFC pauses that cause upstream HoL-blocking. DCQCN+Win+SFQ still adjusts flow rates in an end-to-end manner. As a result, PFCs are triggered at a similar rate (Fig. 6b).

HPCC has been proposed as an improvement to DCQCN, by using link utilization information instead of ECN as a feedback mechanism, as well as an improved control algorithm. It is still an end-to-end protocol, however. Compared to both DCQCN and DCQCN+Win, HPCC reduces tail latency, tail buffer occupancy, and PFC pauses. For short flows, compared to BFC, the FCT’s are still 8-10 × worse with incast, and 3-4 × without. While HPCC’s control algorithm reacts more quickly than DCQCN, it is still limited by its end-to-end architecture, especially during incast.

Ideal-FQ achieves the lowest tail latencies among all the schemes. It avoids HoL-blocking at the congested switch by scheduling equally among all queued flows, and it uses (infinite) buffering to avoid PFC triggered HoL-blocking at upstream switches. However, buffer occupancy can grow to an unfeasible level.

BFC achieves the best tail latency among all realizable schemes (close to optimal). Without incast, BFC performance tracks optimal quite closely. With incast, there are enough incoming flows to exhaust the number of physical queues. For single packet flows, BFC avoids contention for physical queues even with incast by putting those packets into a separate, high-priority queue. This deviates from fair queuing, and so the distinction with Ideal-FQ in this case is not meaningful. Multi-packet flows, however, must contend for scarce physical queues and therefore see somewhat elevated tail latency relative to perfect fair queuing. This effect is largest for the smallest flows. However, it is important to note that tail latency with BFC is still 3-15 × better than existing schemes. Physical queues are used more efficiently than with stochastic fair queueing, and buffers are managed to prevent PFCs and cascading HoL-blocking.

Interestingly, in the absence of incast (Fig. 5c), PFC was never triggered for DCQCN+Win+SFQ and HPCC. Even so, their FCTs are still worse than BFC. BFC’s improvement in this case can be attributed to its efficient use of physical queues and reduced buffer occupancy. BFC’s performance improvement is not a mere consequence of reduction in PFC pauses, and so replacing PFC with smarter retransmission strategies would not, by itself, close the gap with BFC.

Compared to BFC and Ideal-FQ, tail latency for medium flows (100-500KB) is slightly better with some variants of DCQCN. Since DCQCN only slows down long flows, medium flows can get through more quickly. This significantly hurts tail latency for large flows (Fig. 5c).

**Physical queue assignment:** To understand the importance of dynamically assigning flows to physical queues, we
repeated the experiment in Fig. 5a with our straw proposal from § 3.2 (referred as BFC-VFID). Recall that BFC-VFID uses hashing to statically assign flows to physical queues (as in SFQ), with pauses propagated to the corresponding physical queue at upstream switches. In BFC, the physical queue assignment is dynamic. In our experiment, BFC-VFID also uses the high priority queue for very short flows, and to isolate the effect of changing the physical queue assignment, the pause thresholds are same as BFC.

Fig. 7a shows the tail latency. Compared to BFC, tail latency for BFC-VFID is much worse for all flow durations. Without the VFID indirection, flows are often hashed to the same physical queue, triggering HoL-blocking and sometimes upstream pauses, even when there are unoccupied physical queues. Fig. 7b is the CDF of such collisions. BFC-VFID experiences collisions in a high fraction of cases (20%). In contrast, BFC has collisions 1% of the time. Even with incast, the number of active flows in BFC is smaller than the physical queues almost all of the time.

Contention for physical queues can also result in unfairness. This is especially bad for short flows as they may be assigned a physical queue with a flow that already has filled a significant amount of the buffer. To better understand this effect, we implemented SFQ+InfBuffer, where we statically assign flows to physical queues, but with a buffer size set to infinity so that there are no drops or pauses. Flows have a window size cap of one end-to-end bandwidth-delay product. This is also shown in Fig. 7. The tail latency for very short flows (< 1KB) is worse in SFQ+InfBuffer as it doesn’t use a high priority queue. For the remaining flows, SFQ+InfBuffer performs better than BFC-VFID, as it avoids upstream HoL-blocking due to pauses. However, it remains far worse for tail latency than BFC for most size flows.

**Impact of incast on utilization:** With existing systems, incast events can build up large buffers at the switches, triggering PFCs (or BFC pauses) and degrading utilization. To test this, in the T2 topology, we create 4 long-lived flows for each receiver from 4 random senders. We introduce a periodic incast of aggregate size 20 MB every 500 us. We vary the fan in (number of senders) of the incast from 10 to 800. Fig. 8 shows the utilization and the tail 99th percentile switch buffer occupancy.

As the fan in increases, the size of each incast flow becomes smaller. For DCQCN, this reduces its ability to control the incast traffic, triggering PFCs, and decreasing utilization. The utilization drops to 70% at a fan in of 200, where the size of an incast flow is 100 KB (∼1 end-to-end BDP). BFC achieves higher utilization (close to 100%) and lower buffer occupancy compared to DCQCN+Win across all scenarios. BFC does incur a drop in utilization at very high fan in, when it runs out of physical queues and flows start sharing and triggering HoL-blocking, reducing utilization.

From here on, the remainder of the experiments use the smaller T2 topology.

**Cross datacenter environments:** To understand the impact of BFC on managing cross-data center congestion in a metro
area, we created a topology with two T2 data centers, with 10 Gbps links (for computation speed) and 9 MB switch buffers. Relative to the 100 Gbps topology in earlier graphs, we would expect that BFC would have a smaller benefit on a 10 Gbps topology because more of the bytes are sent in flows that take multiple round trips and thus are more amenable to end-to-end control. Two gateway switches connect the data centers using a 100 Gbps link with 200 us of one way delay (i.e. the base round trip delay of the link is 400 us). A longer hop will require more buffering in BFC to keep the link busy, and so we configured the buffer size of the gateway switch to be 60 MB. The experiment consists of both intra- and inter-data center flows derived from the FB_Hadoop distribution. The aggregate load (including both intra- and inter-data center flows) is 65%. 20% of all the flows are inter-data center.

Fig. 9 shows the tail latency in FCT slowdown for intra- and inter-data center flows for BFC and DCQCN+Win. BFC is better for both types of flows. In particular, the slowdown for BFC is close to 1 (ideal) for inter-data center flows. In contrast, tail latency for DCQCN+Win is 2.5× higher, implying that CTs in absolute terms are 600 us higher. Because the end-to-end RTT is so much higher, DCQCN will be slow in reacting to congestion for any traffic that crosses multiple data centers. An inter-data center flow can have up to (∼500KB) of buffering at the switches (both gateway and internal data center switches). This can lead to deep packet buffers and PFC pauses, degrading tail latency even for intra-data center traffic.

In contrast, BFC reacts at the scale of the hop-by-hop RTT. Even though inter-data center flows have higher end-to-end RTTs, on switches within the data center, BFC will pause/resume flows on a hop-by-hop RTT timescale (2 us). As a result, with BFC, tail latencies of intra-data center flows are unaffected by the presence of inter-data center flows, while the opposite is true of DCQCN. We confirmed this experimentally in a separate experiment where we removed all the inter-data center traffic. The tail latency of DCQCN improved in this case, while it had no effect on BFC. With BFC, inter-data center flows build buffers at the gateway switches where they are needed to keep the utilization high, and not at internal switches where the extra buffering disrupts other traffic.

4.3 Design choices

Buffer occupancy management: In §3.5, we introduced an optimization to limit the buffering at a physical queue when multiple flows are mapped to it. To test the effectiveness of this optimization, we created concurrent long lived flows destined to the same receiver. We varied the number of concurrent flows from 8 to 256. This implies that the average number of flows per physical queue varied from 0.25 to 8 (32 physical queues per port). We also ran a variant of our scheme without the buffer optimization (BFC-BufferOpt). In BFC-BufferOpt, an egress port resumes all virtual flows associated with a physical queue when the physical queue goes below the pause threshold.

Fig. 10 shows the tail (99th percentile) per physical queue buffer occupancy at the switch. By limiting the rate we resume flows per physical queue (two per 1-Hop RTT), BFC keeps worst case buffer utilization to about twice the per-hop BDP, regardless of the number of concurrent flows sharing a particular physical queue. Note that this result does not apply for all workloads; a flood consisting entirely of one or two packet flows can drive buffer occupancy up, even for BFC.

In contrast, without this optimization, BFC-BufferOpt unpauases more aggressively and fails to limit the size of the physical queue. In fact, the buffer size grows linearly with the number of concurrent flows.

High priority queue: Using the high priority queue for single packet flows has two key advantages. First, it gives priority to the large number of messages that fit within a single packet. Since this is the typical use case for RDMA, that is particularly appropriate. In addition, since these short flows no longer occupy physical queues, these very short flows are not trapped behind other flows. This is especially advantageous in workloads where most flows are very short (in the Google workload more than 80% flows are < 1KB). Further, longer flows are more likely to find an empty physical queue, improving the ability of the scheduler to optimize flow completion time for other competing flows.

We ran BFC with and without this optimization (BFC-HighPriorityQ). We use the Google workload with 85% average load + 5% additional traffic from 100-to-1 incast. Fig. 11 shows the cumulative distribution of the number of physical queues in use at the switch as well as the tail latency. High load is likely to make the impact of high priority queue particularly pronounced. The high priority queue improves FCT for singleton flows significantly. More interestingly, it also improves tail latency for small and moderate-sized flows, by reducing the number of collisions for physical queues.
4.4 Sensitivity to resource constraints

Number of physical queues: To measure the impact of changing the number of physical queues per egress port, we ran BFC on 60% average load from the Google workload + 5% incast (same as Fig. 5a) as we vary the number of physical queues per port from 8 to 128. Fig. 12 shows the fraction of flows that share a same physical queue (collisions) and the tail latency in each scenario. The figure also shows the tail latency for Ideal-FQ as a point of comparison. As expected, decreasing the number of physical queues increases collisions and hurts tail latency. For this workload, the knee of the curve is at 32 queues, while 64 queues shows modest further improvement to the point of diminishing returns.

Note that with 64 queues, the tail latency for multi-packet flows is better than with Ideal-FQ. This is due to the fact that with incast, the true optimal behavior is to defer sending the incast traffic to the bottleneck switch. Ideal-FQ is greedy, sending as much data downstream as quickly as possible. However, this means that incast packets destined to stall in downstream buffers compete for bandwidth with other traffic that can be more quickly delivered out the downstream switch. Slowing down the cross-traffic hurts tail latency in this scenario. By contrast, BFC allows some incast traffic to go forward, but holds the remainder in upstream buffers, possibly all the way back to the source, until it is really needed.

Number of VFIDs: We repeated the experiment from Fig. 12, but varied the number of VFIDs instead. Fig. 13a shows HoL blocking due to flows getting mapped to the same entry in the hash table (collisions) and the fraction of overflows. Fig. 13b shows the tail latency across different scenarios. On this workload, BFC’s performance is not sensitive to the number of VFIDs. Tail latencies are similar even when the number of VFIDs is reduced to 1024. This might allow the switch to use a smaller hash table, but we note that other workloads, such as those with higher utilization and smaller flows, may be more sensitive to the hash table size.

Size of bloom filter: To test the sensitivity of the results to the size of the bloom filter, we ran the experiment from Fig. 12 and varied the size of bloom filter from 16B to 128B. Fig. 14 shows the tail latencies in each scenario. BFC’s performance is largely unaffected, reflecting the fact that on this workload there are relatively few paused flows. False positive matches in the bloom filter cause unnecessary pauses of upstream queues, and this has a greater relative effect on tail latency than average latency, and for short flows, the tail latency increases by $1.5 \times$ when the bloom filter is reduced to 16 Bytes.

5 RELATED WORK

ATM era schemes: Our work was inspired by work in the early 90s on hop-by-hop credit-based flow control for managing high speed switches in ATM networks [6,19]. ATM networks have small fixed-sized cells and per-connection state at every switch. Credit-based flow control was also introduced by multiprocessor hardware designs of the same era [8,18,21]; the technique is still in wide use in that setting. In these systems, each switch methodically tracks its buffer space, granting permission to upstream switches to forward traffic if and only if there is room. The result is a network that has no congestion loss by design. Deadlock can be prevented in a variety of ways, but is typically through per-destination queues. BFC differs from the earlier ATM approaches by using on/off pause frames rather than numerical credits, and it does not assume per-connection state. As a result, BFC is not lossless.

Sender-based congestion control for RDMA: Data center network operators have developed and deployed several end-to-end congestion control protocols for RDMA networks. These differ in the feedback information available to the algorithms (TIMELY [26] uses delay variation, DCQCN uses ECN marks [40], and HPCC uses link load [22]). For these systems, congestion information can be out of date by the time it arrives back at the sender, and so careful control theory design is needed for stability, with multiple parameters that must be carefully tuned to the workload and physical network. By contrast, because it operates only over a single hop, BFC’s control loop is shorter and simpler to configure, and can therefore support higher utilization for a given target tail latency.

Selective retransmission: We can remove the need for all-or-nothing PFC pause frames for RDMA by changing the standard to support selective retransmission in hardware, as suggested by IRN [27]. Handling arbitrary out-of-order
would schedule packets. We considered using AFQ but could Approximate Fair Queuing (AFQ) [30] assigns packets to worse flow completion tail latency than HPCC [22]. This could be used to control congestion but faces similar virtual-to-physical queue mapping accomplishes the same goal but avoids head-of-line blocking when the number of active flows is less than the number of physical queues. Measurements of production data center networks suggest that congestion can occur throughout the network [31]. In earlier work on preventing denial-of-service attacks [35, 37], routers themselves give credits to the sender (via the receiver). This could be used to control congestion but faces similar control theory challenges to other end to end schemes.

Receiver-driven congestion control: Because sender-based congestion control systems generally perform poorly on incast workloads, some researchers have proposed shifting to a scheme where the receiver prevents congestion by explicitly allocating credits to senders for sending traffic. Three examples are NDP [15], pHost [13] and HOMA [28]. Unlike BFC, these schemes only address congestion that occurs at the last hop, assuming congestion in the network is negligible. Measurements of production data center networks suggest that congestion can occur throughout the network [31]. In earlier work on preventing denial-of-service attacks [35, 37], routers themselves give credits to the sender (via the receiver). This could be used to control congestion but faces similar control theory challenges to other end to end schemes.

Switch scheduling: Several efforts have looked at improving switch scheduling to reduce average flow completion times. Stochastic fair queueing [23] hashes flows to a smaller number of FIFO queues to approximate fair queueing; our virtual-to-physical queue mapping accomplishes the same goal but avoids head-of-line blocking when the number of active flows is less than the number of physical queues. Approximate Fair Queuing (AFQ) [30] assigns packets to priority queues to approximate the order that fair queueing would schedule packets. We considered using AFQ but could not devise a simple way to pause a flow once its packets had been assigned to a priority level. pFabric [4] labels packets at the sender with the amount of remaining data in the flow, allowing the switch to assign packets to priority queues to implement shortest remaining flow first. We think of pFabric as complementary; it is future work to study how best to integrate BFC with switch scheduling algorithms like pFabric.

6 CONCLUSION
In this paper, we presented Backpressure Flow Control (BFC), a new architecture for per-flow per-hop flow control designed to reduce RDMA tail latency for typical workloads found in modern data center networks. In BFC, switches dynamically assign flows to physical queues, allowing fair scheduling among competing flows as long as the number of queued flows is no larger than the number of physical queues. BFC uses selective backpressure on a per-flow basis to manage the buffering at each switch to avoid packet loss and minimize upstream HOL-blocking. On synthetic workloads derived from measurements of data centers, BFC dramatically outperforms existing RDMA congestion control mechanisms, both with and without the presence of competing incast traffic. We further show that our approach generalizes to yield performance improvements for cross-data center RDMA communication.

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