Nezha: Deployable and High-Performance Consensus Using Synchronized Clocks

[Technical Report]

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ABSTRACT

This paper presents a high-performance consensus protocol, Nezha, designed for single-cloud-region environments, which can be deployed by cloud tenants without any support from their cloud provider. Nezha bridges the gap between protocols such as Multi-Paxos and Raft, which can be readily deployed and protocols such as NOPaxos and Speculative Paxos, that provide better performance, but require access to technologies such as programmable switches and in-network prioritization, which cloud tenants do not have.

Nezha uses a new multicast primitive called deadline-ordered multicast (DOM). DOM uses high-accuracy software clock synchronization to synchronize sender and receiver clocks. Senders tags messages with deadlines in synchronized time; receivers process messages in deadline order, on or after their deadline.

We compare Nezha with Multi-Paxos, Fast Paxos, Raft, a NOPaxos version we optimized for the cloud, and 2 recent protocols, Domino and TOQ-based EPaxos, that use synchronized clocks. In throughput, Nezha outperforms all baselines by a median of 5.4× (range: 1.9–20.9×). In latency, Nezha outperforms five baselines by a median of 2.3× (range: 1.3–4.0×), with one exception: it sacrifices 33% of latency performance compared with our optimized NOPaxos in one test. We also prototype two applications, a key-value store and a fair-access stock exchange, on top of Nezha to show that Nezha only modestly reduces their performance relative to an unreplicated system. Nezha is available at https://gitlab.com/steamgjk/nezhav2.

1 INTRODUCTION

Our goal in this paper is to build a high-performance consensus protocol for a local-area network such as a single cloud zone, which can be deployed by cloud tenants with no help from their cloud provider. We are motivated by the fact that the cloud hosts a number of applications that need both high performance (i.e., low latency and high throughput) and fault tolerance. We provide both current and futuristic examples motivating our work below.

First, modern databases (e.g., Cosmos DB, TiKV and CockroachDB) would like to provide high throughput and strong consistency (linearizability) over all their data. Yet, they often need to split their data into multiple instances because a single instance's throughput is limited by the consensus protocol [11, 44, 52]. Second, microsecond-scale applications are pushing the limits of computing [3, 23, 26, 31, 34]. Such applications often have stateful components that must be made fault-tolerant (e.g., the matching engine within a fair-access cloud stock exchange [17], details in §10). To effectively support such applications on the public cloud, we need the consensus protocol to provide low latency and high throughput.

Despite significant improvements in consensus protocols over the years, the status quo falls short in 2 ways. First, protocols such as Multi-Paxos [29] and Raft [49] can be (and are) widely deployed without help from the cloud provider. However, they only provide modest performance: latency in the millisecond range and throughput in the 10K requests/second range [14]. Second, high-performance alternatives such as NOPaxos [33], Speculative Paxos [53], NetChain [25], NetPaxos [25], and Mu [3], require technologies such as programmable switches, switch multicast, RDMA, priority scheduling, and control over routing—most of which are out of reach for the cloud tenant.[1]

Here, we develop a protocol, Nezha, that provides high performance for cloud tenants without access to such technologies. Our starting point in designing Nezha is to observe that a common approach to improve consensus protocols is through *optimism*: in an optimistic protocol, there is a common-case fast path that provides low client latency, and there is a fallback slow path that suffers from a higher latency. Examples of this approach include Fast Paxos [28], EPaxos [47], Speculative Paxos [53], and NOPaxos [33].

For optimism to succeed, however, the fast path must indeed be the common case, i.e., the fraction of client requests that take the fast path should be high. For a sequence of client requests to take the fast path, these requests must arrive in the same order at all servers involved in the consensus protocol. In the public cloud, however, cloud tenants have no control over paths from clients to these servers. As we empirically demonstrate in §3, this leads to frequent cases of *reordering*: client requests arrive at servers in different orders. Thus, for an optimistic protocol to improve performance in the public cloud, reordering must be reduced. This observation influenced the design of Nezha, which has 3 key ideas.

Deadline-ordered multicast. Nezha uses a new network primitive, called deadline-ordered multicast (DOM), designed to reduce the rate of reordering in the public cloud. DOM is a multicast that works as follows. The sender's and receivers' clocks are synchronized to each other to produce a global time shared by the sender and all receivers. The sender attaches a deadline in global time to its message and multicasts the message to all its receivers. Receivers process a message on or after its deadline, and process multiple messages in increasing order of deadline. Because the deadline is a message property and common across all receivers of a message, ordering by deadline provides the same order of processing at all receivers and undoes the reordering effect. DOM is best effort: messages arriving after their deadlines or lost messages are no longer DOM's responsibility. Thus, for DOM to be effective, the deadline should be set so that most messages arrive before

[1] We note that many of these technologies are available to *cloud providers*, but in most cases they are not exposed to tenants of the cloud. RDMA instances [45] are an exception, but such instances are expensive.

their deadlines—despite variable network delays and despite clock synchronization errors. However, if messages arrive after their deadlines, correctness is still maintained because Nezha falls back to the slow path. Here, DOM follows Liskov's suggestion of using accurate clock synchronization for performance improvements, but not as a necessity for correctness [35].

Speculative execution. DOM combats reordering and increases the fraction of client requests that take the fast path. Our next idea reduces client latency of Nezha in the slow path, by decoupling execution of a request from committing the request. Consensus protocols like Multi-Paxos and Raft wait until the request is committed at a quorum of servers before executing the request at one of them (typically the leader). However, the leader in Nezha executes the request before it is committed and sends the execution result to the client. The client then accepts the leader's execution result only if it also gets a quorum of replies from other servers that indicate commitment; otherwise, the client just retries the request. Thus a leader's execution is *speculative* in that the execution result might not actually be accepted by a client because (1) the leader was deposed after sending its execution result and (2) the new leader executed a different request instead.

Proxy for deployability. Performing quorum checks, multicasting, and clock synchronization at the client creates additional overhead on a Nezha client relative to a typical client of a protocol like Multi-Paxos or Raft. This overhead arises because the client is now doing additional work per client request relative to a typical consensus client. To address this, Nezha uses a proxy (or a fleet of proxies if higher throughput is needed), which multicasts requests, checks the quorum sizes, and performs clock synchronization—on the client's behalf. Because Nezha's proxy is stateless, it is easy to scale with the number of clients and it is easier to make fault tolerant.

Evaluation. We compare Nezha to six baselines in public cloud: Multi-Paxos, Fast Paxos, our optimized version of NOPaxos, Raft, Domino and TOQ-based EPaxos under closed-loop and open-loop workloads. In a closed-loop workload, commonly used in the literature [33, 43, 47, 53], a client only sends a new request after receiving the reply for the previous one. In open-loop workloads, recently suggested as a more realistic benchmark [60], clients submit new requests according to a Poisson process, without waiting for replies for previous ones. We find:

(1) In closed-loop tests, Nezha (with proxies) outperforms all the baselines by 1.9–20.9× in throughput, and by 1.3–4.0× in latency at close to their saturation throughputs.

(2) In open-loop tests, Nezha (with proxies) outperforms all the baselines by 2.5–9.0× in throughput. It also outperform five baselines by 1.3–3.8× at close to their saturation throughputs. The only exception is that, it sacrifices 33% of latency compared with our optimized version of NOPaxos.

(3) Nezha can achieve better latency without a proxy, if clients perform multicasts and quorum checks. In open-loop tests, Nezha (without proxies) outperforms all the baselines by 1.3–6.5× in latency at close to their respective saturation throughputs. In closed-loop tests, Nezha (without proxies) outperforms them by 1.5–6.1×.

(4) We also use Nezha to replicate two applications (Redis and a prototype financial exchange) and show that Nezha can provide fault tolerance with only a modest performance degradation: compared with the unreplicated system, Nezha sacrifices 5.9% throughput for Redis; it saturates the processing capacity of CloudEx and prolongs the order processing latency by 4.7%.

Nezha is open-sourced at https://gitlab.com/steamgjk/nezhav2.

2 BACKGROUND

2.1 Clock Synchronization

In a distributed system, each node (server or VM) may report a different time due to clock frequency variations. Clock synchronization aims to bring clocks close to each other, by periodically correcting each node's current clock offset and/or frequency. Given two nodes i and j, their clock times are denoted as $c_i(t)$ and $c_j(t)$ at a certain real time t. We consider their clocks synchronized if

$$|c_i(t) - c_j(t)| \leq \epsilon$$

where ϵ is the clock error bound, indicating how closely the clocks are synchronized at t.

Guaranteeing an ϵ is difficult because clock synchronization may fail occasionally, and the error bound can grow arbitrarily in such cases. Therefore, Nezha does not depend on clock synchronization for correctness. However, well synchronized clocks can improve the performance of Nezha by increasing the fraction of requests that can be committed in the fast path of Nezha using the DOM primitive. While Nezha is compatible with any clock synchronization algorithm, in our implementation, we chose to build Nezha on Huygens [16], because Huygens is a high-accuracy software-based system that can be easily deployed in the public cloud.

2.2 Consensus Protocols

In this section, we overview consensus protocols most closely related to Nezha, namely, Multi-Paxos [29]/Raft [49], Fast Paxos [28], Speculative Paxos [53], and NOPaxos [33]. Table 1 summarizes the basic properties of them and Nezha. §11 provides a more detailed comparison to related work.

Deployable but low-performance. Multi-Paxos/Raft and Fast Paxos are generally deployable, but have lower performance. It takes 4 message delays (i.e., 2 RTTs) for Multi-Paxos/Raft to commit one request and the leader can become a throughput bottleneck. Fast Paxos can save 1 message delay if the request is committed in the fast path. However, when there is no in-network functionality that increases the frequency of the fast path (e.g., the MOM primitive [53] or the OUM primitive [33]), Fast Paxos performs much worse (§9.2) because most requests can only be committed in the slow path, but its slow path (2.5 RTTs) is even slower than Multi-Paxos/Raft and causes heavier bottleneck for its leader.

High-performance but needs in-network functionality. Speculative Paxos [53] and NOPaxos [33] can achieve high performance. However, both require considerable in-network functionality. Speculative Paxos requires that most requests arrive at replicas in the same order to commit them in 1 RTT. Speculative Paxos achieves this by (1) controlling routing to ensure the same path length for client-to-replica requests and (2) using in-network priorities to ensure that these requests encounter low queues. When reordering occurs, the request can only be committed via the

Table 1: Basic Properties of Typical Consensus Protocols; F:fast path; S:slow path.

Protocols	Message Delays	Load on Leader [1]	Quorum Size	Quorum Check	Inconsistency Penalty	Deployment Requirement
Multi-Paxos/Raft	4	$2(2f+1)$	$f+1$	Leader	Low	No special requirement
Fast Paxos	F:3 S:5	$2f+2$	F: $f+\lceil f/2 \rceil+1$ S: $f+1$	F/S: Leader	Medium	No special requirement
Speculative Paxos	F:2 S:6	2	F: $f+\lceil f/2 \rceil+1$ S: $f+1$	F: Client S: Leader	High (rollback)	Priority scheduling, SDN control, etc
NOPaxos	F: 2 or 3 [2] S: 4 or 5	2	F: $f+1$ S: $f+1$	F: Client S: Leader	Medium	Priority scheduling, programmable switch [3], SDN control, etc
Nezha (No proxy)	F: 2 S: 3	$2+2f/m$ [4]	F: $f+\lceil f/2 \rceil+1$ S: $f+1$	F/S: Client	Low	Clock synchronization
(With proxy)	F: 4 S: 5			F/S: Proxy		

[1] Load on Leader indicates how many messages the leader processes per client request.

[2] When NOPaxos uses a switch-based sequencer, the message flow client→sequencer→replica incurs one message delay because the switch is on path; hence, the overall latency is 2 message delays. With a software sequencer, both client→sequencer and sequencer→replica incur one message delay; hence, the overall latency is 3 message delays. The same holds for the slow path.

[3] Programmable switch serves as the hardware sequencer, and it is unnecessary when NOPaxos uses software sequencer. However, software sequencers can reduce throughput

[4] 2 of the messages are the request and the reply. The other $2f$ messages are *sync* messages with much smaller size, and can be batched, so the load is amortized by a factor of m.

slow path (3 RTTs); the slow path also requires application-specific rollback. Speculative Paxos is very sensitive to packet reordering. Its throughput drops by 10× with a 1% reordering rate [53][Figure 9]; such rates can easily occur in public cloud, where routing is out of control and packets can go different paths. We include a micro-benchmark in §3, which shows the reordering rate can usually be more than 20%. NOPaxos requires a programmable switch as the sequencer to achieve its optimal latency (1 RTT). When such a switch is unavailable, NOPaxos uses a server as a software sequencer, which adds 1 additional message delay to its fast path. Besides, as we show (§9.2), NOPaxos also loses throughput when using a software sequencer in public cloud. In particular, it is not resistant to bursts in our open-loop tests, which further increases packet reordering/drop and trigger its slow path, causing distinct degradation.

3 MOTIVATION

Consensus protocols are often used to provide the abstraction of a replicated state machine (RSM) [57], where multiple servers cooperate to present a fault-tolerant service to clients. In the RSM setting, the goal of consensus protocols is to get multiple servers to reach agreement on the contents of an ordered log, which represents a sequence of operations issued to the RSM. This amounts to 2 requirements, one for the order of the log and one for the contents of the log. We state these 2 requirements below.

For any two replicas R_1 and R_2:

- **Consistent ordering.** If R_1 processes request a before request b, then R_2 should also process request a before request b, if R_2 received both a and b.
- **Set equality.** If R_1 processes request a, then R_2 also processes request a.

Many *optimistic* protocols leverage the fact that the ordering of messages from clients to replicas is usually consistent at different locations: they employ a fast path during times of consistent ordering and fall back to a slow path when ordering is not consistent [28, 33, 53, 69]. However, for an optimistic protocol to actually improve performance, the fast path should indeed be the common case. If not, such protocols can potentially hurt performance [28, 53] relative to a protocol that doesn't optimize for the common case like Raft or Multi-Paxos.

Consistent ordering is violated if messages arrive in different orders at different receivers. This situation is especially common in the public cloud where there is frequent reordering: messages from one or more senders to different receivers take different network paths and arrive in different orders at the receivers.

We measure reordering rate with a simple experiment on Google Cloud. We use two receiver VMs, denoted as R_1 and R_2. We use a variable number of sender VMs to multicast messages to R_1 and R_2. We vary the rate of a Poisson process used by each sender to generate multicast messages (Figure 1) or vary the number of multicasting senders (Figure 2). After the experiment, R_1 receives a sequence of messages, which serves as the ground truth: each message is assigned a sequence number based on its arrival order at R_1. We use these sequence numbers to check how reordered R_2 is. We count a message received by R_2 as reordered if (1) its sequence number is smaller than the sequence number of R_2's previous message, or (2) the message does not exist in R_1's sequence (because it was dropped at R_1). Figure 1 shows that when we vary the submission rate, keeping the number of senders fixed at 2, the reordering rate quickly exceeds 20%. Further in Figure 2, when we vary the number of senders, keeping the submission rate fixed at 10K messages/second, the reordering rate increases rapidly up to 36% with the number of senders.

In the public cloud, with such high reordering rates, optimistic protocols are forced to take the slow path often, which reduces their performance (§9.2). In order to design a high-performance protocol, we need to reduce the rate of reordering. This motivates us to design the *deadline-ordered multicast* (DOM) primitive to guarantee consistent ordering among replicas. DOM does not guarantee set equality. This is intentional and is also why we need a consensus

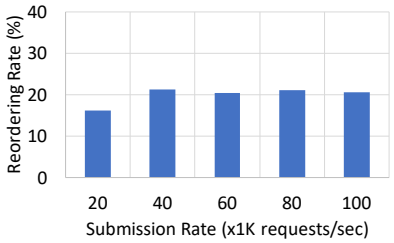


Figure 1: Packet reordering vs. submission rate on Google Cloud

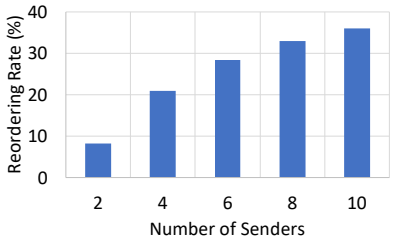


Figure 2: Packet reordering vs. number of senders on Google Cloud

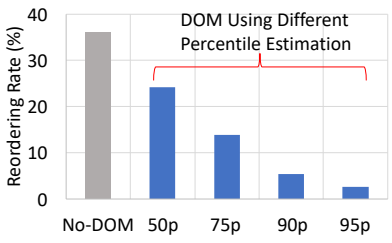


Figure 3: Effectiveness of DOM on packet reordering on Google Cloud

protocol, Nezha, to go along with DOM because guaranteeing both requirements has been shown to be as hard as consensus itself [7].

4 DEADLINE-ORDERED MULTICAST

Informally, deadline-Ordered Multicast (DOM) is designed to reduce the rate of reordering by (1) waiting to process a message at a receiver until the message's deadline has passed and (2) delivering messages to the receiver in deadline order. This gives other messages with a lower deadline the ability to "catch up" and reach the receiver before a message with a later deadline is processed.

Formally, in DOM, a sender wishes to send a message M to multiple receivers R_1, R_2, ..., R_n. The sender attaches a deadline $D(M)$ to the message, where $D(M)$ is specified in a global time that is shared by senders and receivers because their clocks are synchronized. Then the DOM primitive attempts to deliver M to receivers within $D(M)$. Receivers (1) can only process M on or after $D(M)$ and (2) must process messages in the order $D(M)$ regardless of M's sender.

We stress that DOM is a *best-effort* primitive: a sequence of messages is processed in order at a receiver *if* they all arrive before their deadline, but DOM does not guarantee that messages arrive *reliably* at all receivers either before the deadline or ever. There are two situations that cause DOM messages to arrive late or be lost.

The first is network variability: messages may not reach some receivers or reach them so late that the other messages with larger deadlines have been processed. The second is a temporary loss of clock synchronization. If clocks are poorly synchronized, the deadline on a message might be much earlier in time than the actual time at which the receiver receives the message.

While DOM is a general primitive, we comment briefly on its specific use for consensus as in Nezha. When DOM is used for consensus, because DOM makes no guarantees on late or lost messages, it is up to the slow path of the consensus protocol to handle such messages. If client requests are lost because of drops in the network and haven't been received by a quorum of replicas, it is up to clients to retry the requests. These weaker guarantees in DOM are important because providing both reliable delivery and ordering of multicast messages is just as hard as solving consensus [7]. The use of clock synchronization for performance (i.e., increasing the frequency of the fast path) rather than correctness (i.e., linearizability) is also in line with Liskov's suggestion on how synchronized clocks should be used [35].

Setting DOM deadlines. Setting deadlines is a trade-off between avoiding message reordering and adding too much waiting time. In the public cloud, where VM-to-VM latencies can be variable and reordering is common, these deadlines should be set adaptively based on recent measurements of one-way delays (OWDs), which are also enabled by clock synchronization. We pick the deadline for a message by taking the maximum among the estimated OWDs from all receivers and adding it to the sending time of the message. The estimation of OWD is formalized as below.

$$\widetilde{OWD} = \begin{cases} P + \beta(\sigma_S + \sigma_R), & 0 < \widetilde{OWD} < D \\ D \end{cases}$$

To track the varying OWDs, each receiver maintains a sliding window for each sender, and records the OWD samples by subtracting the message's sending time from its receiving time. Then the receiver picks a percentile value from the samples in the window as P. We previously tried moving average but found that just a few outliers (i.e. the tail latency samples) can inflate the estimated value. Therefore, we use percentiles for robust estimation. The percentile is a DOM parameter set by the user of DOM.

Besides P, DOM also obtains from the clock synchronization algorithm [16] the standard deviation for the sending time and receiving time, denoted as σ_S and σ_R [2]. σ_S and σ_R provide an *approximate* error bound for the synchronized clock time, so we add the error bound with a factor β to P and obtain the final estimated OWD. The involvement of $\beta(\sigma_S+\sigma_R)$ enables an adaptive increase of the estimated value, leading to a graceful degradation of Nezha as the clock synchronization performs worse. Moreover, in case that clock synchronization goes wrong and provides invalid OWD values (i.e. very large or even negative OWDs), we further adopt a clamping operation: If the estimated OWD goes out of a predefined scope $[0, D]$, we will use D as the estimated OWD. The estimated OWD will be replied to the sender to decide the deadlines of subsequent requests.

To illustrate DOM's benefits, we redo our experiments from §3 with 10 Poisson senders, each submitting 10K requests/sec to 2 receivers. Figure 3 shows different percentiles (i.e., 50^{th}, 75^{th}, 90^{th}, and 95^{th}) for DOM to decide its deadlines. We can see that a higher percentile leads to more reduction of reordering. However, a higher percentile also means a longer holding delay for messages in DOM, which in turn undermines the latency benefit of Nezha protocol.

[2] σ_S and σ_R are calculated based on the method in Appendix A of [15].

5 NEZHA OVERVIEW

We use DOM as a building block to develop a consensus protocol, called Nezha, atop DOM. We overview the protocol here and describe it in detail in subsequent sections. Recall that DOM maintains consistent ordering across replicas by ordering messages based on their deadlines. This allows Nezha to use a fast path that assumes consistent ordering across replicas. When DOM fails to deliver a message to enough replicas before the message's deadline (either because of delays or drops), Nezha uses a slow path instead.

Model and assumptions. Nezha assumes a fail-stop model and does not handle Byzantine failures. It uses $2f + 1$ replicas: 1 leader and $2f$ followers, where at most f can be faulty and crash. Nezha guarantees safety (linearizability) at all times and liveness under the same assumptions as Multi-Paxos/Raft. However, Nezha's performance is improved by DOM, whose effectiveness depends on accurate clock synchronization among VMs and the variance of OWDs between proxies and replicas. Here "accurate" means the clocks among proxies and replicas are synchronized with a small error bound *in most cases*. But Nezha does not assume the existence of a worst-case clock error bound that is never violated, because clock synchronization can also fail [35, 37, 38].

Nezha architecture. Nezha uses a stateless proxy/proxies (Figures 4 and 5) interposed between clients and replicas to relieve clients of the computational burden of quorum checks and multicasts. Using a stateless proxy also makes Nezha a drop-in replacement for Raft/Multi-Paxos because the client just communicates with a Nezha proxy like it would with a Raft leader. This proxy serves as the DOM sender, while the replicas serve as DOM receivers. The DOM deadline is set to the maximum of a sliding window median (50^{th} percentile) of OWD estimates between the proxy and each replica; these deadlines also take into account the current estimate of clock synchronization errors (§4). Another benefit of a proxy is that it is sufficient if the proxy's clock is synchronized with the receivers; the client can remain unsynchronized.

Fast/Slow path sketch. We very briefly describe Nezha's fast path and slow path, leaving details to later sections. Figure 4 shows the fast path. The request is multicast from the proxy ①. If the request's deadline is larger than the last request released from the *early-buffer*, the request enters the *early-buffer* ②. It will be released from the *early-buffer*s at the deadline, so that replicas can append the request to their *log*s ③. The *log* list is ordered by request deadline. After that, followers immediately send a reply to the proxy without executing the request ⑤, whereas the leader first executes the request ④ and sends a reply including the execution result. The proxy considers the request as committed after receiving replies from the leader and $f + \lceil f/2 \rceil$ followers. The proxy also obtains the execution result from the leader's reply, and then replies with the execution result to its client. The fast path requires a super quorum ($f + \lceil f/2 \rceil + 1$) rather than a simple quorum ($f + 1$) for the same reason as Fast Paxos [28]: Without leader-follower communication, a simple quorum cannot persist sufficient information for a new leader to always distinguish committed requests from uncommitted requests (details in §6.3).

Figure 5 shows the more involved slow path: when a multicasted ① request goes to the *late-buffer* because of its small deadline ②, followers do not handle it. However, the leader must pick it out of its *late-buffer* eventually for liveness. So the leader modifies the request's deadline to make it eligible to enter the *early-buffer* ③. After releasing and appending this request to the log ④, the leader broadcasts this request's identifier (a 3-tuple consisting of *client-id*, *request-id*, and request deadline) to followers ⑦, to force followers to keep consistent logs with the leader. On hearing this broadcast ⑧, the followers add/modify entries from their log to stay consistent with the leader: as an optimization, followers can retrieve missing requests from their *late-buffer*s without having to ask the leader for these entries ⑨. After this, followers send replies to the proxy ⑩. Meanwhile, the leader has executed the request ⑤ and replied to the proxy ⑥. After collecting $f + 1$ replies (including the leader's reply), the proxy considers the request as committed. Notably, Nezha differs from the other optimistic protocols (e.g. [28, 33, 53]): it also decouples the request execution and quorum check in the slow path. Such a decoupling design enables a *faster* slow path for the proxy to commit requests in the slow path. Besides, through the quorum check, the proxy can ensure that the speculative execution result from the leader replica is safe to use.

6 THE NEZHA PROTOCOL

6.1 Replica State

Figure 6 summarizes the state variables maintained by each replica. We omit some variables (e.g., *crash-vector*) related to Nezha's recovery (§7). Below we describe them in detail.

replica-id: Each replica is assigned with a unique *replica-id*, ranging from 0 to $2f$. The *replica-id* is provided to the replica during the initial launch of the replica process, and is then persisted to stable storage, so that the replica can get its *replica-id* after crash and relaunch.

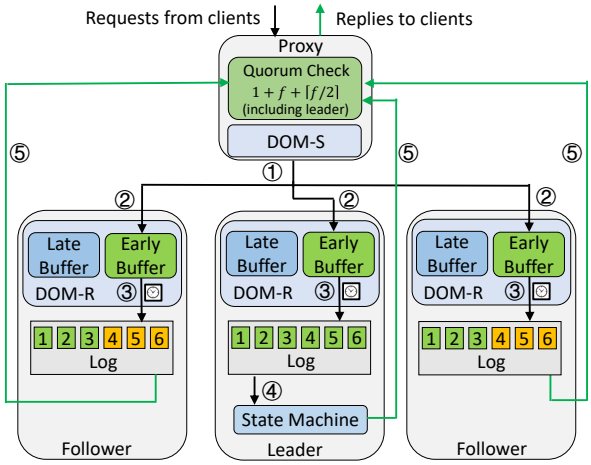
view-id: Replicas leverage a view-based approach [36]: each view is indicated by a *view-id*, which is initialized to 0 and incremented by one after every view change. Given a *view-id*, this view's leader's *replica-id* is *view-id*$\%(2f + 1)$.

status: Replicas switch between three different *status*es. Replicas are initially launched in NORMAL status. When the leader is suspected of failure, followers switch from NORMAL to VIEWCHANGE and initiate the view change process. They will switch back to NORMAL after completing the view change. For a failed replica to rejoin the system, it starts from RECOVERING status and will switch to NORMAL after recovering its state from the other replicas.

early-buffer: *early-buffer* is implemented as a priority queue, sorted by requests' deadlines. *early-buffer* is responsible for (1) conducting eligibility checks of incoming requests: if the incoming request's deadline is larger than the last released one from *early-buffer*, then the incoming request can enter the *early-buffer*; and (2) release its accepted requests in their deadlines' order, thus maintaining DOM's consistent ordering across replicas.

late-buffer: *late-buffer* is implemented as a map using the *<client-id, request-id>* as the key. It is used to hold those requests which are not eligible to enter the *early-buffer*. Replicas maintain such a buffer because those requests may later be needed in the slow path (§6.4). In that case, replicas can directly fetch those requests locally instead of asking remote replicas.

log: Requests released from the *early-buffer* will be appended to the *log* of replicas. The requests then become the entries in the *log*. The *log* is ordered by request deadline.



① The proxy broadcasts the request to all replicas via DOM-S, and the request is tagged with the sending time s and the latency bound l, summing up to be $deadline$.

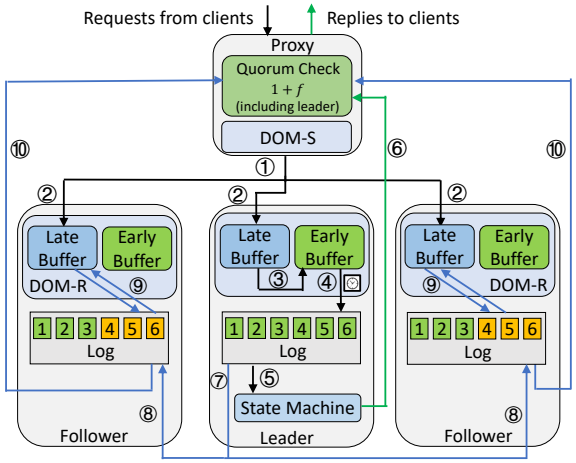
② The request goes into Early Buffer if its $deadline$ is larger than the last appended entry in Log (i.e. the last released entry from Early Buffer).

③ Replicas (DOM-Rs) release the request when the clock time has passed its $deadline$, and appends it to Log.

④ The leader replica executes the request.

⑤ Replicas send *fast-replies* (including a hash value of Log) . The proxy considers the request as committed after receiving 1 leader's and $f + \lfloor f/2 \rfloor$ followers' replies with the same hash. Then the proxy acks a reply to the client.

Figure 4: Fast path of Nezha



① Same as ① in the fast path.

② The request goes into Late Buffer if its $deadline$ is smaller than that of the last released request from Early Buffer.

③ The leader sets the request's $deadline$ slightly larger than the last released request, and puts it into Early Buffer.

④ ⑤ ⑥ are the same as ③ ④ ⑤ in the fast path.

⑦ (In parallel with ⑤) The leader broadcasts the indices of log entries to followers.

⑧ Followers refer to the indices and add/delete entries to/from Log to keep consistent with the leader.

⑨ When missing entries, the follower first tries to recover it from Late Buffer. If it is still missing, then the follower fetches it from others.

⑩ Followers send *slow-replies*. The proxy considers the request as committed after receiving 1 leader's *fast-reply* and f followers' *slow-replies*. Then the proxy replies to client.

Figure 5: Slow path of Nezha

- *replica-id*— replica identifier $(0, 1, \cdots, 2f)$.
- *view-id*— the view identifier, initialized as 0 and increased by 1 after every view change.
- *status*— one of NORMAL, VIEWCHANGE, or RECOVERING.
- *early-buffer*— the priority queue provided by DOM, which sorts and releases the requests according to their deadlines.
- *late-buffer*– the map provided by DOM, which is searchable by <*client-id, request-id*> of the request.
- *log*— a list of requests, which are appended in the order of their deadlines.
- *sync-point*— the log position indicating this replica's *log* is consistent with the leader up to this point.
- *commit-point*— the log position indicating the replica has checkpointed the state up to this point.

Figure 6: Local state of Nezha replicas

sync-point: Followers modify their *log*s to keep consistent with the leader (§6.4). *sync-point* indicates the log position up to which this replica's *log* is consistent with the leader. Specially, the leader always advances its *sync-point* after appending a request.

commit-point: Requests (log entries) up to *commit-point* are considered as committed/stable, so that every replica can execute requests up to *commit-point* and checkpoint its state up to this position. *commit-point* is used in an optional optimization (§8.3).

6.2 Message Formats

There are five types of messages closely related to Nezha. We explain their formats below. Since Nezha uses a view-based approach for leader change, we omit the description of messages related to leadership changes; these messages have been defined in Viewstamped Replication [36].

request: *request* is generated by the client and submitted to the proxy. The proxy will attach some necessary attributes and then submit *request* to replicas. *request* is represented as a 5-tuple:

request=<client-id, request-id, command, s, l>

client-id represents the client identifier and *request-id* is assigned by the client to uniquely identify its own request. On one replica, *client-id* and *request-id* combine to uniquely identify the request. *command* represents the content of the request, which will be executed by the leader. s and l are tagged by proxies. s is the sending time of the *request* and l is the estimated latency bound. When the request arrives at the replica, the replica can derive the request's deadline as $s + l$. Meanwhile, the replica can also derive the proxy-replica OWD by subtracting s from its receiving time.

fast-reply: *fast-reply* is sent by every replica after they have appended or executed the request, and it is used for quorum checks in the fast path. *fast-reply* is represented as a 6-tuple:

fast-reply = <*view-id*, *replica-id*, *client-id*, *request-id*, *result*, *hash*>

view-id and *replica-id* are from the replica state variables (see §6.1). *client-id* and *request-id* are from the appended request that lead to this reply. *result* is only valid in the leader's *fast-reply*, and is *null* in followers' *fast-replies*. *hash* captures a hash of the replica's *log* (the hash calculation is explained in §8.1). Proxies can check the *hash* values to know whether the related replicas have consistent *log*s.

log-modification: *log-modification* message is broadcast by the leader to convey its *log* identifier (*deadline*+*client-id*+*request-id*) to followers, making the followers modify their *log*s to keep consistent with the leader. Meanwhile, *log-modification* also doubles as the leader's heartbeat. *log-modification* is represented as a 5-tuple:

log-modification=<*view-id*, *log-id*, *client-id*, *request-id*, *deadline*>

view-id is from the replica state. *log-id* indicates the position of this log entry (request) in the leader's *log*. *client-id* and *request-id* uniquely identify the request on each replica. *deadline* is the request's deadline shown in the leader's *log*, which is either assigned by proxies on the fast path or overwritten by the leader on the slow path (i.e., ③ in Figure 5). *log-modification* messages can be batched under high throughput to reduce the leader's burden of broadcast.

slow-reply: *slow-reply* is sent by followers after all the entries in their *log*s have become consistent with the leader's *log* up to this request. It is used by the client to establish the quorum in the slow path. *slow-reply* is represented as a 4-tuple:

slow-reply = <*view-id*, *replica-id*, *client-id*, *request-id*>

The four fields have the same meaning as in the *fast-reply*.

log-status: *log-status* is periodically sent from the followers to the leader, reporting the *sync-point* of the follower's *log*, so that the leader can know which requests have been committed and update its *commit-point*. *log-status* is represented as a 3-tuple.

log-status=<*view-id*, *replica-id*, *sync-point*>

The three fields come from the followers' replica state variables.

6.3 Fast Path

Nezha relies on DOM to increase the frequency of its fast path. As shown earlier, the percentile at which DOM estimates OWDs is a parameter set by the DOM user. A lower percentile will set smaller deadlines, which improve fast path latency, but reduce the frequency of the fast path. Higher percentiles have the opposite problem. For Nezha, we use the 50^{th} percentile to strike a balance between the two. This does reduce the fast path frequency compared with using a higher percentile; hence, Nezha compensates for this by optimizing its slow path for low client latency as well.

To commit the request in the fast path (Figure 4), the proxy needs to get the *fast-reply* messages from both the leader and $f + \lceil f/2 \rceil$ followers. (1) It must include the leader's *fast-reply* because only the leader's reply contains the execution result. (2) It also requires the $f + \lceil f/2 \rceil + 1$ replicas have matching *view-id*s and the same *log* (requests). In §8.1 we will show how to efficiently conduct the quorum check by using the *hash* field included in *fast-reply*. If both (1) and (2) are satisfied, the proxy can commit the request in 1 RTT.

As briefly explained in the sketch of the fast path (§5), the fast path requires a super quorum ($f + \lceil f/2 \rceil + 1$) rather than a simple quorum ($f+1$), because a simple quorum is insufficient to guarantee the correctness of Nezha's fast path. Consider what would happen if we had used a simple majority ($f + 1$) in the fast path. Suppose there are two requests *request-1* and *request-2*, and *request-1* has a larger *deadline*. *request-1* is accepted by the leader and f followers. They send *fast-replies* to the proxy, and then the proxy considers *request-1* as committed and delivers the execution result to the client application. Meanwhile, *request-2* is accepted by the other f followers. After that, the leader fails, leaving f followers with *request-1* accepted and the other f followers with *request-2* accepted. Now, the new leader cannot tell which of *request-1* or *request-2* is committed. If the new leader adds *request-2* into the recovered *log*, it will be appended and executed ahead of *request-1* due to *request-2*'s smaller *deadline*. This violates linearizability [21]: the client sees *request-1* executed before *request-2* with the old leader and sees the reverse with the new leader.

6.4 Slow Path

The proxy is not always able to establish a super quorum to commit the request in the fast path. When requests are dropped or are placed into the *late-buffers* on some replicas, there will not be sufficient replicas sending fast replies. Thus, we need the slow path to resolve the inconsistency among replicas and commit the request. We explain the details of the slow path (Figure 5) below in temporal order starting with the request arriving at the leader.

Leader processes request. After the leader receives a *request*, the leader ensures it can enter the *early-buffer*: if it is not eligible due to its small deadline ②, the leader will modify its deadline to make it eligible ③. The leader then conducts the same operations as in the fast path (i.e., appending the request ④, applying it to the state machine ⑤, and sending *fast-reply* ⑥). The leader also broadcasts the *log-modification* message ⑦ in parallel with ⑤-⑥.

Leader broadcasts log-modification. Every time the leader appends a request to its *log*, it broadcasts a *log-modification* message to followers ⑦. Every time a follower receives a *log-modification* message ⑧, it checks its log entry at the position *log-id* included in the *log-modification* message. (1) If the entry has the same 3-tuple <*client-id*, *request-id*, *deadline*> as that included in the *log-modification* message, it means the follower has the same log entry as the leader at this position. (2) If only the 2-tuple <*client-id*, *request-id*> is matched with that in the *log-modification* message, it means the leader has modified the deadline, so the follower also needs

to replace the deadline in its entry with the deadline from the *log-modification* message. (3) Otherwise, the entry has different *<client-id, request-id>*, which means the follower has placed a wrong entry at this position. In that case, the follower removes the wrong entry and tries to put the right one. It first searches its *late-buffer* for the right entry with matching *<client-id, request-id>*. As a rare case, when the entry does not exist on this replica because the request was dropped or delayed, the follower fetches it from other replicas and puts it at the position.

Follower sends slow-reply. After the follower has processed the *log-modification* message, and has ensured the requests in its *log* are consistent with the leader, the follower updates its *sync-point*, indicating its *log* is consistent with the leader up to the log position indicated by the *sync-point*. The leader itself can directly advance its *sync-point* after appending the request to *log*. Then, the follower sends a *slow-reply* message for every synced request ⑩. The *slow-reply* will be used to establish the quorum in the slow path. Specially, a *slow-reply* can be used in place of the same follower's *fast-reply* in the fast path's super quorum, because it indicates the follower's *log* is consistent with the leader. By contrast, the follower's *fast-reply* cannot replace its *slow-reply* for the quorum check in the slow path.

Proxy conducts quorum check. The proxy considers the request as committed when it receives the related *fast-reply* from the leader and the *slow-replies* from f followers. The execution result is still obtained from the leader's *fast-reply*. Our decoupling design enables the proxy to know whether the request is committed even earlier than the leader replica. Meanwhile, replicas can continue to process subsequent requests and are *not blocked by the quorum check in the slow path*, which proves to be an advantage compared to other opportunistic protocols like NOPaxos (see §9.2). Unlike the quorum check of the fast path (§6.3), the slow path does not need a super quorum ($1 + f + \lceil f/2 \rceil$). This is because, before sending *slow-replies*, the followers have updated their *sync-point*s and ensured that all the requests (log entries) are consistent with the leader up to the *sync-point*s. A simple majority ($f + 1$) is sufficient for the *sync-point* to survive the crash. All requests before *sync-point* are committed requests, whose log positions have all been fixed. During the recovery (§7), they are directly copied to the new leader's *log*.

In the background: followers report sync-statuses. In response to *log-modification* messages, followers send back *log-status* messages to the leader to report their *sync-point*s. The leader can know which requests have been committed by collecting the *sync-point*s from $f+1$ replicas including itself: the requests up to the smallest *sync-point* among the $f+1$ ones are definitely committed. Therefore, the leader can update is *commit-point* and checkpoints its state at the *commit-point*. It can also broadcast the *commit-point* to followers, which enables them to checkpoint their states for acceleration of recovery (§8.3). Note that the followers' reporting *sync-status* is not on the critical part of the client's latency on the slow path; it happens in the background. Therefore, the slow path only needs three message delays (1.5 RTTs) for the proxy to commit the request.

6.5 Timeout and Retry

The client starts a timer while waiting for the reply from the proxy. If the timeout is triggered (due to packet drop or proxy failure), the client retries the request with the same or different proxy (if the previous proxy is suspected of failure), and the proxy resubmits the request with a different sending time and (possibly) a different latency bound. Meanwhile, as in traditional distributed systems, replicas maintain *at-most-once* semantics. When receiving a request with duplicate *<client-id, request-id>*, the replica simply resends the previous reply instead of appending/executing it twice.

7 RECOVERY

Assumptions. We assume replica processes can fail because of process crashes or a reboot of its server. When a replica process fails, it will be relaunched on the same server. However, we assume that there is some stable storage (e.g., disk) that survives process crashes or server reboots. A more general case, which we do not handle, is to relaunch the replica process from a different server with a new disk where the stable storage assumption no longer holds. We also do not handle the case of changing Nezha's f parameter by adding or removing replicas from the system. Both cases are handled by the literature on reconfigurable consensus [36, 64], which we believe can be adapted to Nezha as well.

Recovery protocol. Nezha's recovery protocol consists of two components: replica rejoin and leader change. After a replica fails, it can only rejoin as a follower. If the failed replica happens to be the leader, then the remaining followers will stop processing requests after failing to receive the leader's heartbeat for a threshold of time. Then, they will initiate a view change to elect a new leader before resuming service. We describe the recovery protocol in pseudo-code in Appendix §A, and include a model-checked TLA+ specification in Appendix §I. We also include the correctness proof in Appendix B. Here, we only sketch the major steps for the new leader to recover its state (*log*).

After the new leader is elected via the view change protocol, it contacts the other f survived replicas, acquiring their *log*s, *sync-point*s and *last-normal-view*s (i.e., the last view in which the replica's status is NORMAL). Then, it recovers the *log* by aggregating the *log*s of those replicas with the largest *last-normal-view*. The aggregation involves two key steps.

(1) The new leader chooses the largest *sync-point* from the qualified replicas (i.e., the replicas with the largest *last-normal-view*). Then the leader directly copies all the *log* entries up to the *sync-point* from that replica.

(2) For the remaining part, if the *log* entry has a larger *deadline* than the *sync-point*, the leader checks whether this entry exists on $\lceil f/2 \rceil + 1$ out of the qualified replicas. If so, the entry will also be added to the leader's *log*. All the entries are sorted by their *deadline*s.

After the leader rebuilds its *log*, it executes the entries in their *deadline* order. It then switches to NORMAL *status*. After that, the leader distributes its rebuilt *log* to followers. Followers replace their original *log*s with the new ones, and also switch to NORMAL.

In some cases, the leader change can happen not only because of a process crash but also because of a network partition, where followers fail to hear from the leader for a long time and start a view change to elect the new leader. When the deposed leader notices the existence of a higher view, it needs to abandon its current state, because its current state may have diverged from the state of the new leader. In other words, the state of the deposed leader may include the execution of some uncommitted requests, which do not

exist in the new view. To maintain correct state, the deposed leader transfers the state from another replica in the fresh view.

Avoiding disk writes during normal processing. While designing the recovery protocol, we aim to avoid disk writes as much as possible. This is because disk writes can add significant delays (0.5ms~20ms per write), significantly increasing client latency. At the same time, we also want to preserve the correctness of our protocol from *stray messages* [27], which proved to cause bugs to multiple diskless protocols (e.g., [36][53][33][69]). Nezha adopts the *crash-vector* technique invented by Michael et al. [41, 42], to develop Nezha's recovery protocol. While Nezha still uses stable storage to distinguish whether it is the first launch or reboot, it does not use disk writes during normal processing and preserves its correctness from the *stray message* effect. In Appendix §A we describe how to use *crash-vector* to prevent *stray-message*s for Nezha.

Correctness. In Appendix §B, we have proved Nezha's three correctness properties. The three properties have also been model-checked in our TLA+ specification (Appendix §I).

- **Durability:** if a client considers a request as committed, the request survives replica crashes.
- **Consistency:** if a client considers a request as committed, the execution result of this request remains unchanged after the replica's crash and recovery.
- **Linearizability:** A request appears to be executed exactly once between start and completion. The definition of linearizability can also be reworded as: if the execution of a request is observed by the issuing client or other clients, no contrary observation can occur afterwards (i.e., it should not appear to revert or be reordered).

8 OPTIMIZATION

8.1 Incremental Hash

In Nezha's fast path, *fast-replies* from replicas can form a super quorum only if these replies indicate that the replicas' ordered logs are identical. This is because—unlike the slow path—replicas do not communicate amongst themselves first before replying to the client. One impractical way to check that the ordered logs are identical is to ship the logs back with the reply. A better approach is to perform a hash over the sequence corresponding to the ordered log, and update the hash every time the log grows. However, if the log is ever modified in place (like we need to in the slow path), such an approach will require the hash to be recomputed from scratch starting from the first log entry.

Instead, we use a more efficient approach by decomposing the equality check of two ordered logs into two components: checking the contents of the 2 logs and checking the order of the 2 logs. Because logs are always ordered by deadline at all our replicas, it suffices for us to check the contents of the 2 logs. The contents of the logs can be checked by checking equality of the 2 sets corresponding to the entries of the 2 logs: this requires only a hash over a set rather than a hash over a sequence.

To compute this hash over a set, we maintain a running hash value for the set. Every time an entry is added or removed from this set, we compute a hash of this entry (using SHA-1) and XOR this hash with the running hash value. This allows us to rapidly update the hash every time a log entry is appended (an addition to the set) or modified (a deletion followed by an addition to the set). The proxy checks for equality of this set hash across all replicas, knowing that equality of the set of log entries guarantees equality of the ordered logs because logs are always ordered by deadlines.

To be more specific, when the replica sends the *fast-reply* for its n^{th} request, the hash represents the set of all previously appended entries using an incremental hash function [5]:

$$H_n = \bigoplus_{1 \le i \le n} h(request_i)$$

Here, $h(*)$ is a standard hash function (we use SHA1) and $\oplus$ is the XOR operation. To calculate $h(request_i)$, we concatenate the values of the request's *deadline*, *client-id*, *request-id* into a bitvector, and then transform it into a hash value.

To avoid *stray message* effect [41, 42], we also XOR H_n with the hash of *crash-vector* to get the final hash value:

$$hash_n = H_n \oplus h(\textit{crash-vector})$$

Here, $h(\textit{crash-vector})$ is calculated by concatenating every integer in the vector and transforming it into a hash value. The inclusion of *crash-vector* is necessary for the correctness of Nezha and we explain this in Appendix §A.1.

The replica includes $hash_n$ while sending the *fast-reply* for the n^{th} request in its *log*. Assuming no hash collisions, $hash_n$ represents the replica state when replicas reply to the proxy. By comparing the *hash* in the *fast-replies* from different replicas, the proxy can check if they have the same requests. Because the hash is computed over the set (but it represents the sequence) of entries, adding/deleting requests only requires incremental computation of XOR and $h(*)$, instead of recomputing from scratch every time.

8.2 Commutativity Optimization

To enable a high fast commit ratio without a long holding delay of DOM, we employ a commutativity optimization in Nezha. As an example, commutative requests refer to those requests operating on different keys in a key-value store, so that the execution order among them does not matter [9, 51]. The commutativity optimization enables us to choose a modest percentile (50^{th} percentile) while still achieving a high fast commit ratio, because it eases the fast path in two aspects.

First, it relaxes the eligibility check condition of the *early-buffer*. Without commutativity, DOM prevents the incoming request from entering the *early-buffer* if its deadline is smaller than the last request released from the *early-buffer* (§4). Otherwise, the consistent ordering property is violated. However, the execution results of commutative requests are not affected by their order [51]. Hence, consistent ordering is only required among non-commutative requests, which enables the relaxation of the *early-buffer*'s entrance check: the request can enter the *early-buffer* if its deadline is larger than the last released request, *which is not commutative with the incoming request.*

Second, it refines the hash computation, making hash consistency among replicas becomes easier. Since read requests do not modify replica state, the *hash* field in the *fast-reply* does not need to encode read requests. Besides, when encoding previous write requests, the *hash* field only considers those that are not commutative to the

current request. To do that, Nezha maintains a table of per-key hashes for the write requests. For every newly appended write request, the replica will XOR its hash to update the corresponding per-key hash in the table according to its key. While sending the *fast-reply* for a specific request, the replica only includes the hash of the same key. For compound requests, which write (and hence do not commute with) multiple keys (e.g., "move 10 from x to y and return x and y"), the replica fetches the hashes of all relevant keys (e.g. x and y), and includes the XORed hash value (e.g. $hash_x \oplus hash_y$) in the *fast-reply*.

We also evaluate across a range of workloads in Appendix §C, with different read/write ratio and skew factors. The result shows that the commutativity optimization helps reduce the latency by 7.7 %-28.9 %.

8.3 Periodic Checkpoints

To (1) accelerate the recovery process after leader failure and (2) enable the deposed leader to quickly catch up with the fresh state, we integrate periodic checkpoints mechanism in Nezha.

Since Nezha only allows the leader to execute requests during normal processing, it can lead to inefficiency during leader change, either caused by leader's failure or network partition. This is because the new leader is elected from followers, and it has to execute all requests from scratch after it becomes the leader. To optimize this, we adopt a similar idea as NOPaxos [33] and conduct synchronization in the background.

Periodically, the followers report their *sync-point*s to the leader, and the leader chooses the smallest *sync-point* among the $f + 1$ replicas as the *commit-point*, and broadcasts the *commit-point* to all replicas. Both the leader and followers checkpoint state at their *commit-point*s. The periodic checkpoints bring acceleration benefit in two aspects: (1) When the leader fails, the new leader only needs to recover and execute the requests from its *commit-point* onwards. (2) When network partition happens, the leader is deposed and it later notices the existence of a higher view. Instead of abandoning its complete state (as what we described in §7), it can start from its latest checkpoint state, and only retrieve from another replica (in the fresh view) the requests beyond its *commit-point*.

9 EVALUATION

We answer the following questions during the evaluation:

(1) How does Nezha compare to the baselines (Multi-Paxos, Fast Paxos, NOPaxos) in the public cloud?

(2) How does Nezha compare to the recent protocols which also use clock synchronization (i.e., Domino and TOQ-based EPaxos)?

(3) How effective are the proxies, especially when there is a large number of replicas?

(4) How fast can Nezha recover from the leader failure?

(5) How does Nezha compare to Raft when both are equipped with log persistence to stable storage?

(6) Does Nezha provide sufficient performance for replicated applications?

9.1 Settings

Testbed. We run experiments in Google Cloud. We employ `n1-standard-4` VMs for clients, `n1-standard-16` VMs for replicas and NOPaxos sequencer, and `n1-standard-32` VMs for Nezha proxies. All VMs are in a single cloud zone. Huygens is installed on all VMs and has an average 99^{th} percentile clock offset of 49.6 ns.

Baselines. We compare with Multi-Paxos, Fast Paxos and NOPaxos. For the 3 baselines, we use the implementation from the NOPaxos repository [32] with necessary modification: (1) we change multicast into multiple unicasts because network-support multicast is unavailable in cloud. (2) we use a software sequencer with multi-threading for NOPaxos because tenant-programmable switches are not yet available in cloud. We also added two recently proposed protocols that leverage synchronized clocks for comparison, i.e., Domino [69] and TOQ-based EPaxos [60]. We choose to compare them with Nezha because they also use clock synchronization to accelerate consensus. For Domino, it is previously tested with clocks synchronized by Network Time Protocol (NTP) in [69], but in our test, we also provide Huygens synchronization for Domino to give it more favorable conditions because Huygens has higher accuracy than NTP [16]. However, since both Domino and TOQ-based EPaxos target WAN settings, we do not expect them to perform better than Nezha or our other baselines in LAN settings, which is verified by our experiments (§9.2). We also intend to compare with Derecho [24]. However, its performance degrades a lot in public cloud (see Appendix §E). We think the comparison is not fair to Derecho and do not include it.

Metrics. We measure execution latency: the time between when a client submits a request to the system and receives an execution result from it along with a confirmation that the request is committed. We also measure throughput. To measure latency, we use median latency because it is more robust to heavy tails. We have attempted to measure tail latency at the 99th and 99.9th percentile. But we find it hard to reliably measure these tails because tail latencies within a cloud zone can exceed a millisecond [22, 46, 66]. This is unlike the WAN setting where tails can be more reliably estimated [60]. We run each experiment 5 times and average values before plotting.

Evaluation method. We follow the method of NOPaxos [33] and run a *null application* with no execution logic. Traditional evaluation of consensus protocols [33, 43, 47, 48, 53, 62] use closed-loop clients, which issue a continuous stream of back-to-back requests, with exactly one outstanding request at all times. However, the recent work [60] suggests a more realistic open-loop test with a Poisson process where the client can have multiple outstanding requests (sometimes in bursts). We use both closed-loop and open-loop tests. While comparing the latency and throughput in §9.2, we use 3 replicas. For the closed-loop test, we increase load by adding more clients until saturation [3]. For the open-loop test, we use 10 clients and increase load by increasing the Poisson rate until saturation.

Workloads. Since the three baselines (Multi-Paxos, Fast Paxos and NOPaxos) are oblivious to the read/write type and commutativity of requests, and the *null application* does not involve any execution logic, we simply measure their latency and throughput under one type of workload, with a read ratio of 50 % and a skew factor [19] of 0.5. We also evaluate Nezha under various read ratios and

[3] Specially, when the system is saturated, the throughput can drop instead of continuously increasing [8, 63], as shown in Figure 7.

skew factors in Appendix §C, which verifies the robustness of its performance.

9.2 Comparison with Multi-Paxos, Fast Paxos and NOPaxos

The closed-loop and open-loop evaluation results are shown in Figure 7. We plot two versions of Nezha. Nezha-Proxy uses standalone proxies whereas Nezha-Non-Proxy lets clients undertake proxies' work. Below we discuss three main takeaways.

First, all baselines yield poorer latency and throughput in public cloud, in comparison with published numbers from highly-engineered networks [33]. Fast Paxos suffers the most and reaches only 4.0K requests/second at 425 µs in open-loop test (not shown in Figure 7b). When clients send at a higher rate, Fast Paxos suffers from heavy reordering, and the reordered requests force Fast Paxos into its slow path, which is even more costly than the Multi-Paxos.

Second, NOPaxos performs unexpectedly poorly in the open-loop test, because it performs *gap handling* and *normal request processing* in one thread. NOPaxos *early binds* the sequential number with the request at the sequencer. When request reordering/drop inevitably happens from the sequencer to replicas, the replicas trigger much gap handling and consume most CPU cycles. We realize this issue and develop an optimized version (NOPaxos-Optim in Figure 7) by using separate threads for the two tasks. NOPaxos-Optim outperforms all the other baselines because it offloads request serialization to the sequencer and quorum check (fast path) to clients. But it still loses significant throughput in the open-loop test compared with the closed-loop test. This is because open-loop tests create more bursts of requests, and cause packet reordering/drop more easily. When 10 open-loop clients submit 10K requests/sec each, NOPaxos replicas trigger gap handling (slow path) for more than 30% of requests. Besides, the gap handling also *blocks the processing of follow-up requests* because NOPaxos still relies on the leader replica to do quorum checks in the slow path. Once a previous request triggers the slow path, all the follow-up requests will be made to wait for at least 1 RTT before the leader completes gap agreement. Thus, all these follow-up requests will count the gap handling cost into their latencies, and they can also continue to cause more gaps.

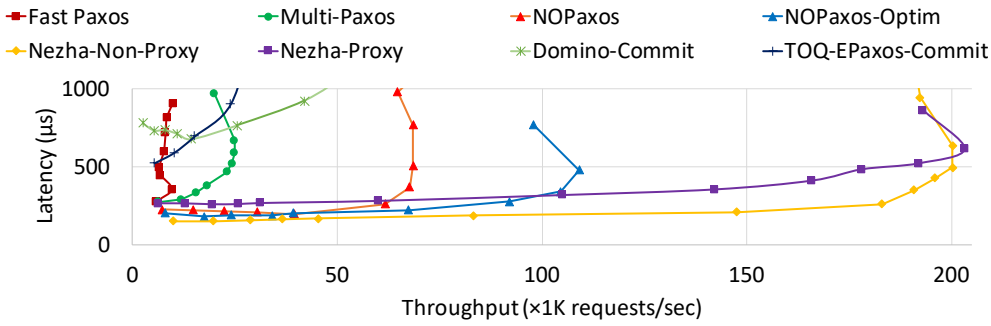
Last, Nezha achieves much higher throughput than all the baselines, and Nezha-Non-Proxy also achieves the lowest latency because of co-locating proxies with clients. Even equipped with standalone proxies, Nezha-Proxy still outperforms all baselines at their saturation throughputs, except NOPaxos-Optim (open-loop). Nezha's improved throughput and latency come from three design aspects: (1) DOM helps create consistent ordering for the replication protocol, and makes it easier for replicas to achieve consistency. (2) Nezha separates request execution and quorum check, letting clients/proxies undertake quorum check instead of the leader, which effectively relieves leader's burden and enables better pipelining (i.e., avoid the blocking problem in NOPaxos). (3) The use of commutativity further reduces the latency by allowing more requests to be committed in fast path. To verify the benefit of each component, we further conduct an ablation study in §9.4.

9.3 Comparison with Domino and TOQ-based EPaxos

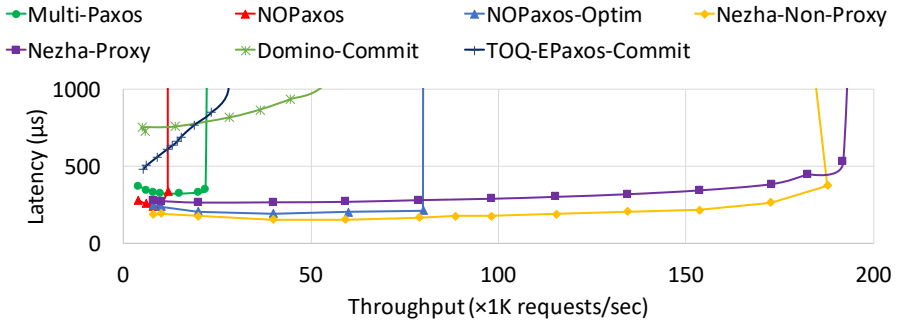
For a fair comparison of Domino and TOQ-based EPaxos with Nezha, we originally wanted to plot their execution latencies in Figure 7. However, both Domino and TOQ-based EPaxos decouple commit from execution, and execution happens much later than commit, which causes high execution latencies. In our experiments, we found that Domino's execution latency exceeds 10 ms and TOQ-based EPaxos' execution latency ranges from 1.3 to 3.3 ms, which are significantly larger than Nezha (as well as our other baselines). This makes it hard to show them in our figure, and hence we plot their commit latencies instead. When comparing the commit latency of Domino and TOQ-based EPaxos to Nezha's execution latency, we still find that Nezha performs better. We believe this is because of differences in implementation: Domino and TOQ-based EPaxos are implemented in Golang with gRPC [56, 68] whereas Nezha and the other baselines are implemented in C++ with UDP [32]). These differences in implementation likely arise from the fact that the additional latency incurred by Golang+gRPC is tolerable for wide-area use cases that typically have higher latencies. We also compare Nezha with Domino and TOQ-based EPaxos below from a design perspective.

Nezha compared with Domino. Nezha and Domino both use synchronized clocks with deadlines attached to messages. They differ in 3 ways. (1) Unlike Domino, Nezha does not decouple commit from execution, which makes it easier for Nezha to be a drop-in replacement for Paxos/Raft, where applications can directly get the execution result from the commit reply. (2) Domino uses a more conservative estimation of OWD at the 95^{th} percentile, rather than Nezha's estimate at the 50^{th} percentile. While a 50^{th} percentile will reduce the fast commit ratio, Nezha compensates for it by leveraging commutativity (see §8.2) and a more optimized slow path that uses speculative execution. (3) Domino directly uses a request's deadline as a position in the log, which causes 2 problems. First, it means that many log entries will be no-ops because there was no request at that particular time/position. Second, when a request arrives later than its deadline, Domino replicas are expected to reject it. But if clock skew occurs at this moment, the replicas will accept it and make the client consider it as committed, even though this request may later be replaced by a no-op. This causes a violation of durability and subsequently linearizability as we explain with error traces in Appendix §F. Compared with Domino's "timestamp-as-log-position," Nezha uses deadlines to reduce packet reordering, but the log positions of requests are eventually decided by the replicas. Therefore, Nezha's correctness is independent of clock skew.

Nezha compared with TOQ-based EPaxos. Nezha differs from TOQ-based EPaxos in 2 ways. (1) TOQ only mitigates the conflict rate for EPaxos but does not optimize the message flow of EPaxos. When it comes to LAN, where there is no difference between LAN RTTs and WAN RTTs, the EPaxos protocol inherently uses more RTTs than Nezha. Besides, there are no proxies to assist EPaxos replicas in request multicast and quorum check, so its replicas' burden is heavier than Nezha. (2) TOQ does not maintain consistent ordering for its released requests, so the protocol (EPaxos) still needs to handle both consistent ordering and set equality to commit



(a) Closed-loop workload



(b) Open-loop workload

Figure 7: Latency vs.throughput

requests, whereas DOM enables Nezha to only focus on set equality. (3) TOQ synchronizes the clocks only among replicas whereas DOM synchronizes clocks among replicas and stateless proxies. Therefore, when it comes to WAN, TOQ-based EPaxos cannot provide optimal WAN latency (1 WAN RTT) for clients in different zones from replicas, whereas DOM can achieve optimal WAN latencies for all clients by deploying a proxy in the same zone with the clients. In Appendix §H, we discuss the comparison between Nezha with EPaxos in WAN, and leave it as our future work to evaluate Nezha and EPaxos in WAN.

9.4 Ablation Study

During the ablation study of Nezha, we remove one component from the full protocol of Nezha each time, and yield three variants, shown as No-DOM, No-QC-Offloading, No-Commutativity in Figure 8. No-DOM variant removes the DOM primitive from Nezha. No-QC-Offloading variant relies on the leader replica to do the quorum check, and it still relies on DOM for consistent ordering (the proxies still perform request mutlicast). No-Commutativity variant disables Nezha's commutativity optimization. We run all protocols under the same setting as Figure 7b.

Figure 8 shows that, removing any of the three components can degrade the performance (i.e., throughput and/or latency).

(1) The No-DOM variant makes the fast path meaningless, because consistent ordering is no longer guaranteed and set equality (i.e. reply messages with consistent hash) no longer indicates the state consistency among replicas. In this case, the No-DOM variant actually becomes the Multi-Paxos protocol with quorum check offloading, and the leader replica still takes the responsibility of ordering and request multicast, which makes No-DOM variant yield a much lower throughput and higher latency.

(2) The No-QC-Offloading variant still uses DOM for ordering and request multicast, but it relies on the leader to do quorum check for every request. Therefore, the leader's burden becomes much heavier than the full protocol, and the heavy bottleneck at the leader replica degrades the throughput and latency performance.

(3) The No-Commutativity variant degrades the fast commit ratio and causes more requests to commit via the slow path. It does not cause a distinct impact on the throughput. However, compared with the full protocol, the lack of commutativity optimization degrades the latency performance by up to 32 %.

9.5 Scalability

Figure 9 shows that, Nezha achieves much higher throughput than the baselines with different number of replicas. However, in open-loop tests with only 10 clients (Figure 9b), the throughput of Nezha-Non-Proxy distinctly degrades from 187.8K requests/sec to 148.7K requests/sec, as the number of replicas grows. This indicates that the clients become the new bottleneck when submitting at high rates. By contrast, when equipped with proxies, Nezha-Proxy maintains a high throughput regardless of the number of replicas. We continue to evaluate the proxy design in §9.6.

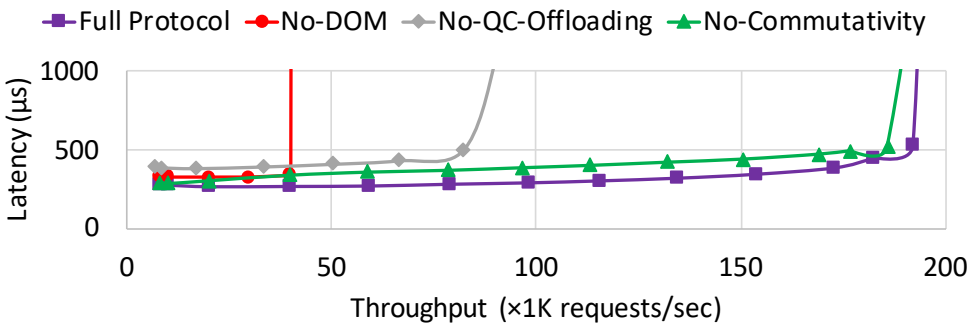
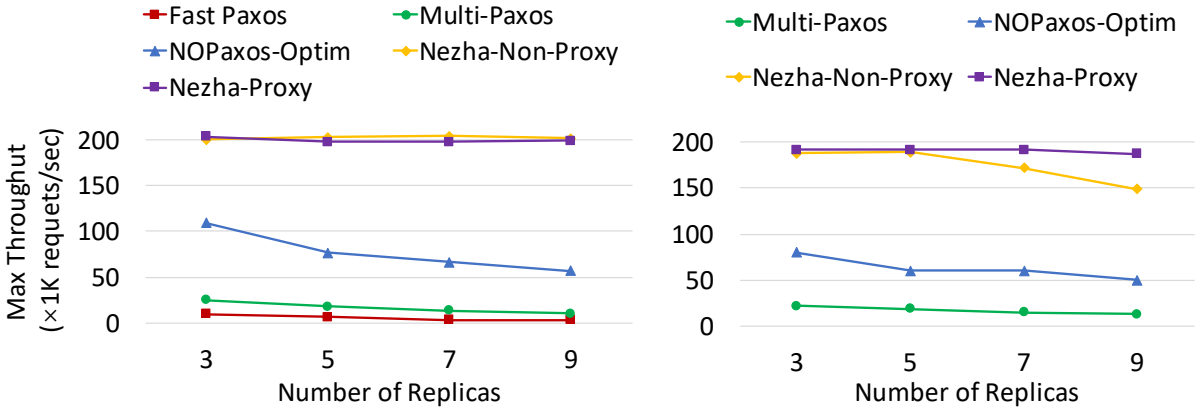
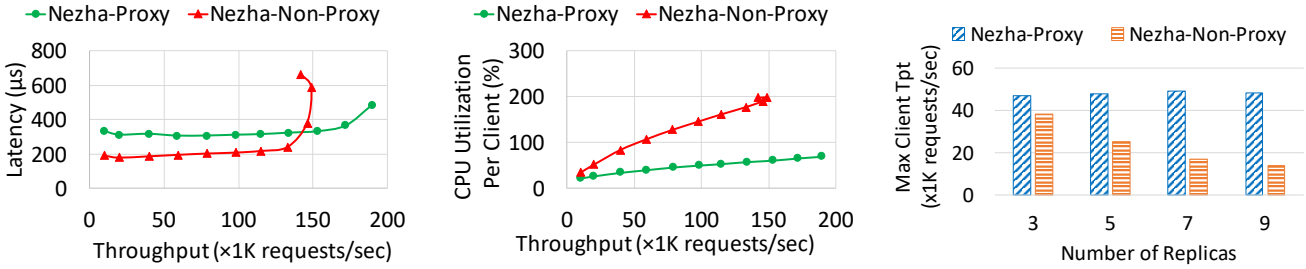


Figure 8: Ablation study of Nezha



(a) Closed-loop **(b) Open-loop**

Figure 9: Max throughput vs. number of replicas



(a) Latency vs. throughput **(b) CPU cost vs. throughput** **(c) Max. client throughput**

Figure 10: Proxy Evaluation

9.6 Proxy Evaluation

Figure 10a and Figure 10b compare the two versions of Nezha with 10 open-loop clients and 9 replicas. Nezha-Proxy also employs 5 proxies. As the client increases its submission rate, we measure the latency and the average CPU utilization per client. Compared with Nezha-Non-Proxy, which sends 9 messages and receives 17 messages (i.e., 9 *fast-replies* and 8 *slow-replies*) for each request, Nezha-Proxy incurs 2 extra message delays, but reduces significant CPU cost at the client side. It achieves even lower latency as the throughput grows, because Nezha-Non-Proxy makes the clients CPU-intensive.

Figure 10c compares the maximum throughput achieved by one client with/without proxies. Given the same CPU resource [4], the throughput of the client without proxies declines distinctly as the number of replicas increases. Such bottlenecks can also occur in the other works with similar offloading design (e.g., Speculative Paxos, NOPaxos, Domino, CURP). By contrast, when equipped with

[4]Every client uses one thread for request submission and another for reply handling.

proxies, the client remains a high throughput regardless of the number of replicas.

9.7 Failure Recovery

We evaluate the failure recovery as shown in Figure 11 and Figure 12. Since follower's crash and recovery do not affect the availability of Nezha, we mainly focus on the evaluation of the leader's crash and recovery. We study two aspects: (1) How long does it take for the remaining replicas to complete a view change with the new leader elected? (2) How long does it take to recover the throughput to the same level as before crash?

We maintain 3 replicas and 10 open-loop clients, and vary per-client submission rate from 1K requests/sec to 20K requests/sec, so the total submission rate varies from 10K requests/sec to 200K requests/sec. Under different submission rates, we kill the leader and measure the time cost of view change. As shown in Figure 11, the time cost grows as the submission rate increases, because there is an increasing amount of state (log) transfer to complete the view change. But the time cost of view change is generally low (150 ms-300 ms) because of the acceleration idea (§8.3) integrated in Nezha.

The time cost to recover the same throughout level (Figure 12) is larger than the time cost of view change, because there are other tasks to complete after the replicas enter the new view. For example, replicas need to relaunch the working threads and reinitialize the contexts; replicas need to handle clients' retried requests, which fail to be responded before crash; followers may need additional state transfer due to lagging too far behind, etc.

Based on the measured trace, we calculate the throughput every 10 ms, and plot the data points in Figure 12. Figure 12 implies that the recovery time is related to the throughput level to recover. A lower throughput level takes a shorter time to recover, and vice versa. It takes approximately 0.7 s, 1.9 s, 4.0 s, to recover to the same throughput level under the submission rate of 20K requests/sec, 100K requests/sec, 200K requests/sec, respectively. As a reference to compare, Figure 3.20 in [61] evaluates the recovery time for an industrial Raft implementation [20], which takes about 6 seconds to recover to 18K requests/sec.

9.8 Nezha vs. Raft

Raft establishes its correctness on log persistence and relies on the stable storage for stronger fault tolerance (e.g. power failure). For a fair comparison to Raft, we convert Nezha from its diskless operation to a disk-based version, making it achieve the same targets as Raft. Before Nezha replicas send replies, they first persist the corresponding log entry (including *view-id* and *crash-vector*) to stable storage. Then, if a replica is relaunched, it can recover its state and replay the *fast-replies*/*slow-replies*. We want to study whether Nezha is fundamentally more I/O intensive than Raft.

We initially use the original Raft implementation [49] (Raft-1 in Figure 13), which is written in C++, but uses a slower communication library based on TCP, and involves additional mechanisms (e.g. snapshotting). Raft-1 can only work in closed-loop tests because of its blocking API. For Raft-1, we use its default batching and pipeline mechanism, and noticed that Raft-1 achieves very low throughput of 4.5K requests/sec on Google Cloud VMs equipped with zonal standard persistent disk [18]. Hence, we implement and optimize Raft (Raft-2), by using the Multi-Paxos code from [32] as a starting point. For both Raft-2 and Nezha, we tune their batch sizes to reach the best throughput. Our evaluation shows that Nezha outperforms Raft in both closed-loop test (Figure 13) and open-loop test (Figure 14). We also see that there is little difference in latency with or without a proxy in Nezha because latencies are now dominated by disk writes, not message delays.

10 APPLICATION PERFORMANCE

Redis. Redis [54] is a typical in-memory key-value store. We choose YCSB-A [67] as the workload, which operate on 1000 keys with `HMSET` and `HGETALL`. We use 20 closed-loop clients to submit requests, which can saturate the processing capacity of the unreplicated Redis. Figure 15 illustrates the maximum throughout of each protocol under 10 ms SLO. Nezha outperforms all the baselines on this metric: it outperforms Fast Paxos by 2.9×, Multi-Paxos by 1.9×, and NOPaxos by 1.3×. Its throughput is within 5.9% that of the unreplicated system.

CloudEx. CloudEx [17] is a research fair-access financial exchange system for public cloud. There are three roles involved in CloudEx: matching engine, gateways and market participants. To provide fault tolerance, we replicate the matching engine and co-locate one gateway with one proxy. Market participants are unmodified. Before porting it to Nezha, we improved the performance of CloudEx, compared with the version in [17], by multithreading and replacing ZMQ [70] with raw UDP transmission. We first run the unreplicated CloudEx with its dynamic delay bounds (DDP) strategy disabled [17]. We configure a fixed sequencer delay parameter (d_s) of 200μs. Similar to [17], we launch a cluster including 48 market participants and 16 gateways, with 3 participants attached to one gateway. The matching engine is configured with 1 shard and 100 symbols. We vary the order submission rate of market participants, and find the matching engine is saturated at 43.10K orders/sec, achieving an inbound unfairness ratio of 1.49%.

We then run CloudEx atop the four protocols with the same setting. In Figure 16, only Nezha reaches the throughput (42.93K orders/sec) to nearly saturate the matching engine, and also yields a close inbound unfairness ratio of 1.97%. We further compare the end-to-end latency (i.e., from order submission to the order confirmation from the matching engine) and order processing latency (i.e., from order submission to receiving the execution result from the matching engine.) between Nezha and the unreplicated CloudEx. In Figure 17, Nezha prolongs the end-to-end latency by 19.7 % (344 µs vs. 288 µs), but achieves very close order processing latency to the unreplicated version (426 µs vs. 407 µs).

11 RELATED WORK

Consensus protocols. Classical consensus protocols, e.g., Multi-Paxos, Raft, and Viewstamped Replication make no distinction between a fast and slow path: all client requests incur the same latency. Nezha uses an optimistic approach to improve latency in the common case. Mencius [40] exploits a multi-leader design to mitigate the single leader bottleneck in Multi-Paxos. However, it introduces extra coordination cost among multiple leaders and further, the crash of any of the leaders temporarily stops progress.

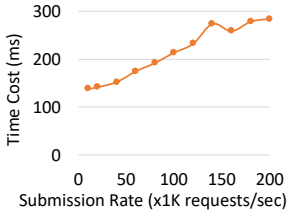
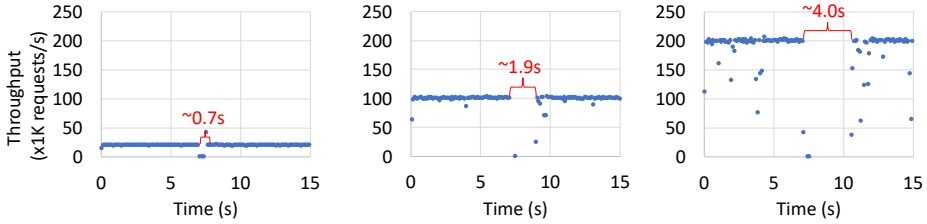


Figure 11: Time cost of view change



(a) 20K requests/sec **(b) 100K requests/sec** **(c) 200K requests/sec**

Figure 12: Time cost to recover to the same throughput level

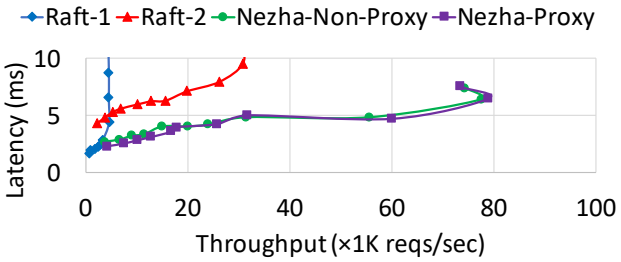


Figure 13: Nezha vs. Raft (closed-loop)

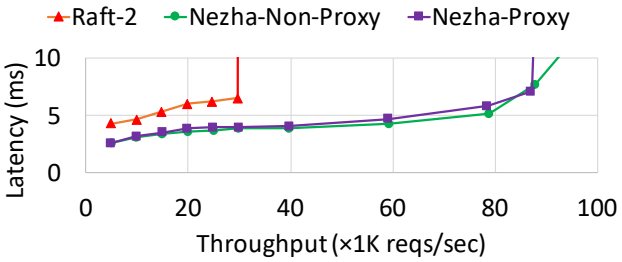


Figure 14: Nezha vs. Raft (open-loop)

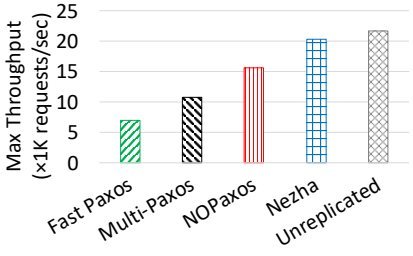


Figure 15: Redis throughput with a 10 ms latency SLO

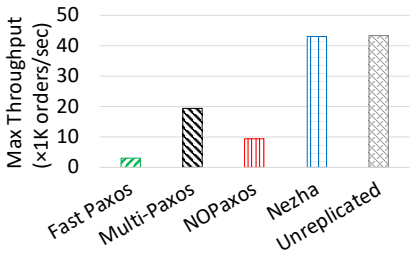


Figure 16: CloudEx throughput

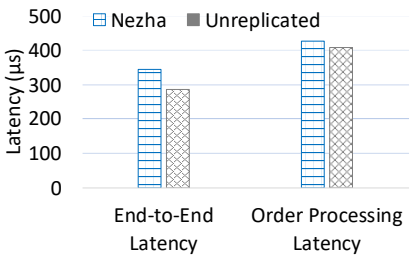


Figure 17: CloudEx latency

By contrast, Nezha reduces the leader's bottleneck using proxies and followers' crash does not affect progress. EPaxos [47] can achieve optimal WAN latency in the fast path, but when it comes to the LAN scenarios we focus on, it performs worse than Multi-Paxos [4]. CURP [51] can complete commutative requests in 1 RTT, but doesn't take advantage of consistent ordering: hence, it costs up to 3 RTTs even if all witnesses process the non-commutative requests in the same order. SPaxos [6], BPaxos [65] and Compartmentalized Paxos [43] address the throughput scaling of consensus protocols with modularity, trading more latency for throughput improvement. The proxy design in Nezha is similar to compartmentalization [43], but Nezha's proxies are stateless. By contrast, [6, 43, 65] use stateful proxies, which complicates fault tolerance.

Network primitives to improve consensus. Recent works consider building network primitives to accelerate consensus protocols. 4 other primitives closely related to DOM, namely, mostly-ordered multicast (MOM) [53], ordered unreliable multicast (OUM) [33], timestamp-ordered queuing (TOQ) [60] and sequenced broadcast (SB) [58]. From the perspective of deployability, DOM and TOQ are both based on software clock synchronization whereas MOM and OUM rely on highly engineered network. This gives DOM and TOQ an advantage over MOM/OUM in environments like the cloud. On the other hand, requests output from MOM and TOQ can still result in inconsistent ordering. By contrast, DOM and OUM guarantee consistent ordering of released requests. DOM's guarantees are stronger than MOM because MOM can occasionally reorder requests, but are weaker than OUM because OUM also provides gap detection. We include a formal comparison in Appendix §G. SB is a new primitive for Byzantine fault tolerance. It works in an epoch-based manner and achieves high throughput through load balancing. However, its latency is in the order of seconds.

Clock synchronization applied to consensus protocols. CRaft [61] and CockRoachDB [59] use clock synchronization to improve the throughput of Raft. However, they base their correctness on the assumption of a known worst-case clock error bound, which is not practical for high-accuracy clock synchronization [35, 37, 38]. Domino [69] and TOQ [60] try using clock synchronization to accelerate Fast Paxos and EPaxos respectively. We evaluate and

compare them with Nezha in §9.3, and include more details in Appendix §F and §H.

12 CONCLUSION AND FUTURE WORK

Recent development of accurate software clock synchronization techniques brings us new opportunities to develop novel consensus protocols to achieve high performance in public cloud. Leveraging this, we present Nezha in the paper, which can be easily deployed in the public cloud, and achieves both higher throughput and lower latency than baselines.

We are considering three lines of future work. First, we intend to replace the Multi-Paxos/Raft backend used by some industrial systems (e.g., Kubernetes, Apache Pulsar, etc) so as to boost their performance. Second, although we target at LAN scenarios in this paper, applying Nezha in WAN will be an interesting follow-up work. Third, we believe DOM can also be applied to other domains. We plan to integrate DOM with concurrency control algorithms (e.g. Two-Phase Locking, Optimistic Concurrency Control, etc) to improve their performance, or invent new concurrency control protocols based on DOM.

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APPENDICES

In this appendix, we include the following:

- The explanation of Nezha's recovery (§A).
- The correctness proof of Nezha (§B).
- The evaluation of Nezha under different workloads (§C).
- The detailed discussion and evaluation about the effect of clock variance on Nezha's performance (§D).
- The deployment experience and evaluation of Derecho in bare-metal servers and public cloud (§E)
- The analysis on the incorrectness of Domino due to clock skew/failure (§F).
- Formal comparison among DOM, MOM, and OUM (§G).
- The discussion about deploying Nezha in WAN and its expected advantages over EPaxos in WAN (§H).
- The TLA+ specification of Nezha (§I).

A RECOVERY PROTOCOL AND ALGORITHMS

We explain how Nezha leverages the diskless crash recovery algorithm [36] from Viewstamped Replication in 3 steps. First, we explain how we adopt the recent concept of crash-vectors [41, 42] to fix the incorrectness in the crash recovery algorithm. Second, we explain how a replica rejoins Nezha following a crash. Third, we describe how the leader election works if the leader crashes.

A.1 Crash Vector

Like Viewstamped Replication, Speculative Paxos and NOPaxos, Nezha also adopts diskless recovery to improve performance. However, in contrast to them, Nezha avoids the effect of *stray messages* [27] (i.e., messages that are sent out but not delivered before replica crash, so that the relaunched replicas forget them) using the *crash-vector* [41, 42]. *crash-vector* is a vector containing $2f + 1$ integer counters corresponding to the $2f + 1$ replicas. Each replica maintains such a vector, with all counters initialized as 0s.

crash-vectors can be aggregated by taking the max operation element-wise to produce a new *crash-vector*. During the replica rejoin (§A.2) and leader change(§A.3) process, replicas send their *crash-vectors* to each other. Receivers can make their *crash-vectors* more up-to-date by aggregating their *crash-vector* with the *crash-vector* from the sender. Meanwhile, by comparing its local *crash-vector* and the sender's *crash-vector*, the receiver can recognize whether or not the sender's message is a potential *stray message* (refer to [42] for detailed description of *crash-vector*).

Nezha uses *crash-vector*s to avoid two types of *stray messages*, i.e. (1) the *stray messages* during the view change process and (2) the *stray messages* (*fast-replies*) during quorum check. (1) has been clearly explained in [42], so here we only sketch how *crash-vector* works preserve the protocol correctness during the view change process. (2) has not been disclosed in prior works, so we will explain more details in §A.4 after we complete the explanation of the replica rejoin (§A.2) and leader change (§A.3).

A.1.1 Stray Message during View Change. There can be *stray messages* during the view change process: replicas mistakenly elect a leader, whose state falls behind the others, finally causing permanent loss of committed requests. The *crash-vector* prevents the *stray messages* effect because it enables the replicas to recognize potential *stray messages* by comparing a *crash-vector* received from a replica with the local *crash-vector*. During recovery, the RECOVERING replica first recovers its *crash-vector* by collecting and aggregating the *crash-vectors* from a majority of NORMAL replicas. Then, the replica increments its own counter (i.e. replica i increments the ith counter in the vector) and tags the new *crash-vector* to the messages sent afterwards. Once the update of *crash-vector* is exposed to the other replicas, they can recognize the *stray messages* sent by the replica before crash (i.e., those messages have a smaller value at the ith counter), and avoid processing those messages. Thus, the recovery will not be affected by *stray messages*.

A.1.2 Stray Message during Quorum Check. Stray messages can also occur during the quorum check in the fast path: some replicas send *fast-replies* and crash after that. The reply messages in the fast path (i.e., *fast-reply*) may become *stray messages* and participate into the quorum check, which makes the proxies/clients prematurely believe the request has been persisted to a super-majority of replicas, but actually not yet (i.e. the recovered replicas may not hold the requests after their recovery).

In brief, the *crash-vector* prevents the effect of such *stray fast-replies*, because we include the information of *crash-vector*s in the *fast-replies* (§6.3). When a failed replica rejoins the system (Algorithm 1), it leads to the update of *crash-vector*s for the leader and other remaining followers, so these replicas will send *fast-replies* with different *hashes* from the *stray fast-replies* sent by the rejoined replica. Therefore, the stray *fast-replies* from the rejoined replica and the normal *fast-replies* from the other replicas cannot form the super quorum together. After we describe the replica rejoin (§A.2) and leader change (§A.3), we come back to explain the details in §A.4.

A.2 Replica Rejoin

Crashed replicas can rejoin the system as followers. After the replica crashes and is relaunched, it sets its *status* as RECOVERING. Before it can resume request processing, the replica needs to recover its replica state, including *crash-vector*, *view-id*, *log* and *sync-point*. With reference to Algorithm 1, we explain how the replica rejoin process works.

Step 1: The replica sets its *status* as RECOVERING (line 2), and broadcasts the same CRASH-VECTOR-REQ to all replicas to request their *crash-vectors*. A *nonce* (line 4) is included in the message, which is a random string locally unique on this replica, i.e., this replica has never used this *nonce* [5].

Step 2: After receiving the CRASH-VECTOR-REQ, replicas with NORMAL status reply to the recovering replica with <CRASH-VECTOR-REP, *nonce*, *crash-vector*> (line 40-47).

Step 3: The recovering replica waits until it receives the corresponding replies (containing the same *nonce*) from a majority ($f + 1$) of replicas (line 23). Then it aggregates the $f + 1$ *crash-vectors* by taking the maximum in each dimension (line 7, line 99-104). After obtaining the aggregated crash vector *cv*, the replica increment its own dimension, i.e. $cv[replica\text{-}id] = cv[replica\text{-}id] + 1$ (line 8).

[5] There are many options available to generate the locally unique *nonce* string [36, 42]. Nezha uses the universally unique identifier (UUID) (GENERATE-UUID in line 4), which have been widely supported by modern software systems.

Algorithm 1 Replica rejoin

Local State: ▷
nonce, ▷ A locally unique string on this replica
C, ▷ Reply set of CRASH-VECTOR-REP
R, ▷ Reply set of RECOVERY-REP
r, ▷ short for *replica-id* (the message sender)
cv, ▷ short for *crash-vector*
status, *view-id*, *last-normal-view*, *log*, *sync-point* ▷ Other attributes

1: **upon** RECOVER **do**
2: $status$ =RECOVERING
3: $C = \emptyset$
4: $nonce$ = GENERATE-UUID
5: READ-CRASH-VECTOR
6: *cv-set*= $\{m.cv | m \in C\}$
7: cv =AGGREGATE(*cv-set* $\cup \{cv\}$)
8: $cv[r] = cv[r] + 1$ ▷ Increment its own counter
9: **do**
10: $R = \emptyset$
11: READ-RECOVERY-INFO
12: *highest-view* = $\max\{m.v | m \in R\}$
13: *leader-id* = *highest-view* % $(2f + 1)$
14: **while** (*leader-id* = r)
15: Pick $m \in R$: $m.v$ = *highest-view*
16: STATE-TRANSFER(*leader-id*)
17: **function** READ-CRASH-VECTOR
18: $m.type$ = CRASH-VECTOR-REQ
19: $m.r = r$
20: $m.nonce = nonce$
▷ Broadcast CRASH-VECTOR-REQ to all replicas
21: **for** $i \leftarrow$0 to $2f$ **do**
22: SEND-MESSAGE(m, i) ▷ Send message m to the replica i
23: Wait until $|C| \geq f + 1$ ▷ C is initialized as $\emptyset$ by the caller
24: **return**
25: **function** READ-RECOVERY-INFO
26: $m.type$ =RECOVERY-REQ
27: $m.r = r$
28: $m.cv = cv$
▷ Broadcast RECOVERY-REQ to all replicas
29: **for** $i \leftarrow$0 to $2f$ **do**
30: SEND-MESSAGE(m, i)
31: Wait until $|R| \geq f + 1$ ▷ R is initialized as $\emptyset$ by the caller
32: **return**
33: **function** STATE-TRANSFER(i)
34: $m.type$ =STATE-TRANSFER-REQ
35: $m.r = r$
36: $m.cv = cv$
37: SEND-MESSAGE(m, i)
38: Wait until $status$ = NORMAL
39: **return**
40: **upon** *RECEIVING* CRASH-VECTOR-REQ, m **do**
41: **if** $status \neq$ NORMAL **then**
42: **return**
43: $m'.type$ =CRASH-VECTOR-REP
44: $m'.r = r$
45: $m'.nonce = m.nonce$
46: $m'.cv = cv$
47: SEND-MESSAGE(m', $m.r$)
48: **upon** *RECEIVING* CRASH-VECTOR-REP, m **do**
49: **if** $status \neq$ RECOVERING **then**
50: **return**
51: **if** $nonce \neq m.nonce$ **then**
52: **return**
53: $C = C \cup \{m\}$
54: **upon** *RECEIVING* RECOVERY-REQ, m **do**
55: **if** $status \neq$ NORMAL **then**
56: **return**
57: cv =AGGREGATE(cv, $m.cv$)
58: $m'.type$ = RECOVERY-REP
59: $m'.r = r$
60: $m'.v = view\text{-}id$
61: $m'.cv = cv$
62: SEND-MESSAGE(m', $m.r$)
63: **upon** *RECEIVING* RECOVERY-REP, m **do**
64: **if** $status \neq$ RECOVERING **then**
65: **return**
66: **if** CHECK-CRASH-VECTOR(m, cv)=**false then**
67: Resend RECOVERY-REQ to $m.r$
68: **else** ▷ Remove stray messages and add the fresh one
69: $R' = \{m' \in R \mid m'.cv[m'.r] < cv[m'.r]\}$
70: $R = R \cup \{m\} - R'$
71: $\forall m' \in R'$, resend RECOVERY-REQ to $m'.r$
72: **upon** *RECEIVING* STATE-TRANSFER-REQ, m **do**
73: **if** $status \neq$ NORMAL **then**
74: **return**
75: **if** CHECK-CRASH-VECTOR(m, cv)=**false then**
76: **return**
77: $m'.type$ = STATE-TRANSFER-REP
78: $m'.log = log$
79: $m'.v = view\text{-}id$
80: $m'.sp = sync\text{-}point$
81: $m'.cv = cv$
82: SEND-MESSAGE(m', $m.r$)
83: **upon** *RECEIVING* STATE-TRANSFER-REP, m **do**
84: **if** $status \neq$ RECOVERING **then**
85: **return**
86: **if** CHECK-CRASH-VECTOR(m, cv)=**false then**
87: **return**
88: $log = m.log$
89: $last\text{-}normal\text{-}view = view\text{-}id = m.v$
90: $log = m.log$
91: $sync\text{-}point = m.sp$
92: $status$ = NORMAL ▷ Rejoin as a NORMAL follower
93: **function** CHECK-CRASH-VECTOR(m, cv)
94: **if** $m.cv[m.r] < cv[m.r]$ **then** ▷ A potential *stray message*
95: **return false**
96: **else**
97: cv = AGGREGATE($\{cv, m.cv\}$) ▷ Update local cv
98: **return true**
99: **function** AGGREGATE(*cv-set*)
100: $ret = [\underbrace{0 \ldots 0}_{2f+1}]$
101: **for** $c \in$ *cv-set* **do**
102: **for** $i \leftarrow 0$ to $2f$ **do**
103: $ret[i] = \max(ret[i], c[i])$
104: **return** ret

Step 4: The recovering replica broadcasts a recovery request to all replicas, which includes its *crash-vector*, i.e. <RECOVERY-REQ, *cv*> (line 11, line 26-30).

Step 5: After receiving the RECOVERY-REQ, replicas with NORMAL status update their own *crash-vectors* by aggregating with *cv*, obtained from the request in step 4. Then, these replicas send back a reply including their own *view-id* and *crash-vector*, i.e. <RECOVERY-REP, *view-id*, *crash-vector*> (line 54-63).

Step 6: The recovering replica waits until it receives the recovery replies from $f + 1$ replicas (line 31). If the RECOVERY-REP is not a *stray message*, it updates its own *crash-vector* by aggregating it with the *crash-vectors* included in these replies (line 66); otherwise, it resends RECOVERY-REQ to that replica, asking for a fresh message (line 67). Because the *crash-vectors* may have been updated (line 66), those RECOVERY-REP which have been received can also become *stray messages* because their *crash-vectors* are no longer fresh enough. Therefore, we also remove them (R' in line 69) from the reply set R (line 70), and resend requests to the related replicas for fresher replies (line 71).

Step 7: The RECOVERING replica picks the highest *view-id* among the $f + 1$ replies (line 12). From the highest *view-id*, it knows the corresponding leader of this view (line 13). If the RECOVERING replica happens to be the leader of this view, it keeps broadcasting the recovery request (line 9-14), until the majority elects a new leader among themselves. Otherwise, the RECOVERING replica fetches the *log*, *sync-point*, *view-id* from the leader via a state transfer (line 16, line 33-39). After that, the replica set its *status* to NORMAL and can continue to process the incoming requests.

Specially, the RECOVERING replica(s) do not participate in the view change process(§A.3). When the majority of replicas are conducting a view change (possibly due to leader failure), the RECOVERING replica(s) just wait until the majority completes the view change and elects the new leader.

A.3 Leader Change

When the follower(s) suspect that the leader has failed, they stop processing new client requests. Instead, they perform the view change protocol to elect a new leader and resume request processing. With reference to Algorithm 2, we explain the details of the view change process.

Step 1: When a replica fails to receive the heartbeat (i.e., *sync* message) from the leader for a threshold of time, it suspects the leader has failed. Then, it sets its *status* as VIEWCHANGE, increments its *view-id*, and broadcasts a view change request to all replicas including its *crash-vector*, i.e. <VIEW-CHANGE-REQ, *view-id*, *replica-id*, *cv*> (line 6-10) [6]. The replica switches its *status* from NORMAL to VIEWCHANGE, and enters the view change process.

Step 2: After receiving a VIEW-CHANGE-REQ message, the recipient checks the *cv* and *replica-id* with its own *crash-vector* (line 32). If this message is a potential *stray message*, then the recipient ignores it. Otherwise, the recipient updates its *crach-vector* by aggregation. After that, the recipient also participates in the view change (line 35) if its *view-id* is lower than that included in the VIEW-CHANGE-REQ message.

[6]The view change request will be rebroadcast if the replica times out but is still waiting for the view change process to complete. The same is also true for the view change message described in the next step.

Step 3: All replicas under view change send a message <VIEW-CHANGE, *view-id*, *log*, *sync-point*, *last-normal-view*> to the leader of the new view (*replica-id*= *view-id*%$(2f + 1)$) (line 11). Here *last-normal-view* indicates the last view in which the replica's *status* was NORMAL.

Step 4: After the new leader receives the VIEW-CHANGE messages from f followers with matching *view-id*s, it can recover the system state by merging the *log*s from the f+1 replicas including itself (line 67). The new leader only merges the *log*s with the highest *last-normal-view*, because a smaller *last-normal-view* indicates the replica has lagged behind for several view changes, thus its *sync-point* cannot be larger than the other replicas with higher *last-normal-view* values. Therefore, it makes no contribution to the recovery and does not need to join.

Step 5: The new leader initializes an empty log list (denoted as *new-log*) (line 74). Among the VIEW-CHANGE messages with the highest *last-normal-view*, it picks the one with the largest *sync-point* (line 75-77). Then it directly copies the log entries from that message up to the *sync-point* (line 78-82).

Step 6: Afterwards, the new leader checks the remaining entries with larger *deadline*s than *sync-point* (ling 83-88). If the same entry (2 entries are the same iff they have the same <*deadline*, *client-id*, *request-id*>) exists in at least $\lceil f/2 \rceil + 1$ out of the $f + 1$ *log*s, then leader appends the entry to *new-log*.

Step 7: After *new-log* is built, the new leader broadcasts <START-VIEW, *cv*, *view-id*, *new-log*> to all replicas (line 68-70).

Step 8: After receiving the START-VIEW message with a *view-id* greater than or equal to its *view-id*, the replica updates its *view-id* and *last-normal-view* (line 97), and replaces its *log* with *new-log* (line 98). Besides, it updates *sync-point* as the last entry in the new *log* (line 98), because all the entries are consistent with the leader. Finally, replicas set their *status*es to NORMAL (line 100), and the system state is fully recovered.

Step 9: After the system is fully recovered, the replicas can continue to process the incoming requests. Recall in §8.2 that the incoming request is allowed to enter the *early-buffer* if its deadline is larger than *the last released request* which is not commutative. To ensure consistent ordering, the eligibility check is still required for the incoming request even if it is the first one arriving at the replica after recovery. The replica considers the entries (requests) in the recovered *log*, which are not commutative to the incoming request, and chooses the one as *the last released request* with the largest deadline among them. The incoming request can enter the *early-buffer* if its deadline than *the last released request*, otherwise, it is put into the *late-buffer*.

Note that the view change protocol chooses the leader in a round-robin way (*view-id*%$(2f + 1)$). Specially, a view change process may not succeed because the new leader also fails (as mentioned in [36]). In this case (i.e. after followers have spent a threshold of time without completing the view change), followers will continue to increment their *view-id*s to initiate a further view change, with yet another leader.

After the replica rejoin or leader change process, replicas' *crash-vector*s will be updated. Due to packet drop, some replicas may fail to receive the update of *crash-vector*s during the recovery, thus they cannot contribute to the quorum check of the fast path in the following request processing, because their *crash-vector*s are still

Algorithm 2 Leader change

```
    Local State:                                                          ▷
    V,                                                  ▷ Reply set of VIEW-CHANGE
    r,                                  ▷ short for replica-id (the message sender)
    cv,                                                 ▷ short for crash-vector
    last-normal-view,                                   ▷ The most recent view in which
                                                        ▷ the replica's status is NORMAL
    status, view-id, log, sync-point                    ▷ Other attributes
 1: upon SUSPECT LEADER FAILURE do
 2:     do
 3:         INITIATE-VIEW-CHANGE(view-id + 1)
 4:     while (status ≠ NORMAL)
 5: function INITIATE-VIEW-CHANGE(v)
 6:     status = VIEWCHANGE
 7:     view-id = v
 8:     V = ∅
        ▷ Broadcast VIEW-CHANGE-REQ to all replicas
 9:     for i ←0 to 2f do
10:         SEND-VIEW-CHANGE-REQ(i)
        ▷ Send VIEW-CHANGE to the new leader
11:     SEND-VIEW-CHANGE(v%(2f + 1))
12:     Wait until status = NORMAL or TIMEOUT
13:     return
14: function SEND-VIEW-CHANGE-REQ(i)
15:     m.type = VIEW-CHANGE-REQ
16:     m.v = view-id
17:     m.cv = cv
18:     SEND-MESSAGE(m, i)
19:     return
20: function SEND-VIEW-CHANGE(i)
21:     m.type = VIEW-CHANGE
22:     m.v = view-id
23:     m.cv = cv
24:     m.log = log
25:     m.sp = sync-point
26:     m.lnv = last-normal-view
27:     SEND-MESSAGE(m, i)
28:     return
29: upon RECEIVING VIEW-CHANGE-REQ, m do
30:     if status = RECOVERING then
31:         return
32:     if CHECK-CRASH-VECTOR(m, cv)=false then
33:         return
34:     if m.v > view-id then
35:         INITIATE-VIEW-CHANGE(m.v)
36:     else
37:         if status = NORMAL then
38:             SEND-START-VIEW(m.r)
39:         else          ▷ The leader is asking for fresher VIEW-CHANGE
40:             SEND-VIEW-CHANGE(m.r)
41: function SEND-START-VIEW(i)
42:     m.type = START-VIEW
43:     m.v = view-id
44:     m.cv = cv
45:     m.log = log
46:     SEND-MESSAGE(m, i)
47:     return
48: upon RECEIVING VIEW-CHANGE, m do
49:     if status = RECOVERING then
50:         return
51:     if CHECK-CRASH-VECTOR(m, cv)=false then
52:         return
53:     if status = NORMAL then
54:         if m.v > view-id then
55:             INITIATE-VIEW-CHANGE(m.v)
56:         else                                  ▷ The sender lags behind
57:             SEND-START-VIEW(m.r)
58:     else if status = VIEWCHANGE then
59:         if m.v > view-id then
60:             INITIATE-VIEW-CHANGE(m.v)
61:         else if m.v < view-id then            ▷ The sender lags behind
62:             SEND-VIEW-CHANGE-REQ(m.r)
63:         else          ▷ Remove stray messages and add the fresh one
64:             V' = {m' ∈ V | m'.cv[m'.r] < cv[m'.r]}
65:             V = V ∪ {m} − V'
66:             ∀m' ∈ V', resend VIEW-CHANGE-REQ to m'.r
67:             if |V| ≥ f + 1 then
68:                 log = MERGE-LOG(V)
69:                 for i ←0 to 2f do
70:                     SEND-START-VIEW(i)
71:                 last-normal-view = view-id
72:                 status = NORMAL               ▷ Leader becomes NORMAL
73: function MERGE-LOG(V)
74:     new-log = ∅
75:     largest-normal-view = max{m.lnv|m ∈ V}
76:     largest-sync-point = max{m.sp|m ∈ V
                and m.lnv = largest-normal-view}
77:     Pick m ∈ V:
            m.lnv = largest-normal-view and
            m.sp = largest-sync-point
        ▷ Directly copy entries up to sync-point
78:     for e ∈ m.log do          ▷ m.log is already sorted by deadlines
79:         if e.deadline ≤ largest-sync-point.deadline then
80:             new-log.append(e)
81:         else
82:             break
        ▷ Add other committed entries beyond sync-point
83:     entries = {e|e ∈ m.log
                    and e.deadline > largest-sync-point.deadline
                    and m.lnv = largest-normal-view}
84:     for e ∈ entries do
        ▷ Check how many replicas contain e
85:         S = {m|m ∈ V and e ∈ m.log}
86:         if |S| ≥ ⌈f/2⌉ + 1 then
87:             log.append(e)
88:     Sort new-log by entries' deadlines
89:     return new-log
90: upon RECEIVING START-VIEW, m do
91:     if status = RECOVERING then
92:         return
93:     if CHECK-CRASH-VECTOR(m, cv)=false then
94:         return
95:     if m.v < view-id then
96:         return
97:     last-normal-view = view-id = m.v
98:     log = m.log
99:     sync-point = log.last()
100:    status = NORMAL                           ▷ Followers become NORMAL
```

old and cannot generate the consistent hash with the leader's hash. To enable every replica to obtain the fresh information of *crash-vector*s rapidly, the leader can piggyback the fresh *crash-vector*s in the *sync* messages, so that replicas can check and update their *crash-vector*s as soon as possible.

A.4 Why Does *crash-vector* Prevent *Stray Message* Effect during Quorum Check?

The *stray message*s can cause a common bug for most optimistic consensus protocols (e.g., [33, 53, 69]) when they conduct the quorum check in the fast path. Below we summarize the general pattern to cause the loss of committed requests (durability violation). Then, we explain why Nezha can avoid such problems by use of *crash-vector*.

General Error Pattern. Consider a request is delayed in the network, whenever it arrives at one replica, that replica sends a reply and immediately crashes afterwards, then the crashed replica recover from the others and gets an empty log list (because the other replicas have not received the request). After each replica completes such behavior, the client gets replies from all the replicas but actually none of them is holding the request. Such a pattern does not violate the failure model, but causes permanent loss of committed requests.

Reviewing the existing opportunistic protocols, Speculative Paxos, NOPaxos and Domino all suffer from such cases. CURP [51] can avoid the *stray message* effect by assuming the existence of a configuration manager, which never produces stray messages (e.g., by using stable storage). Whenever the witnesses crash and are relaunched, the configuration manager need to refresh the information for the master replica as well as the clients, so that clients can detect the stray messages during quorum check and avoid incorrectness.

Nezha avoids such error cases by including the information of *crash-vector* in the hash of *fast-replies* (§8.1), which prevent stray reply messages from forming the super-quorum in the fast path and creating an *illusion* to the proxies/clients. We analyze in more details below.

Regarding the general pattern above,

(1) When the follower(s) fail, they need to contact the leader and complete the state transfer before their recovery (Algorithm 1).

- If the leader has already received the request before the state transfer, then after the follower's recovery, it can remember the *fast-reply* that it has sent before crash, and can replay it. In this case, the *fast-reply* is not a stray message.
- If the leader has not received the request before the state transfer, then the leader's *crash-vector* will be updated after receiving the follower's STATE-TRANSFER-REQ (line 75-76 in Algorithm 1), which includes a different *crash-vector* (the follower has incremented its own counter). Therefore, the *hash* of the leader's *fast-reply* is computed with the aggregated *crash-vector*, and will be different from that included in the *fast-reply* (stray message) sent by the follower before crash, i.e. the leader's *fast-reply* and the followers' stray *fast-replies* cannot form a super quorum.

(2) When the leader fails, based on Algorithm 2, the view change will elect a new leader. *crash-vectors* ensure the view change process is not affected by *stray messages*. After the view change is completed, the *view-id* is incremented. At least $f + 1$ replicas after the view change will send *fast-replies* with higher *view-id*s. Because the quorum check requires reply messages have matching *view-id*s, the *stray fast-replies* (sent by the old leader) can not form a super quorum together with the *fast-replies* sent by the replicas after the view change.

Nezha's slow path does not suffer from *stray message* effect, because there is causal relation between the leader's state update (advancing its *sync-point*) and followers' sending *slow-replies*.

(1) When followers crash and recover, they copy the state from the leader. The followers' state before crash is no fresher than their recovered state, so the followers have no *stray slow-replies*, i.e. the followers can remember the *slow-replies* they have sent before crash and can replay them.

(2) When the leader crashes and recovers, it can only rejoin as a follower replica after the new leader has been elected (§A.2), so the old leader's reply messages before crash have smaller *view-id*s, compared with the *slow-replies* of replicas after the view change. With matching *view-id*s, these reply messages cannot form a quorum together in the slow path.

A.5 Reconfiguration

While it has not been implemented, Nezha can also use the standard reconfiguration protocol from Viewstamped Replication [36] (with its incorrectness fixed by *crash-vector* [41, 42]) to change the membership of the replica group, such as replacing the failed replicas with the new ones that have a new disk, increasing/decreasing the number of replicas in the system, etc.

B CORRECTNESS PROOF OF NEZHA

With the normal behavior described in §6.3~§6.4, we can prove that the recovery protocol of Nezha guarantees the following correctness properties.

- **Durability:** if a client considers a request as committed, the request survives replica crashes.
- **Consistency:** if a client considers a request as committed, the execution result of this request remains unchanged after the replica's crash and recovery.
- **Linearizability:** A request appears to be executed exactly once between start and completion. The definition of linearizability can also be reworded as: if the execution of a request is observed by the issuing client or other clients, no contrary observation can occur afterwards (i.e., it should not appear to revert or be reordered).

B.1 Proof of Durability

The client/proxy considers *req* as committed after receiving the corresponding quorum or super quorum of replies. Since the quorum checks on both the fast path and slow path require the leader's reply, a committed request indicates that the request must have been accepted by the leader. If a follower crashes, it does not affect durability because the recovered followers directly copy *log* from the leader via state transfer (Step 7 in §A.2) before serving new requests. Hence, we consider the durability property during leader crashes.

(1) If the client/proxy commits *req* in the fast path, it means the request has been replicated to the leader and at least $f + \lceil f/2 \rceil$ followers. When the leader crashes, among any group of $f + 1$ replicas, *req* exists in at least $\lceil f/2 \rceil + 1$ of them because of quorum intersection. Hence, *req* will be added to the *new-log* in Step 6 in §A.3, and eventually recovered.

(2) If the client/proxy commits *req* in the slow path, it means *req* has been synced with the leader by at least $f + 1$ replicas, i.e., there are at least $f + 1$ replicas containing a *sync-point* whose *deadline* is greater than or equal to *req*'s *deadline*. Due to quorum intersection, there will at least one replica which has the *sync-point* in Step 4 of §A.3. Therefore, *req* will be directly copied to *new-log* in Step 5 of §A.3, and eventually recovered.

B.2 Proof of Consistency

Without considering the acceleration of recovery mentioned in §7, we prove consistency. It is also easy to check that the recovery acceleration is a performance optimization that does not affect the consistency property. So, ignoring acceleration of recovery for simplicity, followers do not execute requests. Thus, we only need to consider the leader's crash and recovery. We assume the client/proxy has committed *req* before the leader crash.

(1) If the client/proxy commits *req* in the fast path, it means at least $f + \lceil f/2 \rceil$ followers have consistent log entries with the leader up to this request *req*. Therefore, on the old leader, all the log entries *before req* are *also* committed, because they also form a super quorum with consistent hashes. So, they can survive crashes and be recovered in Steps 5 and 6 of §A.3. Additionally, consider an uncommitted request *ureq*, which is not commutative to *req* and has a smaller *deadline* than *req*, it cannot be appended by any of the $f + \lceil f/2 \rceil + 1$ replicas which have appended *req*, because the *early-buffer* of DOM only accepts and releases requests in the ascending order of deadlines (§4). Even if all the other $\lfloor f/2 \rfloor$ have appended *ureq*, they fail to satisfy the condition in Step 5/Step 6, so *ureq* cannot appear in the recovered *log*s to affect the execution result of *req*.

(2) If the client/proxy commits *req* in the slow path, it means at least f followers have consistent log entries with the leader up to *req*, i.e., the *deadline*s of their *sync-points* are greater than or equal to the *deadline* of *req*. Therefore, on the old leader, all the log entries before *req* are committed, and they can survive crashes and be recovered in Step 5 of §A.3. Additionally, if the follower's *log* contains the request *ureq*, which is not commutative to *req* and has a smaller deadline than *req*, but does not exist on the leader, then *ureq* cannot appear in the recovery *log* of the new leader. This is because, based on the workflow of the slow path (§6.4), the follower advances its *sync-point* strictly following *sync* messages from the leader. Since the *sync* message does not include the *ureq*'s 3-tuple <*client-id, request-id, deadline*>, the follower will delete *ureq* before updating its *sync-point*. Therefore, it is impossible for *ureq* to appear in the recovered *log*s and affect the execution result of *req*.

After recovery, the survived log entries will be executed by the new leader according to the ascending order of their *deadline*s, thus the same execution order is guaranteed and provides the consistent execution result for *req*.

B.3 Proof of Linearizability

We assume there are two committed requests, denoted as *req-1* and *req-2*. The submission of *req-2* is invoked after the completion of *req-1*, i.e. the client has observed the execution of *req-1* before submitting *req-2*. We want to prove that no contrary observation can occur after crash and recovery. Here we assume *req-1* and *req-2* are not commutative with each other, because the execution of commutative requests cause no effect on each other, regardless of their execution order.

Since *req-2* is invoked after the completion of *req-1*, *req-2* must have a larger *deadline* than *req-1*, otherwise, it cannot be appended to the *log*. Based on the durability property, *req-1* and *req-2* will be both recovered after a crash. According to the recovery algorithm, the new leader still executes the two requests based on their *deadline*s. Therefore, the execution of *req-1* on the new leader cannot observe the effect of *req-2*. By contrast, while executing *req-2*, the effect of *req-1*'s execution has already been reflected in the leader's replica state. Therefore, no contrary observation (i.e., revert or recorder) can occur after the crash and recovery.

C EVALUATION OF NEZHA UNDER DIFFERENT WORKLOADS

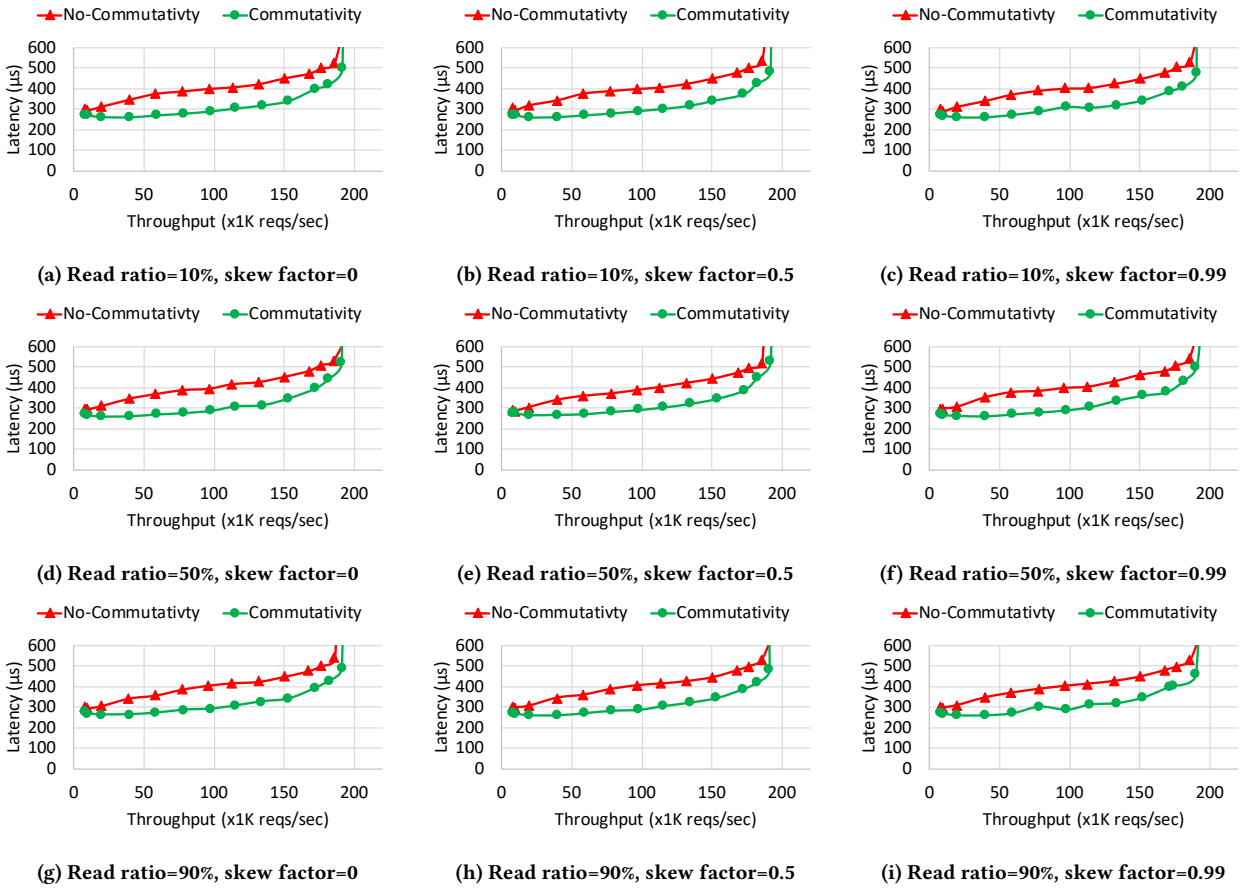
We adopt the similar approach as [60] to conduct extensive evaluation under different workloads: we maintain 1 million unique keys and choose different values of read ratio and skew factors to generate different workloads. As for the read ratio, we choose three different values, i.e. read-10% (write-intensive), read-50% (medium) and read-90% (read-intensive). As for the skew factor, we also choose three different values, i.e. skew-0.0 (evenly distributed), skew-0.5 (medium) and skew-0.99 (highly skewed). The combination of the two dimensions create 9 different workloads. We measure the median latency and throughput under each workload, as shown in Figure 18. Considering the variance in cloud environment, we run each test case for 5 times and plot the average values.

Although the latency in the cloud can vary over time [22, 46, 66] and introduces some noise to performance results, in general, the commutativity optimization remains effective across all workloads and helps reduces the latency by 7.7 %-28.9 %: under low throughput, the effectiveness of the commutativity optimization is not distinct because the No-Commutativity variant can also keep a high fast commit ratio (~75 %). However, as the throughput increases, the fast commit ratio of No-Commutativity variant drops distinctly but the commutativity variant can still maintain a high fast commit ratio (80 %-97 %), so the commutativity optimization becomes more effective. Then, as the throughput continues to grow, it reaches closer to the capacity of the replicas and even overloads the replicas, so the reduction of latency becomes less distinct again and eventually negligible.

D EVALUATION UNDER DIFFERENT CLOCK VARIANCE

D.1 Explanation of Clock Assumption

Nezha depends on clock synchronization for performance but not for correctness. An accurate clock synchronization provides



(a) Read ratio=10%, skew factor=0 (b) Read ratio=10%, skew factor=0.5 (c) Read ratio=10%, skew factor=0.99

(d) Read ratio=50%, skew factor=0 (e) Read ratio=50%, skew factor=0.5 (f) Read ratio=50%, skew factor=0.99

(g) Read ratio=90%, skew factor=0 (h) Read ratio=90%, skew factor=0.5 (i) Read ratio=90%, skew factor=0.99

Figure 18: Latency vs. throughput (open-loop)

favorable conditions for Nezha to achieve high performance. Here "accurate" means the clocks among replicas and proxies are synchronized with a small error bound *in most cases*, but note that Nezha does not assume a deterministic worst-case error bound, which is impractical because Huygens is built atop a probabilistic model (SVM), and the Huygens agents (or other clock synchronization algorithms) can also fail while the consensus protocol is still running without awareness of that.

Besides, Nezha's correctness does not require the assumption of monotonously increasing clock time either. In other words, even if the local clock time goes back and forth (this can happen because Huygens or other clock synchronization algorithms may correct the clocks with some negative offset), Nezha's correctness is still preserved thanks to the entrance condition of the *early-buffer*. Recall in §4, the eligibility check to enter the *early-buffer* is to compare the incoming request's deadline with the deadline of the last released one (rather than the replica's local clock time). Requests in the *early-buffer* are organized with a priority queue and released according to their deadline order. The clock skew can only cause requests to be released prematurely, but the released requests all follow the ascending order of deadlines, therefore, the invariant of uniform ordering is preserved by DOM. Establishing protocol correctness independent of clock skew is desirable and we will show in §F that the other protocol, Domino, loses its correctness due to clock skew (i.e. it can lose committed requests permanently if replica clocks go from large values to small values).

D.2 Quantifying the Effect of Bad Clock Synchronization on Nezha Performance

Although we have not experienced significant clock skew in our evaluation, it is worthwhile to quantify the effect on Nezha performance imposed by different clock synchronization quality. To simplify the discussion below, we consider most VM/server's clocks are synchronized to the reference clock time within a tight bound, whereas the other ones suffer from distinct skew and are not well synchronized with the reference clock time. Thus, we mainly focus on three categories.

(1) The leader replica's clock is badly synchronized with the other VMs.
(2) The follower replica's clock is badly synchronized with the other VMs.
(3) The proxy's clock is badly synchronized with the other VMs.

Method. To create the effect of bad clock synchronization, we choose one or multiple target VMs (i.e. the leader replica, or the follower replica, or the proxies) and inject artificial offsets when the

clock APIs are called on the VM. To be more specific, we generate random offsets based on normal distribution $N(\mu, \sigma)$. For each test case, we choose different mean values (μ) and standard deviation (σ) for the distribution to mimic bad clock synchronization of different degrees. When the clock API is called, instead of returning the clock value, we take an offset sample from the distribution and add it to the clock value, and then return this summed value, to make its clocks faster/slower than the others.

Test Setting. Similar to the setting in §9.2, we set up 3 replicas and 2 proxies, and use 10 open-loop clients to submit at 10K request/sec each, thus yielding a throughput ~100K request/sec. We measure the latency for each test case and study how the latency evolves as the clock synchronization quality varies. We maintain the same parameters for the adaptive latency bound formula (refer to §4). Specifically, the sliding window size is 1000 to calculate the moving median M; $beta = 3$; $D = 200\mu s$. During our tests, we observe the σ_S and σ_R returned by Huygens are both very small, typically $1 - 2\mu s$. We choose 10 different normal distributions (as shown in Figure 19) to mimic bad clock synchronization of different degrees, from the slowest clock to the fastest clock.

For example, $N(-300, 30)$ indicates that the mean value of the normal distribution is -300 µs with a standard deviation of 30 µs. When we choose an offset (typically a large negative value) from this distribution and add it to the clock value, it will make the clock value smaller than the synchronized clock value by hundreds of microseconds, i.e., the clock becomes slower than the other clocks due to the offset we have added.

D.2.1 Bad Clock Synchronization of Leader Replica. As shown in Figure 19a, when the leader's clock fails to be synchronized with the other VMs and goes faster or slower, it will inflate the latency performance of Nezha. Comparing the faster-clock cases and the slower-clock cases, we can see that a slower clock on the leader replica causes more degradation than a faster clock.

When the leader replica has a slower clock, it will accept most requests into its *early-buffer* but keep them for a much longer time. The requests cannot be committed until the leader releases it. Therefore, the slower the leader's clock is, the long latency Nezha will suffer from.

When the leader replica has a faster clock, it will cause two main effects. First, the leader replica will prematurely release requests with large deadlines, causing the subsequent requests unable to be accepted by its *early-buffer*, so the subsequent requests can only be committed in the slow path. Second, the leader will provide overestimated one-way delay (OWD) values and piggyback them to the proxies (recall that the OWD is calculated by using leader's receiving time to subtract the proxies' sending time) and cause the proxies to use large latency bound (i.e. the max of the estimated OWDs from all replicas) for its following requests multicast. However, the second effect is mitigated by DOM-Rs, because we use the clamping function: when the estimated OWD goes beyond the scope of $[0, D]$, it will use D as the estimated value. Therefore, the negative impact due to the leader's slower clock is constrained. The major impact is that more requests can only be committed in the slow path, which can degrade the latency performance, but the degradation is bounded.

D.2.2 Bad Clock Synchronization of Follower Replica. As shown in Figure 19b, similar to the cases in Figure 19a, the follower's bad clock synchronization also inflates the latency. However, the negative impact of the follower's bad clock synchronization is less distinct than leader's bad clock synchronization: both a faster clock and a slower clock of the follower only cause bounded degradation of latency performance.

When the follower has a faster clock, it may prematurely release requests with large deadlines and cause subsequent requests not accepted by the *early-buffer* (similar to the case where the leader has a faster clock). But eventually the request can be committed in the slow path, so the slow-path latency will bound the degradation.

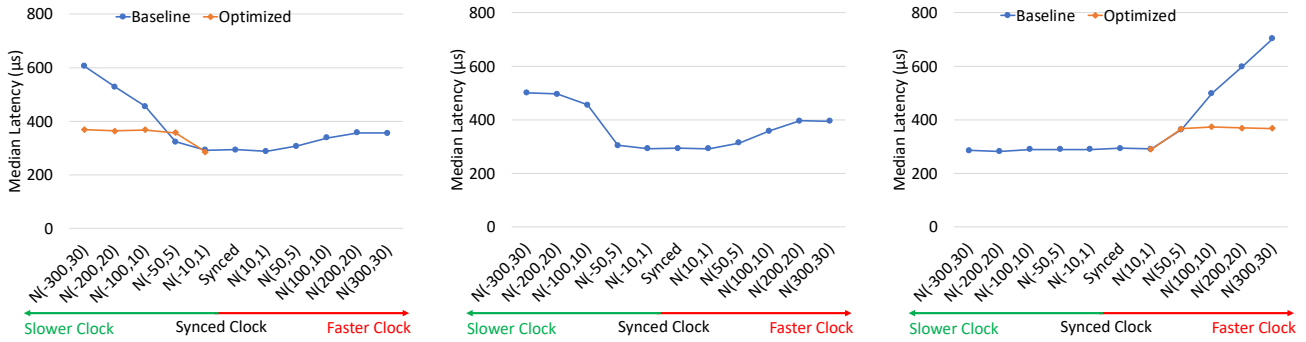
When the follower has a slower clock, it will hold the requests in its *early-buffer* for longer time. However, if the leader and the other follower(s) have well synchronized clocks, they can still form a simple majority to commit the request in the slow path. Therefore, this follower's slower clock can not degrade the latency without bounds (whereas the leader's slower clock can).

The major negative impact caused by the follower's faster/slower clock is that, it will lead to inaccurate estimation of OWDs. If the follower has a faster clock, it will piggyback large OWDs to the proxies, thus causing the proxies to choose large latency bound for the following requests. If the follower has a slower clock, it will piggyback small OWDs (or even negative OWDs) to the proxies. However, thanks to the clamping operation during the latency bound estimation, the latency bound will fall back to D ($D = 200$ µs) when the estimated OWD goes too large or negative. In this way, the negative impact of follower's faster/slower clock is constrained.

D.2.3 Bad Clock Synchronization of Proxy. As shown in Figure 19c, the proxies' having slower clocks do not cause degradation of the latency performance, so long as replicas have well synchronized clocks. However, the proxies' having faster clocks can lead to unbounded degradation of latency performance.

When the proxies have slower clocks, it does not affect the latency so long as replicas are still well synchronized with each other. This is because, although proxies' slower clocks cause smaller sending time, it also leads to larger OWD, which is calculated by the replicas using its local clock time to subtract the sending time. The OWDs are piggybacked to the proxies and eventually lead to large latency bound. Therefore, although the clocks of proxies lag behind, the over-estimated latency bound compensate the lag, and summing up the sending time and latency bound still yields a proper deadline. Therefore, the latency performance does not degrade when proxies have slower clocks.

When the proxies have faster clocks, the latency can go up without bound. When there is only a small clock offset occurring (e.g. $N(10, 1)$), i.e. the proxies' clocks are not fast enough, it will not degrade the latency performance of Nezha. This is because, although the faster clock leads to a larger sending time, it also leads to a smaller latency bound, summing them up still yields a proper deadline. However, when proxies' clocks are too fast (e.g. $N(300, 30)$), the sending time becomes even larger than the receiving time obtained at replicas. In this case, replicas will get negative OWD values, so the estimated OWD will be clamped to D and piggybacked to the proxies. Then proxies will use D as the latency bound. Since the proxies' clocks are already faster than the replicas' clocks, the request deadline will become much larger than



(a) Impact of the leader's clock synchronization quality on Nezha performance

(b) Impact of the follower's clock synchronization quality on Nezha performance

(c) Impact of the proxies' clock synchronization quality on Nezha performance

Figure 19: Nezha latency vs. clock synchronization quality

the replicas' clock time when the request arrives at replicas, leading to long holding delay in the *early-buffer*, and eventually causing much degradation of Nezha's latency performance.

D.2.4 Optimization: Bounding Latency Degradation. Reviewing the cases described in §D.2.1-§D.2.3, we note that the leader's slower clock and proxies' faster clocks can cause unbounded latency performance degradation to Nezha. Although such cases do not affect Nezha's correctness and can hardly be long-lasting in practice (because Huygens will keep monitoring its agents and correct the error bounds), we propose an optimization strategy to bound the latency even when such cases of bad clock synchronization become long-lasting. The key idea of the optimization is to let leader force the request to be committed in the slow path.

Recall in the design of DOM (§4), DOM-R will not accept the request into *early-buffer* only if its deadline is smaller than the last released one, which is not commutative to this request (§8.2). We can enforce the entrance condition of *early-buffer*: If the request's deadline is much larger than the current clock time of the leader, which means the request will suffer from a long holding delay if it is put into the *early-buffer*, then the leader also modifies its deadline to be slightly larger than the last released one and then put it into the *early-buffer*. This step is similar to ③ in Figure 5. The difference is, here we modify a large deadline to a smaller one so as to make it release from the *early-buffer* earlier. By contrast, step ③ in Figure 5 is to modify a small deadline to a larger one, so that it will not violate uniform ordering with previously released requests from the *early-buffer*.

The effectiveness of the optimization is shown in Figure 19a and Figure 19c. We configure a threshold for the leader replica: if the request's deadline is larger than the replica's current clock time by 50 µs, then the request will not be directly put into the *early-buffer* (as the baseline does). Instead, the leader replica modifies the request's deadline to be slightly larger than the deadline of the last released request (which is not commutative to this request), so that the request can enter the leader's *early-buffer* and be released much earlier without violating uniform ordering. Eventually, the request can be committed in the slow path. After installing the optimization strategy, we can see from Figure 19a and Figure 19c that, the degradation of the latency performance becomes bounded, which provides Nezha with stronger resistance to bad clock synchronization.

In this section, we only discuss the three typical cases. Theoretically, there exists some possibility that these cases can happen simultaneously, which creates even more complicated scenarios. For example, when proxies and the leader replica both have slower clocks, the effect due to the bad clock synchronization can be counteracted to some extent. However, the optimization strategy discussed here is still effective to bound the latency degradation and help Nezha to resist the impact of bad clock synchronization.

E DERECHO IN PUBLIC CLOUD

Derecho [24, 55] is a recent high-performance state machine replication system. It works with both RDMA and TCP, and achieves very high throughput with RDMA. Since Derecho is also deployable in public cloud (with TCP), we intend to compare Nezha with Dereco in public cloud.

We follow the guidelines from the Derecho team [1]: First, we try to tune the configuration parameters for Derecho and reproduce the performance number in [55] by using bare-metal machines. We set up a cluster in Cloudlab [13]. We use 3 `c6525-100g` instances (equipped with 100GB RDMA NICs) to form a Derecho subgroup size of 3. We use ramfs [10] as the storage backend for Derecho to avoid disk writes. Then, we evaluate the throughput of Derecho in all-sender mode and one-sender mode. As for the all-sender mode, Derecho yields the throughput of 626K request/sec with 1KB message size and 634K request/sec with 100B message size. As for the one-sender mode, Derecho yields the throughput of 313K request/sec with 1KB message size and 305K request/sec with 100B message size. These numbers are close to the reported number in [55], which convinces us that the configuration parameters have been properly set.

Then, we keep using the cluster and the configuration files for Derecho, but switch the backend from RDMA to TCP. After switching to TCP, we find Derecho's performance drops much: with 100B message size, the all-sender mode achieves the throughput

of 17.4K request/sec with the median latency of 2.33 ms; the one-sender mode achieves the throughput of 5.68K request/sec with the median latency of 2.35 ms. The throughput becomes even lower after we move back to Google Cloud: with 100B message size, the all-sender mode achieves the throughput of 16.5K request/sec with the median latency of 2.0 ms; the one-sender mode achieves the throughput of 4.93K request/sec with the median latency of 2.54 ms.

We speculate that the low performance of Derecho is due to `libfabric` it uses for communication. Although `libfabric` supports both RDMA and TCP communication, it is mainly optimized for RDMA, and the TCP backend is mainly used for test and debug [50]. We expect Derecho can achieve much higher performance if equipped with a better TCP backend. Therefore, we think the comparison is unfair to Derecho and do not include it.

F ERROR TRACES OF DOMINO

Domino [69] is a recently proposed solution to achieve consensus with clock synchronization. When clock skew happens, clients may consider the request as committed, but eventually the request is lost from the replicas, leading to durability violation. The key reason for the durability violation is because clocks cannot always maintain monotonically increasing value [2]. In this section, we will use an error trace to show the durability violation.

Error Trace 1: There are 5 replicas in Domino, and we denote them as R0-R4. Suppose R0 is the DFP (Domino's Fast Paxos) leader and the others are the followers. There are two requests included in the trace, denoted as request-1 and request-2.

(1) R1-R5's clocks are synchronized. R1-R4 report their current clock time T to the coordinator R0, indicating they have accepted no-ops for all log positions before T (as described in 5.3.2 of Domino paper [69]).
(2) R0 receives request-1 with predefined arrival time T+1. So R0 accepts this request.
(3) R0 intends to execute request-1. Before execution, R0 broadcasts request-1 with the other replicas.
(4) R1 and R2 also accept request-1 and reply to R0, whereas R3 and R4 do not receive request-1 from either the client or the replica because the request is dropped.
(5) R0 considers request-1 is committed because it has received the majority of replies (R1, R2 and itself). R0 considers it safe to execute the request, because R1-R4 have reported T to R0, and R1 and R2 also accept request-1.
(6) R0 executes the request, but has not broadcast the execution to learners (i.e. the other replicas).
(7) R1 and R2 fail (i.e. so the NTP services of R1 and R2 also fail). When R1 and R2 are relaunched, the NTP services on their nodes are reinitialized, but the reinitialized NTP gives a time T1, which is smaller than T.
(8) R3 and R4's NTP services encounter a skew and get a clock time T2, which is smaller than T.
(9) The client submits request-2, which has a pre-arrival time smaller than T but larger than both T1 and T2.
(10) R1-R4 all accept request-2 and send replies to the client. The client considers request-2 as committed.
(11) R1-R4 waits for the notification from the coordinator R0, when the notification arrives, R1-R4 will do either (a) replace request-2 with no-op and only execute request-1 or (b) execute both requests but with request-2 first and request-1 second.

The choice between (a) and (b) in Step 11 depends on how Domino implements the coordination between the leader and the other learners (followers), which is not shown in the Domino paper [69]. We have studied the implementation [68] of Domino and found that, followers will choose (a) because DFP leader will also broadcast the log positions (refer to `NonAcceptTime` variable in [68]) which the leader fills no-ops. When followers choose (a), Domino violates durability because request-2, which have been considered committed, is lost permanently. As an alternative, if followers choose (b), consistency will be violated: After the DFP leader (i.e., coordinator) fails and one replica among R1-R4 becomes the new leader, it will have different system state (which executes both request-2 and request-1) from the old leader (which only executes request-1)

Furthermore, the durability property is a necessary condition for the consistency and linearizability properties:

- Because one committed request can affect the execution result of the subsequent requests, the loss of it will lead to different execution results for the subsequent requests, thus violating consistency.
- Because the committed request can be observed by clients, the loss of it causes contrary observation afterwards, thus violating linearizability.

Hence, Error Trace 1 has shown that Domino can violate durability, and consequently violate consistency and linearizability property.

Nezha avoids such error cases. In Nezha's design, the key difference from Domino to avoid such error cases is that, Nezha exploits synchronized clocks to *reduce packet reordering in the network, rather than directly decide ordering with clock time*. The design of the *early-buffer* maintains the invariant of consistent ordering regardless of clock skew/failure, because the eligibility check for the request to enter the *early-buffer* is to compare its deadline with the last released one (§4). Even after the replica fails and recovers, the consistent ordering invariant is still guaranteed: in this case, the *last released request* is the last appended entry in the recovered *log* (Step 9 in §A.3). Therefore, Nezha's correctness is independent of the clock behavior. However, the clock synchronization indeed affects the performance of Nezha. For example, if the clock time of the replica becomes much faster and goes to a very large value, it can release some requests with very large deadlines. The large deadlines will be used in the eligibility check of the *early-buffer*, making the subsequent requests unable to enter the *early-buffer* and trigger more frequent slow path. We have discussed in §D different cases of bad clock synchronization and their impact on Nezha.

By contrast, *Domino directly uses the clock time for ordering*, and does not expect that the clocks can also give a smaller time than before(Step 7 and Step 8 in Error Trace 1), which leads to the incorrectness of the protocol.

G FORMAL COMPARISON OF DIFFERENT PRIMITIVES

Concretely, the mostly ordered multicast (MOM) primitive [53] used by Speculative Paxos creates a network environment to make most requests arrive at all replicas in the same order. The ordered unreliable multicast (OUM) primitive [33] used by NOPaxos ensures ordered delivery of requests without a reliability guarantee using a programmable switch as a request sequencer. By contrast, the deadline-ordered-multicast (DOM) primitive used by Nezha leverages clock synchronization to guarantee consistent ordering, so as to ease the work for replication protocols to achieve state consistency (i.e. to satisfy both consistent ordering and set equality). In this section, we aim to make a formal comparison among the three primitives.

G.1 Notation

- Replicas: $R_1, R_2, \ldots$
- Messages: $M_1, M_2, \ldots$
- $a(M_i, R_k)$: the arrival time of M_i at R_k. It is the **reference time which is not accessible by replicas**, replicas can only get an approximate $\hat{a}(M_i, R_k)$ by calling their local clock API once M_i arrives.
- $r(M_i, R_k)$: the reference time at which M_i is released by the primitive to R_k's protocol. It is the reference time. The primitive does not deliver a reference time $r(M_i, R_k)$ to replication protocols, instead, it delivers an approximate $\hat{r}(M_i, R_k)$ by calling the clock API before releasing M_i.
- $S(M_i)$: the sequential number of M_i given by Sequencer.(OUM Oracle Information).
- $D(M_i)$: the **planned** deadline of M_i to arrive at all replicas (DOM Oracle Information). Replicas know the value of $D(M_i)$ but cannot decide when is exactly the time point of $D(M_i)$ by simply checking their local clock.

G.2 Definition

- Packet drop: M_i is lost to R_x if $r(M_i, R_x) = \infty$
- consistent ordering: R_1 and R_2 are said to be *consistently ordered* (denoted as $UO(R_1, R_2, M_1, M_2)$) with respect to M_1 and M_2 if:
  - $r(M_1, R_1) > r(M_2, R_1)$ and $r(M_1, R_2) > r(M_2, R_2)$
  - Or $r(M_1, R_1) < r(M_2, R_1)$ and $r(M_1, R_2) < r(M_2, R_2)$

  For simplicity, we omit discussing the edge case $r(M_1, R_1) = r(M_1, R_1)$ and/or $r(M_1, R_2) = r(M_1, R_2)$, which can be categorized into either of the two aforementioned outcomes. Similar edge cases are also omitted in the discussion of §G.3.
- Set equality: R_1 and R_2 are *set-equal* with respect to M_1 (denoted as $SE(R_1, R_2, M_1)$) if
  - $r(M_1, R_1) = \infty$ and $r(M_1, R_2) = \infty$
  - Or $r(M_1, R_1) < \infty$ and $r(M_1, R_2) < \infty$

  Set equality is similar to the term *reliable delivery* in NOPaxos [33]. While NOPaxos describes the property from the network perspective, our description is more straightforward by describing it from the replica perspective.
- Consistency: R_1 and R_2 are consistent if

$$\forall M_i, M_j: \quad UO(R_1, R_2, M_i, M_j) \;\&\; SE(R_1, R_2, M_i) \;\&\; SE(R_1, R_2, M_j)$$

  Satisfying both *UO* and *SE* property is equivalent to implementing an atomic broadcast primitive [12], which is as hard as the consensus protocol.

G.3 Primitive Actions

Given a replica R_k, and two messages M_1 and M_2, we can formally describe the actions of the three primitives as follows.

G.3.1 MOM.

$$\begin{aligned} r(M_1, R_k) &= a(M_1, R_k) \\ r(M_2, R_k) &= a(M_2, R_k) \end{aligned} \tag{1}$$

$r(*, *)$ is completely determined by $a(*, *)$ without guaranteeing consistent ordering.

G.3.2 OUM. Without loss of generality, the OUM Oracle gives $S(M_1) < S(M_2)$.

If $a(M_1, R_k) < a(M_2, R_k)$ (**Branch 1**), then

$$\begin{aligned} r(M_1, R_k) &= a(M_1, R_k) \\ r(M_2, R_k) &= a(M_2, R_k) \end{aligned} \tag{2}$$

Otherwise $a(M_1, R_k) > a(M_2, R_k)$ (**Branch 2**), then

$$\begin{aligned} r(M_1, R_k) &= \infty \\ r(M_2, R_k) &= a(M_2, R_k) \end{aligned} \tag{3}$$

Equation 2 captures the case where M_1 and M_2 arrive in an order *consistent* with their sequence numbers.

Equation 3 captures the case where M_1 and M_2 arrive in an order *inconsistent* with their sequence numbers, in which case M_1 is immediately declared lost.

consistent ordering is guaranteed by OUM because messages arrive at different replicas either consistent with their sequence numbers (which are unique to a message and not a replica) or messages are declared lost.

G.3.3 DOM. To simplify the following comparison analysis, we assume the local clock of each replica is monotonically increasing, which is a common assumption in clock modeling [30, 37, 39]. However, it is worth noting that, the correctness of Nezha does not require this assumption: recall that the replicas compare the deadlines of the incoming requests with the *last released request* to decide whether the incoming request can enter the *early-buffer* (§4). Therefore, even if the local clock time goes back and forth, the entrance condition of the *early-buffer* still preserves consistent ordering among the released requests. The clock behavior only affects the performance of Nezha but not its correctness.

DOM can satisfy the monotonically increasing property as follows: DOM tracks the returned value every time it calls the clock API. If the returned value is smaller than the last one (i.e. violating the monotonically increasing property), DOM disposes of the value and retries the clock API. When the replica fails, DOM can rely on the replication protocol to recover the committed logs, and then it starts using the clock time which is larger than the deadline of the last log entry. In this way, DOM guarantees that each replica clock follows monotonically increasing property.

The *monotonically increasing clock time* leads to the following fact:

$$\begin{aligned} r(M_1, R_k) < r(M_2, R_k) &\iff \hat{r}(M_1, R_k) < \hat{r}(M_2, R_k) \\ a(M_1, R_k) < a(M_2, R_k) &\iff \hat{a}(M_1, R_k) < \hat{a}(M_2, R_k) \end{aligned} \tag{4}$$

Without loss of generality, assume DOM Oracle decides two deadlines $D(M_1)$ and $D(M_2)$, satisfying $D(M_1) < D(M_2)$.

If $\hat{a}(M_1, R_k) < \hat{a}(M_2, R_k)$ or $\hat{a}(M_1, R_k) < D(M_2)$ (**Branch 3**), then

$$\begin{aligned} \hat{r}(M_1, R_k) &= \max\{D(M_1), \hat{a}(M_1, R_k)\} \\ \hat{r}(M_2, R_k) &= \max\{D(M_2), \hat{a}(M_2, R_k)\} \end{aligned} \tag{5}$$

Based on the condition and the formula, it is easy to check that $\hat{r}(M_1, R_k) < \hat{r}(M_2, R_k)$, thus $r(M_1, R_k) < r(M_2, R_k)$.

Otherwise $\hat{a}(M_1, R_k) > \hat{a}(M_2, R_k)$ and $\hat{a}(M_1, R_k) > D(M_2)$ (**Branch 4**), then

$$\begin{aligned} \hat{r}(M_1, R_k) &= \infty \\ \hat{r}(M_2, R_k) &= \max\{D(M_2), \hat{a}(M_2, R_k)\} \end{aligned} \tag{6}$$

Equation 5 captures the cases where either M_1 arrives before M_2 arrives, or M_1 arrives before M_2's deadline (based on the local clock of the replica). In both cases, M_1 can be released to the protocol before M_2 is released.

Equation 6 captures what happens when M_1 arrives after both M_2's deadline and M_2's arrival. Here, M_1 has to be declared lost (so that both $\hat{r}(M_1, R_k)$ and $r(M_1, R_k)$ are ∞) and M_2 is released to the protocol.

consistent ordering is guaranteed by DOM because messages are released to replicas according to their deadline's order. Those which have violated the increasing deadline order will not be released by DOM and should be handled by the replication protocol.

G.4 Understanding Difference between MOM, OUM and DOM

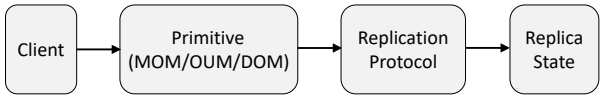


Figure 20: General Model of Speculative Paxos/NOPaxos/Nezha

As shown in Figure 20, primitives are decoupled from the replication protocol. None of the primitives guarantees consistency defined in §G.2. The primitives are just used to create favorable message sequences for the replication protocol to achieve consistency more efficiently.

Considering two replicas R_1 and R_2, with M_1 and M_2 arriving at both replicas. We aim to study the question: When equipped with MOM/OUM/DOM, how likely (easily) can R_1 and R_2 reach consistency without sacrificing liveness (i.e. both M_1 and M_2 should appear in the consistent replica state) given the message sequences output from the primitive?

- MOM simply relies on the highly-engineered network to remove inconsistency flavor, and its output is exactly the output of the network. There is no guarantee on either consistent ordering or set equality. When it comes to the general network (e.g. public cloud), Rarely both $r(M_1, R_1) < r(M_2, R_1)$ and $r(M_1, R_2) < r(M_2, R_2)$ (or the other direction $r(M_1, R_1) > r(M_2, R_1)$ and $r(M_1, R_2) > r(M_2, R_2)$) are satisfied at the same time, thus most consensus work still needs to be undertaken by the replication protocol.
- OUM is potentially better than MOM, because it does the serialization between the clients and replicas with a standalone sequencer, so that the reordering occurrence in the path between clients and the sequencer does not matter. However, when reordering happens in the path between the sequencer and replicas, it leads to **Branch 2**, thus the replication protocol (e.g. NOPaxos) has to handle the loss of M_1 for the affected replicas (e.g. fetching from other replicas or starting gap agreement). Although the consistency property is still satisfied if all replicas take **Branch 2**, that leads to liveness problem: the client which submits M_1 has to retry a new submission. If it goes to the extreme case, when clients submit a series of requests and only the one with the largest sequential number arrives first on all replicas, then all the other requests are declared loss by OUM. In this case, the replicas reach consistency, but little progress is made.
- DOM performs better than OUM in general network because it maintains stronger resistance to reordering. Based on the Equation 4, we can easily derive that **Branch 3** of DOM is a superset of **Branch 1** of OUM. In other words, replicas equipped with DOM are more likely to take DOM's "good" branch (i.e. **Branch 3**), whereas replicas equipped with OUM are less likely to take OUM's "good" branch (i.e. **Branch 1**). However, DOM's strong resistance is obtained at the expense of extra pending delay. According to Equation 6, even when M_1 and M_2 come in order and before their deadlines, they still need to be held until $D(M_1)$ and $D(M_2)$. By contrast, OUM can immediately present M_1 and M_2 to the replication protocol, according to Equation 2.

G.5 Why does Clock Synchronization Matter to DOM?

Clock synchronization affects the effectiveness of DOM for two reasons. First, clock synchronization affects whether DOM can resist the reordering. Second, clock synchronization is closely related to the measurement of client-replica one-way delay, thus (indirectly) affecting whether the client can decide a proper deadline for its messages (requests). We use two cases to illustrate how bad clock synchronization and bad deadlines can affect DOM's effectiveness, and use one case to illustrate the effective DOM with good clock synchronization and proper deadlines.

Bad Case-1: Bad clock synchronization. M_1 and M_2 arrive at R_1 out of order but $a(M_1, R_1) < D(M_2)$. Meanwhile, the two messages arrive at the other replicas in order. If R_1's clock had been well synchronized with the reference clock, $\hat{a}(M_1, R_1)$ should be very close to $a(M_1, R_1)$, leading to $\hat{a}(M_1, R_1) < D(M_2)$, and then DOM should be able to rectify the reordering on R_1, so that it outputs the consistent message sequence as the others. However, R_1's clock fails at that time and gives a very large $\hat{a}(M_2, R_1)$ that leads to $\hat{a}(M_2, R_1) > D(M_3)$. In this case, DOM becomes ineffective and R_1 takes **Branch 4**, leaving more consensus work for the replication protocol to complete.

Bad Case-2: Improper deadline. Suppose the clock synchronization goes wrong on some replicas (e.g. R_2), and the clocks on the problematic replicas are much faster than the reference clock, so the one-way delay (OWD) measurement gives very large value and elevates the latency bound estimation (§4). When M_1 and M_2 are given very large deadlines $D(M_1)$ and $D(M_2)$.

The replicas (e.g. R_1) will take **Branch 3** and DOM is able to rectify possible reordering. However, M_1 suffers from the pending time of $D(M_1) - \hat{r}(M_1, R_1)$ whereas M_2 suffers from the pending time of $D(M_2) - \hat{r}(M_2, R_1)$ on R_1 (Assume R_1's clock is well synchronized with the reference clock).

Good Case: Clocks are well synchronized and $D(M_i)$s are properly decided, i.e. $D(M_i)$ is close to (but slightly larger than) the arrival time $a(M_i, R_x)$ regarding most replicas. In this case, when the network is good, DOM delivers the message to the replication protocol with both consistent ordering and set equality, just like MOM and OUM. More than that, when the network causes message reordering, both MOM and OUM will present the reordering effect to the replication protocol, and triggers the replication protocol to take extra effort. Specifically, MOM presents non-uniformly ordered messages to the replication protocol which causes Speculative Paxos to go to the slow path and costly rollback; OUM presents consistent order messages with gaps (equation 3), which also causes NOPaxos to go to the slow path and make the following messages pending before the gap is resolved. By contrast, so long as the out-of-order message (M_1) does not break the deadline ($D(M_2)$) of the message (M_2), the reordering between M_1 and M_2 can be rectified by DOM in equation 5 and is insensible to the replication protocol, so that the workload of replication protocol (Nezha) is much relieved.

H NEZHA VS. EPAXOS IN WAN

EPaxos [47, 60] is a consensus protocol that proves to outperform Multi-Paxos in Wide Area Netowrk (WAN) scenario. EPaxos fully exploits the fact that Local Area Network (LAN) message delays are negligible when compared with WAN message delays. Therefore, EPaxos distributes its replicas across multiple zones. Such design enjoys two benefits: First, the long-distance (cross-zone) communication between replicas is fully controlled by the service providers, so the service providers can use private backbone network to provide better quality of service. By contrast, if replicas are co-located together and far away from clients. The long distance from clients to replicas is out of control, which may cause longer latency and more frequent message drop. Second, Although EPaxos also incurs 2 RTTs in the fast path, one of them is LAN RTT (i.e. client→replica and replica→client message delays) that can be ignored. Therefore, EPaxos claims to achieve optimal RTT (1 WAN RTT) in the fast path and 2 RTTs in the slow path, which makes it outperform Multi-Paxos in latency. Besides, by using the multi-leader design and commutativity, EPaxos also enjoys less throughput bottleneck compared with Multi-Paxos.

While in this paper we only focus on LAN deployment and have shown that Nezha outperformed TOQ-based EPaxos in LAN (Figure 7), we would like to highlight that Nezha is also deployable in WAN environment, and we believe Nezha can also earn more advantages over EPaxos when deployed in WAN. We analyze the advantages below and leave the experimental evaluation as our follow-up work.

H.1 Latency

When deployed in WAN, Nezha shares the same benefit as EPaxos: Nezha deploys its stateless proxies in every zone, so the client→proxy and proxy→client message delays are also LAN message delays that can be ignored. Therefore, Nezha also achieves 1 WAN RTT as EPaxos, but Nezha achieves only 1.5 WAN RTTs in the slow path, compared with 2 WAN RTTs achieved by EPaxos.

Besides, Nezha earns more performance advantages over EPaxos when there are more zones than replicas. For instance, consider a 3-replica consensus protocol with 10 different zones, and clients are evenly distributed in every zone, EPaxos cannot benefit all clients regarding the latency. Since there are only three replicas, at most the clients in three zones can enjoy 1 WAN RTT to commit their requests in the fast path. The majority of clients (70%) still suffer 2 WAN RTTs to commit in the fast path, and even worse (3 WAN RTTs) to commit in the slow path. The large number of zones makes EPaxos lose most of its latency benefit. To let all clients enjoy 1 WAN RTT fast path, EPaxos has to deploy one replica in each zone (i.e. 10 replicas), but in that case, the quorum check will become much heavier and more interference/conflicts among replicas can occur. In contrast, Nezha *distributes proxies instead of replicas across zones, and proxies are highly scalable.* Regardless of the number of zones, Nezha can still maintain 1 WAN RTT for all clients, so long as sufficient proxies are deployed in every zone.

Besides, when data center failure is not considered (i.e. the number of zone failures is assumed to be 0), Nezha can even co-locate all replicas in the same zone and connect them with high-end communication (e.g. DPDK, RDMA). In this case, inter-replica communication is also LAN message delays, and Nezha can achieve optimal WAN RTT (1 WAN RTT) for both fast path and slow path, which gives Nezha more latency benefit than EPaxos.

H.2 Throughput

While EPaxos uses multiple leaders to mitigate single-leader bottleneck, Nezha adopts an alternative design: Nezha still maintains single leader but offloads most workload to proxies. The proxy brings two major advantages for Nezha regarding the throughput. First, the inter-replica communication is much more lightweight because the leader only multicast index messages (rather than request messages) to other followers, which have much smaller sizes than requests and can be batched to amortize the communication cost. Second, replicas do not undertake quorum checks, and proxies can conduct the quorum check concurrently. Although EPaxos can share the workload of request multicast and quorum check among replicas, the number of replicas is limited and it is still likely that the quorum check workload can overwhelm the capacity of multiple leaders. However, the number of proxies in Nezha can be considered without constraint (i.e. as many as Huygens can support), and Nezha can deploy as many proxies as needed to tackle the workload of request multicast and quorum check. Therefore, we expect Nezha can also achieve higher throughput than EPaxos.

H.3 Clock Synchronization in WAN

As mentioned in our paper, the performance of Nezha is closely related to the synchronization performance of clocks. A reasonable concern about deploying Nezha in WAN is that the clock error can become very large and cause performance degradation. Such concerns prove to be unnecessary. According to the discussion with the developer team of Huygens, when deployed in public cloud across multiple data centers, the clock accuracy provided

by Huygens will be in the order of 10s of microseconds, with occasional spikes if the WAN link is unstable. Such claims have been verified in [60], which evaluates Huygens in the WAN setting and observes the clock offsets between 20 µs and 1 ms. Considering the inter-datacenter latency is usually tens of or even hundreds of milliseconds (as shown in Figure 5 of [60]), the synchronization performance of Huygens is sufficient for Nezha to achieve fast consensus in WAN.

I NEZHA TLA+ SPECIFICATION

The TLA+ specification of Nezha is available at the anonymous repository https://gitlab.com/steamgjk/nezhav2/-/blob/main/docs/Nezha.tla.