Sync HotStuff: Simple and Practical Synchronous State Machine Replication

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***Abstract*—Synchronous solutions for Byzantine Fault Tolerance (BFT) can tolerate up to minority faults. In this work, we present Sync HotStuff, a surprisingly simple and intuitive synchronous BFT solution that achieves consensus with a latency of 2Δ in the steady state (where Δ is a synchronous message delay upper bound). In addition, Sync HotStuff ensures safety in a weaker synchronous model in which the synchrony assumption does not have to hold for all replicas all the time. Moreover, Sync HotStuff has responsiveness, i.e., it advances at network speed when less than one-quarter of the replicas are not responding, a small sacrifice in availability compared with optimal partially synchronous solutions. Borrowing from practical partially synchronous BFT solutions, Sync HotStuff has a two-phase leader-based structure, and has been fully prototyped under the standard synchronous message delay assumption. When tolerating a single fault, Sync HotStuff achieves a throughput of over 280 Kops/sec under typical network performance, which is comparable to the best known partially synchronous solution and more than a factor of two better than the state-of-the-art synchronous solution.**

I. Introduction

Byzantine Fault Tolerance (BFT) protocols relying on a synchrony assumption have the advantage of tolerating up to one-half Byzantine faults [1], while asynchronous or partially synchronous protocols tolerate only one-third [2]. On the flip side, synchronous protocols are often considered impractical for three main reasons. First, existing synchronous protocols require a large number of rounds. Second, most synchronous protocols require lock-step execution (i.e., replicas must start and end each round at the same time), making them hard to implement and further exacerbating the latency problem. Third, an adversary may attack the network to violate the synchrony assumption, causing the protocol to be unsafe.

In this work, we introduce Sync HotStuff, a synchronous BFT state machine replication (SMR) protocol that addresses the above concerns with a surprisingly simple and intuitive solution (see Figure 1). In Sync HotStuff, in the standard synchrony model, a leader broadcasts a proposal; the replicas echo it; and each replica can commit after waiting for the maximum round-trip delay unless it hears by that time an equivocating proposal signed by the leader. (If the leader does not propose, replicas time out and perform a leader change; details on this step are given in the body of the paper.)

*Part of the work was done while the authors were at VMware Research.

Simple yet powerful, Sync HotStuff achieves the following desirable properties. First, as in most synchronous solutions, Sync HotStuff tolerates up to one-half Byzantine replicas. Second, inspired by Hanke et al. [3], Sync HotStuff does not require lock-step execution in the steady state. Third, with minor modifications, Sync HotStuff can handle a weaker and more realistic synchrony model suggested by Chan et al. [4]. Finally, Sync HotStuff is prototyped and shown to offer practical performance. It achieves a throughput comparable to partially synchronous protocols and the commit latency is roughly a single maximum round-trip delay. Given the above properties, we believe Sync HotStuff can be the protocol of choice for single-datacenter replicated services as well as consortium blockchain applications.

We proceed to elaborate on the key techniques and key results of Sync HotStuff, which removes performance barriers on synchronous BFT under weaker assumptions.

Near-optimal latency. The first key contribution is the aforementioned extremely simple steady state protocol (Figure 1). We observe that waiting for a single maximum round-trip delay suffices for replicas to commit. Furthermore, our protocol does not have to be executed in a lock-step fashion, despite relying on synchrony. In other words, other than the concurrent waiting step, replicas move to the next step upon receiving enough messages of the previous step, without waiting for the conservative synchrony delay bound. This gives a latency of $2\Delta + O(\delta)$ in steady state where Δ denotes the known bound assumed on maximal network transmission delay and δ denotes the actual network delay, which can be much smaller than Δ.

Assuming $\delta \ll \Delta$, the above latency is within a factor of two of the optimal latency that can be obtained by synchronous protocols: we give a minor adaptation to the proof of Dwork et al. [2] to show that a Δ latency is necessary for any protocol tolerating more than one-third Byzantine faults. The Δ latency lower bound should not be surprising because a protocol that commits faster than Δ, in a way, does not take advantage of synchrony and is thus subject to the one-third partial synchrony barrier. In fact, we conjecture a stronger latency lower bound of 2Δ. Our intuition is that replicas can be out-of-sync by Δ, so one Δ is needed for lagging replicas to

catch up and another Δ is needed for messages to be delivered (the current lower bound proof only captures the latter Δ).

Moreover, though $O(\delta)$ latency is impossible to *guarantee* under more than one-third faults, it can be achieved optimistically. The Thunderella protocol [5] achieves $O(\delta)$ latency when the leader is honest and more than three-quarter of the replicas are responding. We show that our protocol can be adapted to incorporate this idea.

Practical throughput. The key technique to improve throughput is to move the *synchronous waiting steps* off the critical path. In more detail, there are two steps in our protocol that require waiting for a conservative $O(\Delta)$ time. One is to check for a leader equivocation before committing and it is made concurrent to main logic. Thus, replicas start working on the next block without waiting for the previous blocks to be committed. (The non-blocking commit also reduces latency, since a block can now be proposed before the previous block is committed.) The other is before sending a status message to the leader, but it is in the view-change protocol, which occurs infrequently. With these tricks, in the steady state, the replicas are always sending protocol messages and utilizing the entire network capacity, thus behaving exactly like partially synchronous protocols. Our experiments validate that Sync HotStuff achieves throughput comparable to partially synchronous protocols. In fact, since a synchronous solution tolerates more corruption (half vs. one-third), it requires fewer replicas to be deployed to tolerate a given number of faults. In our experiments, we observe that in some cases, Sync HotStuff can even slightly outperform partially synchronous solutions in throughput.

Safety despite some sluggish honest replicas. Synchronous protocols proven secure under the standard synchrony assumption fail to provide safety if a single message between honest replicas is delayed. Recently, Guo et al. [6] proposed a "weak synchrony" model that allows the message delay bound Δ, at any point in time, to be violated for a set of honest replicas. We call these replicas *sluggish*. We call the remaining honest replicas *prompt* and messages of *prompt* replicas can reach each other within Δ time. To reflect reality and be more conservative, the model allows sluggish replicas to be arbitrarily *mobile*, i.e., an adversary decides which replicas are sluggish at any time. Messages sent by or sent to a sluggish replica may take an arbitrarily long time *until* the replica is prompt again. Since "weak synchrony" has been used in the literature to describe other models (e.g., [7], [8], [9]), we will refer to this model as the *mobile sluggish* model in this paper. We call the synchrony mode without mobile sluggish faults *standard synchrony*.

With standard synchrony, if a replica sends a message to another replica, it is guaranteed to arrive within Δ time. Our protocol and proofs crucially use this fact to achieve safety. With a mobile sluggish fault model, on the other hand, the delivery is not guaranteed if the sender or the receiver is sluggish. In that sense, the guarantee for a single replica sending or receiving a message is similar to that of partially synchronous network model. The central observation enabling us to tackle mobile sluggish faults is the following: assuming a minority of the replicas are sluggish or Byzantine at any point in time, if a replica receives a message from a majority of replicas, at least one of the senders must be prompt and honest. We use this observation atop Sync-HotStuff-under-standard-synchrony to obtain a protocol in the mobile sluggish model.

The above technique allows Sync HotStuff to ensure safety as long as the number of sluggish plus Byzantine faults combined is less than one-half; in other words, at any time, a majority of replicas must be honest and prompt (which has been shown to be a necessary condition [6] in the mobile sluggish model). More importantly, with the above technique, Sync HotStuff manages to handle the mobile sluggish model at almost no extra cost: the communication complexity remains the same and the commit latency increases only by $O(\delta)$.

Organization. In the remainder of this section, we define state machine replication. In Section II, we describe Sync HotStuff in the standard synchrony model without sluggish faults. Section III augments this protocol tolerate sluggish faults. Section IV adds an optimistically responsive track to Sync HotStuff with sluggish faults. Section V presents the results based on our implementation and evaluation. Section VI compares with closely related works.

A. Definitions and Model

State Machine Replication (SMR). A state machine replication protocol is used for building a fault-tolerant service to process client requests. The service consists of n replicas, up to f of which may be faulty. The service commits client requests into a linearizable log and produces a consistent view of the log akin to a single non-faulty server. More formally, a state machine replication service provides the following two guarantees:

> **(safety)** if a value is committed by some non-faulty replica at a log position, the value committed at the same position by any non-faulty replicas is also the same,
> **(liveness)** each client request is eventually committed by all non-faulty replicas.

We assume that the network consists of pairwise, authenticated communication channels between replicas. We assume digital signatures and a public-key infrastructure (PKI), and use $\langle x \rangle_p$ to denote a message x signed by replica p. (It is sufficient to sign the hash digest of a message for efficiency.) Wherever it is clear from context, we do not mention the subscript p. We also assume that there is no drift between the clocks used by the replicas, i.e., the clocks run at the same rate. Our protocol is secure under a sluggish mobile adversary. However, for ease of exposition we first explain the protocol in the standard synchrony model. We describe these models in the respective sections.

II. Sync HotStuff under Standard Synchrony

We first present Sync HotStuff in the standard synchrony model (without mobile sluggish faults). Here, the synchrony

assumption states that a message sent at time t by any replica arrives at another replica by time $t + \Delta$ where Δ is a known maximum network delay. We use δ to denote the actual message delay in the network. We show a protocol that tolerates minority Byzantine replicas, i.e., $n = 2f + 1$.

Sync HotStuff takes the Paxos/PBFT's approach of having a *stable leader* in a steady state. The reign of a leader is called a *view*, numbered by integers. The leader of view v can simply be replica $(v \mod n)$, i.e., leaders are scheduled in a round-robin order. The leader is expected to keep making progress by committing client requests at increasing *heights*. If replicas detect Byzantine behavior by the leader or lack of progress in a view, they *blame* the leader and engage in a *view-change* protocol to replace the leader. Figures 1 and 2 describe the steady state and view-change protocols, respectively.

Blocks and block format. As commonly done in SMR, client requests are batched into *blocks*. The protocol forms a chain of blocks. We refer to a block's position in the chain as its *height*. A block B_k at height k has the following format

$$B_k := (b_k, h_{k-1})$$

where b_k denotes a proposed value at height k and $h_{k-1} := H(B_{k-1})$ is a hash digest of the predecessor block. The first block $B_1 = (b_1, \perp)$ has no predecessor. Every subsequent block B_k must specify a predecessor block B_{k-1} by including a hash of it. We say a block is *valid* if (i) its predecessor is valid or $\perp$, and (ii) its proposed value meets application-level validity conditions and is consistent with its chain of ancestors (e.g., does not double spend a transaction in one of its ancestor blocks).

Block extension and equivocation. If a block B_k is an ancestor of another B_l ($l \geq k$), we say B_l *extends* B_k. We say two blocks B_l and $B'_{l'}$ proposed in the same view *equivocate* one another if they are not equal and do not extend one another.

Certificates and certified blocks. A key ingredient of BFT solutions is a *quorum certificate* (certificate for short), a set of signatures on a block by a quorum of replicas. In Sync HotStuff, a quorum consists of $f + 1$ replicas (out of $2f + 1$). We use $\mathcal{C}_v(B_k)$ to denote a quorum of signatures on $h_k = H(B_k)$ from the same view v and call it a certificate for B_k.

Certificate ranking. Certified blocks are ranked first by views and then by heights, i.e., (i) blocks with higher views have higher ranks, and (ii) for blocks from the same view, blocks with higher height have higher rank. During the protocol execution, each replica keeps track of all signatures for all blocks and keeps updating the highest certified block to its knowledge. Looking ahead, the notion of highest certified blocks will be used to guard the safety of a commit.

Block chaining. Blocks across heights are chained by hashes (cf. block format) and certificates (cf. Figure 1). This idea originated from the Bitcoin white paper [10] and it was incorporated into BFT by Casper [11] and HotStuff [12], [13]. It greatly simplifies BFT protocols since now the voting step on a block also serves as a voting step for all its ancestor blocks that have not been committed. Hence, crucially, committing a block commits all its ancestors.

A. Steady State Protocol

The steady state protocol runs in iterations. In each iteration, the leader L proposes a block by sending $\langle \mathsf{propose}, B_k, v, \mathcal{C}_{v'}(B_{k-1}) \rangle_L$ where $B_k := (b_k, h_{k-1})$ is a new block, $v' \leq v$, and B_{k-1} is the highest certified block known to L (Step 1). If the leader has been in the steady state, it should extend the previous block it has proposed in the same view. If a leader has just entered the steady state after a view-change, it should extend the highest certified block it learned during the view-change. If there are multiple such blocks (which implies that a previous Byzantine leader has equivocated), the leader can pick one arbitrarily.

Each replica r, upon receiving a valid block B_k from the leader, broadcasts a vote $\langle \mathsf{vote}, B_k, v \rangle_r$ for it if it extends a highest certified block known to r (Step 2). Although the protocol assumes synchrony, replicas do not progress in lock-steps. A replica r may first hear a block through another replica's vote. A vote for a block can thus be considered a *re-proposal* of the block. In this case, if the block extends the highest certified block known to r, replica r also broadcasts a vote for B_k.

Once replica r votes for B_k, it starts a timer called $\mathsf{commit\text{-}timer}_k$ (Step 3). B_k is committed if r does not observe any equivocating blocks within the next 2Δ time (recall that equivocating blocks can be from different heights but are from the same view). We note again that blocks across heights form a chain, and committing a block commits all its ancestors.

Note that the commit timers do not affect the critical path of progress. A replica starts the next iteration concurrently without waiting for the previous height to be committed. In fact, a replica can potentially have many previous heights whose commit timers are still running.

Why does absence of an equivocating block within 2Δ time suffice to ensure safety? Consider an honest replica r that votes for a block B_k at time t, does not observe any equivocating block, and hence commits B_k at time $t + 2\Delta$. We will show that B_k will be the only certified block at height k. For that, we need to show that (i) B_k will be certified, and (ii) no conflicting height-k block can be certified.

r's vote reaches all honest replicas before $t + \Delta$. Recall that honest replicas vote for the first valid block received at each height in a view, and that they vote on blocks obtained through votes (i.e., votes serve as re-proposals). Thus, if any honest replica votes for a conflicting block B'_k, it must do so before $t + \Delta$; otherwise, it would receive B_k from r first and vote for B_k. Therefore, if any honest replica votes for a conflicting block B'_k, r would have known about B'_k before $t + 2\Delta$. This contradicts with the fact that r commits B_k. Hence, (ii) holds. On the other hand, if r does not receive a vote for an equivocating block before $t + 2\Delta$, r can be sure

Let v be the current view number and replica L be the leader of the current view. The steady state protocol runs the following steps in iterations.

1) **Propose.** The leader L broadcasts $\langle \mathsf{propose}, B_k, v, \mathcal{C}_{v'}(B_{k-1})\rangle_L$ where $B_k = (b_k, h_{k-1})$, b_k is a batch of new client requests, $v' \leq v$, and B_{k-1} is the highest certified block known to L.
2) **Vote.** Upon receiving the first valid proposal for a height-k block B_k from L or through a vote by some other replica, if it is extending a highest certified block, forward the proposal to all other replicas and broadcast a vote in the form of $\langle \mathsf{vote}, B_k, v\rangle$. Set $\mathsf{commit\text{-}timer}_k$ to 2Δ and start counting down.
3) **(Non-blocking) Commit.** When $\mathsf{commit\text{-}timer}_k$ reaches 0, commit B_k and all its ancestors.

Fig. 1: The steady state protocol under standard synchrony.

Let L and L' be the leader of view v and $v+1$, respectively.

i **Blame.** If less than p blocks are received from L in $(2p+1)\Delta$ time in view v, broadcast $\langle \mathsf{blame}, v\rangle$. If a leader L proposes equivocating blocks, broadcast $\langle \mathsf{blame}, v\rangle$ and the two equivocating blocks.

ii **Quit old view.** Upon gathering $f+1$ $\langle \mathsf{blame}, v\rangle$ messages, broadcast them, and quit view v (abort all commit-timer(s) and stop voting in view v).

iii **Status.** Wait for Δ time and enter view $v+1$. Upon entering view $v+1$, send a highest certified block to L' and transition back to steady state.

Fig. 2: The view-change protocol under standard synchrony.

that all honest replicas have voted for B_k before $t+\Delta$, and that it will receive all these votes, which form a certificate for B_k, before $t+2\Delta$. Hence, (i) holds. Note that (i) holds even if the leader did not propose B_k to all replicas. In summary, if a block B_k is committed by some honest replica, it will be the only certified block for that height.

Skipped commit. A commit by some honest replica at height k does not imply a commit by all honest replicas at that height. This is because a Byzantine leader can send an equivocating block to a subset of honest replicas before their commit timers expires, causing them to not commit. But the view-change protocol will ensure that all honest replicas eventually commit the same block, as will be explained in the next subsection.

B. View-change Protocol

The view-change protocol ensures liveness. The leader can prevent progress through two mechanisms – stalling and equivocating. The two blame conditions, based on no progress and equivocation, defend against these two attacks, respectively. A leader is expected to propose a block every 2Δ time: one Δ for its proposal to reach other replicas and one Δ for other replicas' votes to arrive. An extra Δ time is given to the leader in the beginning of the view to receive status. This forces a Byzantine leader to propose a block every 2Δ time to avoid being overthrown. If a leader equivocates by proposing conflicting blocks, the two equivocating blocks serve as a proof of Byzantine behavior and are sent together with the blame message. All honest replicas will blame the leader within Δ time. Note that equivocating blocks may be received directly from L or through a vote from other replicas.

If a replica gathers $f+1$ $\langle \mathsf{blame}, v\rangle$ messages, it starts a view-change process by broadcasting the $f+1$ blame messages (Step ii). At this point, it quitsview v, which involves stopping all commit timers and stopping votes in view v.

The replica then waits for Δ time to learn the highest certified block held among honest replicas, and then reports one such block to L' (Step iii). The Δ wait before sending status ensures that an honest replica learns all blocks committed by all honest replicas in previous views before sending status in a new view (formalized in Lemma 1). After that, a replica transitions back to steady state, and waits for the leader to propose a block that extends the block it reports in status.

C. Safety and Liveness

We say a block B_k is committed *directly* if an honest replica commits it by observing no equivocating block within 2Δ after it votes. We say a block B_k is committed *indirectly* if it is a result of directly committing a block extending B_k.

Lemma 1. *If an honest replica directly commits a block B_l in view v, then (i) every honest replica votes for B_l in that view, and (ii) every honest replica receives $\mathcal{C}_v(B_l)$ before entering the next view.*

Proof. Suppose an honest replica r directly commits B_l at time t. r votes for B_l at time $t-2\Delta$. This vote reaches all honest replicas by time $t-\Delta$. At this point, an honest replica r' will vote for B_l unless one of the two conditions below holds: (1) r' has already voted for a block equivocating with B_l before time $t-\Delta$, or (2) r' has already received (and sent) $f+1$ blame messages before time $t-\Delta$. However, if either happens, r would have received either a vote for an equivocating block or $f+1$ blame messages before time t, which contradicts the hypothesis of r committing B_l at time t. Thus, at time $t-\Delta$, all honest replicas are still in the view and

they all vote for B_l. Moreover, all of them enter the next view at or after time t, and by then, they all receive $\mathcal{C}_v(B_l)$. □

Lemma 2 (Unique Extensibility)**.** *If an honest replica directly commits a block B_l in view v, then (i) there does not exist a certificate for block $B'_l \neq B_l$ in the same view, (ii) all certified blocks with higher ranks (i.e., higher views and/or higher heights) extend B_l.*

Proof. Suppose an honest replica r directly commits a block B_l in view v. For the first part, by part (i) of Lemma 1, if an honest replica directly commits a block B_l in view v, all honest replicas vote for B_l in view v; thus, a conflicting height-l certificate cannot be formed in view v.

We will prove the second part through an induction on the view number. For the base case, observe that a block $B'_{l'}$ with $l' \geq l$ certified in view v must extend B_l. If not, $B'_{l'}$ equivocates B_l, and will not get any honest vote by Lemma 1. For the inductive step, observe that by Lemma 1, every honest replica receives $\mathcal{C}_v(B_l)$ before entering the next view. Due to the inductive hypothesis, any certified block with a higher rank from view v to v' extends B_l. Thus, before entering view $v'+1$, all honest replicas lock on block that extends B_l. Thus, any certified block from view $v'+1$ extends B_l. □

Theorem 3 (Safety)**.** *Honest replicas always commit the same block B_k for each height k.*

Proof. Suppose for contradiction that two distinct blocks B_k and B'_k are committed at height k. Suppose B_k is committed as a result of B_l being directly committed in view v and B'_k is committed as a result of $B'_{l'}$ being directly committed in view v'. This implies B_l extends B_k and $B'_{l'}$ extends B'_k. Without loss of generality, assume $v \leq v'$. If $v = v'$ and $l = l'$, then by part (i) of Lemma 2, $B_l = B'_{l'}$. If $v = v'$ and $l < l'$, or $v < v'$, then by part (ii) of Lemma 2, $B'_{l'}$ extends B_l. In either case, $B'_k = B_k$. □

Theorem 4 (Liveness)**.** *(i) A view-change will not happen if the current leader is honest; (ii) A Byzantine leader must propose p blocks in $(2p+1)\Delta$ time to avoid a view-change; and (iii) If k is the highest height at which some honest replica has committed a block in view v, then leaders in subsequent views must propose blocks at heights higher than k.*

Proof. For (i), note that an honest leader L is able to propose p blocks in $(2p+1)\Delta$ time. Immediately after entering its view, L needs Δ time to gather status; after that, it can propose a block every 2Δ time: one Δ for its proposed block to reach all replicas and another Δ for other replicas' votes to arrive. Thus, an honest leader has sufficient time and does not equivocate, so it will not be blamed by any other honest replica. On the other hand, if a Byzantine leader delays beyond the above allotted time it will be blamed by all honest replicas.

For part (iii), observe that all honest receive $\mathcal{C}_v(B_k)$ due to Lemma 1. Hence, in `status` of subsequent views, they all report a certified block at height at least k. If the leader does not propose a block with a height higher than k within 3Δ, it will be blamed by all honest replicas. □

D. Efficiency Analysis

Throughput. In steady state (Figure 1), the only step that uses the synchrony bound Δ is the commit step. But as we have mentioned, the commit step is not on the critical path (non-blocking). Thus, the choice of Δ, no matter how conservative, does not affect the protocol's throughput in steady state. Thus, Sync HotStuff should have similar throughput as partially synchronous protocols. Our experiments in Section V confirm this.

Latency. From an honest leader's perspective, each block incurs a latency of $2\Delta+\delta$ after being proposed. (Step 1 proceeds at the actual network delay δ.) But for SMR, it is more customary to calculate latency from a client's perspective, that is, the time difference between when a client sends a request and when it receives a response. If a client's request arrives between two leader proposals, it incurs an additional delay before being getting proposed. From a client's perspective, it takes δ to send its request to the replicas; on average, this request will arrive half way between two leader proposals, which are 2δ time apart, so the average queue-up delay is δ; it takes an additional δ time for replicas to reply to the client. So the average client latency of Sync HotStuff is $2\Delta + 4\delta$. Our experiments in Section V confirm this.

For comparison, the best prior synchronous protocol in terms of latency is Hanke et al. [3]. Its average latency is $8\Delta + 9\delta$ from a leader's perspective [14], and $9\Delta + 11.5\delta$ from a client's perspective, following a similar analysis.

E. Bound on Responsiveness

Sync HotStuff commits with a 2Δ latency in the steady state when $f < n/2$. For completeness, in this section, we show a lower bound on the latency when $f > n/3$. The lower bound and the proof closely follow Dwork et al. [2]. For clarity, we present the bound in the Byzantine broadcast formulation. Recall that in Byzantine broadcast, a designated *sender* tries to broadcast a value to n parties. A solution needs to satisfy three requirements:

> **(termination)** all honest parties eventually commit,
> **(agreement)** all honest parties commit on the same value, and
> **(validity)** if the sender is honest, then all honest parties commit on the value it broadcasts.

Theorem 5. *There exists no Byzantine broadcast protocol that simultaneously satisfy the following:*

- *termination, agreement and validity as defined above;*
- *tolerates $f \geq n/3$ Byzantine faults;*
- *terminates in less than Δ time if the designated sender is honest.*

Proof. Suppose such a protocol exists. Divide parties into three groups P, Q and R, each of size f. Parties in P and Q are honest and parties in R are Byzantine. Consider the following three scenarios. In Scenario A, parties in R remain silent and an honest designated sender sends 0. In this scenario,

parties in P and Q commit 0 in less than Δ time. In Scenario B, parties in R remain silent and an honest designated sender sends 1. In this scenario, parties in P and Q commit 1 in less than Δ time. In Scenario C, the designated sender is Byzantine and sends 0 to P and 1 to Q; parties in R behave like Q in Scenario A to P, and behave like P in scenario B to Q. Messages from P take Δ to reach Q and messages from Q take Δ to reach P. Before time Δ, P receive no messages from Q and Q receive no messages from P, and thus the scenario is indistinguishable from Scenario A to P and indistinguishable from Scenario B to Q. Thus, P commits 0 and Q commits 1 in less than Δ time, violating agreement. □

III. Sync HotStuff with Mobile Sluggish Faults

The standard synchrony model used in the previous section requires that every message sent by an honest replica arrives at every other honest replica within Δ time. In practice, such an assumption may not hold all the time due to potential unforeseen aberrations in the network at either the sender or the receiver, causing some messages to be delayed. Under such aberrations, a protocol proved secure under the standard synchrony assumption may lose safety. For our protocol specifically, if a replica that voted for an equivocating block runs into a network glitch, then another honest replica may not receive it in time and may incorrectly commit another block. A potential way to "fix" this is to account for the sender (or receiver) of the delayed message as Byzantine and thus tolerate fewer actual Byzantine faults. Unfortunately, over the course of a long execution, every replica is bound to observe such an aberration and this "fix" will result in a dishonest majority of replicas, thus breaking safety eventually.

A. The Mobile Sluggish Model

Chan et al. [4] consider a weaker model that allows some replicas to be *sluggish*, i.e., allows delays for messages sent/received by sluggish replicas in the network. On the other hand, messages sent by *prompt* replicas will respect the synchrony bound. More specifically, if a replica r_1 is prompt at time t_1, then any message sent by r_1 at time $\leq t_1$ will arrive at a replica r_2 prompt at time t_2 if $t_2 \geq t_1 + \Delta$. Moreover, the set of sluggish replicas can arbitrarily change at every instant of time. Hence, we call this model the mobile sluggish fault model. We denote the number of sluggish replicas by d, the number of Byzantine replicas by b and the total number of faults by f. Thus, $f = d + b$.

We note that the mobile sluggish model expects that a message sent by a sluggish replica would respect the synchrony bound as soon as it becomes prompt. In practice, this model captures temporary loss in network connectivity causing message delays. The replica can resend messages as soon as network connectivity is restored. A similar argument holds for receiving messages. However, it is not a good model for capturing a replica going offline for too long since this would require the replica to either buffer a huge amount of messages to be resent or to resend each message many times, both of which are impractical.

B. Protocol

Under weak synchrony, Guo et al. [6] show that no protocol can tolerate a total number of faults (sluggish plus Byzantine replicas) greater than $n/2$. The intuition is that a majority set consisting of Byzantine and sluggish replicas might reach a commit decision without interacting with the rest of the world and might cause conflicting commits. Thus, we assume $> n/2$ replicas are honest and prompt at any time. Moreover, the protocol guarantees liveness when all honest replicas are prompt for a "sufficiently long" period of time, i.e., there can be sluggish replicas but they are not mobile. The duration is directly related to the time required to commit a block in the protocol.

Protocol intuition. In the synchronous protocol described in Section II, the 2Δ period immediately after a vote was used to ensure the following. If no equivocation was observed, i.e., if the 2Δ period was undisturbed, then a vote for a block B_k by a replica r in view v ensured that both, the replica r and all other replicas receive $\mathcal{C}_v(B_k)$ within 2Δ. This is because no other honest replica would have voted for an equivocating block and thus all honest replicas would vote for B_k. Thus, no other certificate would have been formed or committed. On the other hand, the presence of an equivocation implies the potential existence of another certificate, preventing replica r from committing.

In the presence of sluggish replicas, unfortunately, none of the above arguments hold. A sluggish replica, even in the absence of an equivocation, may not receive a certificate in the 2Δ period. Similarly, other replicas may not receive its votes and consequently certificates. Finally, if an equivocation exists, the replica may not receive it in time.

Intuitively, these arguments fail in the mobile sluggish model because we rely on a single replica's vote to ensure safety but this replica can now be sluggish. The natural fix is to rely on $\geq d + 1$ honest replicas' votes. Then, at least one of the votes would have been sent by a prompt replica and it would ensure safety. Unfortunately, if a replica receives $d+1$ votes, it cannot determine if the votes are received from $d+1$ prompt replicas or a combination of Byzantine and sluggish replicas. Thus, we rely on a certificate $\mathcal{C}_v(B_k)$ which is bound to contain one vote from an honest and prompt replica which is capable of sending it to other prompt replicas.

The above intuition suffices for a certificate to be formed at replica r; it knows it can obtain a certificate because it has already received the certificate. However, it doesn't guarantee a certificate at other honest replicas. This can cause safety violations if at the end of the view, sufficient honest replicas do not have a certificate and some other value is proposed and committed in subsequent views. We solve this by using the same intuition again — if a replica has received a $\mathcal{C}_v(\mathcal{C}_v(B_k))$, then $\mathcal{C}_v(B_k)$ must have been voted by one of the prompt replicas. This ensures a certificate at all prompt replicas.

Finally, observe that even if a replica observes an undisturbed 2Δ period after receiving $\mathcal{C}_v(\mathcal{C}_v(B_k))$, then the commit is not guaranteed. This is because the replica could be sluggish

Let v be the current view number and replica L be the leader of the current view. The steady state protocol runs the following steps in iterations.

1) **Propose.** The leader L broadcasts $\langle \mathsf{propose}, B_k, v, \mathcal{C}_{v'}(B_{k-1})\rangle_L$ where $B_k = (b_k, h_{k-1})$, b_k is a batch of new client requests, $v' \leq v$, and B_{k-1} is the highest certified block known to L.
2) **Vote.** Upon receiving the first valid proposal for a height-k block B_k from L or through a vote by some other replica, if it is extending a highest certified block, forward the proposal to all other replicas and broadcast a vote in the form of $\langle \mathsf{vote}, B_k, v\rangle$. Set $\mathsf{pre\text{-}commit\text{-}timer}_{k-2}$ to 2Δ and start counting down.
3) **(Non-blocking) Pre-commit.** When $\mathsf{pre\text{-}commit\text{-}timer}_k$ reaches 0, pre-commit B_k and broadcast $\langle \mathsf{commit}, B_k, v\rangle$.
4) **(Non-blocking) Commit.** On receiving $\langle \mathsf{commit}, B_k, v\rangle$ from $f+1$ distinct replicas with the same v, commit B_k and all its ancestors.

Fig. 3: The steady state protocol with mobile sluggish faults.

and may not have received an equivocating block. However, if the replica hears from a majority honest replicas about an undisturbed 2Δ period, then an equivocation could not have missed all of them and thus the replica can commit the block. We call the prior state a *pre-commit* and the latter state a commit.

Protocol. Interestingly, despite a weaker model than the standard synchrony assumption, based on the intuition presented above, the protocol is only marginally different from the one presented in Figures 1 and 2. For clarity we present the entire steady state protocol and gray out the repetition in Figure 3. The view-change protocol remains unchanged.

Now, since we require a $\mathcal{C}_v(\mathcal{C}_v(B_k))$ before waiting for the 2Δ period, we start a *pre-commit timer* for B_{k-2} on receiving B_k. When the pre-commit timer expires, the replica broadcasts a commit request $\langle \mathsf{commit}, B_k, v\rangle$ to all replicas. Upon receiving $f+1$ commit requests, the replica commits since at least one of the requests is from a prompt replica.

C. Safety and Liveness

The proof in the mobile sluggish model has the same structure as the standard synchrony model. Lemma 6 below is almost identical to Lemma 1 except that the claim is made for $f+1$ honest replicas instead of all honest replicas, which is as expected because the remaining d honest replicas can be sluggish. We can prove unique extensibility and safety identically as Lemma 2 and Theorem 3 respectively; by invoking Lemma 6 wherever Lemma 1 is invoked in those proofs.

As before, we say a block B_k is committed *directly* if it is committed due to $f+1$ pre-commits. We say a block B_k is committed *indirectly* if it is a result of directly committing a block extending B_k.

Lemma 6 (Lemma 1 extended to mobile sluggish)**.** *If an honest replica directly commits a block B_l in view v, then $f+1$ honest replicas (i) vote for B_l in that view, and (ii) receive $\mathcal{C}_v(B_l)$ before entering the next view.*

Proof. If an honest replica directly commits B_l in view v, then $d+1$ honest replicas pre-commit B_l in view v. Denote the set of these $d+1$ replicas by R. Let the earliest pre-commit among R be performed by replica r_1 at time t. Thus, replica r_1 must have received B_{l+2} at time $t-2\Delta$. This implies that $f+1$ replicas have voted for B_{l+1} before time $t-2\Delta$. Since $f+1$ replicas are honest and prompt at time $t-2\Delta$, at least one of these replicas, say replica r_2, intersects the quorum of replicas that voted for B_{l+1}. Denote the set of honest and prompt replicas at time $t-\Delta$ by R'. Since r_2 is honest and prompt at $t-2\Delta$, its vote for B_{l+1}, which contains $\mathcal{C}_v(B_l)$, reaches all replicas in R' by time $t-\Delta$.

We will now prove that the set R' is the required set that satisfies the two conditions in the lemma. The only scenario in which this is not true is if one of the replicas in R', say replica r' has voted for a conflicting block or has quit view v before time $t-\Delta$. However, because r' is honest and prompt at time $t-\Delta$, before time t, its broadcast of conflicting vote or blame certificate will reach all honest replicas that are prompt at time t. At least one replica in the pre-committing set R would be prompt and would have received this message by time t. It would have prevented the pre-commit of B_l at that replica, a contradiction. □

Remark. Note that the above lemma states that R' receives $\mathcal{C}_v(B_k)$ before quitting view v. Thus, the Δ wait during view-change is not needed for the protocol with sluggish faults.

Liveness. In the mobile sluggish model, liveness is guaranteed only during periods in which all honest replicas stay prompt. In that case, the same arguments in Theorem 4 hold. We do not repeat the proof.

D. Efficiency

In the mobile sluggish model, each pre-commit timer starts two proposals later, adding 4δ latency. The commit messages add another round of communication and δ latency. So the total latency becomes $2\Delta + 4\delta + 5\delta = 2\Delta + 9\delta$.

IV. Sync HotStuff with Optimistic Responsiveness

In this section, we incorporate the Thunderella [5] optimistic responsive mode into Sync HotStuff. In Section II, a certificate/quorum required only $f+1$ votes. Thus, a vote from a single honest replica *can* result in a certificate if all f

Byzantine replicas vote on the same block. Therefore, before committing a block, a replica needs to wait long enough to hear all honest replicas' votes and make sure none of them voted for a conflicting block. The commit latency thus inherently depends on the maximum network delay Δ. In contrast, partially synchronous protocols rule out the existence of a conflicting certificate with larger quorums. For instance, PBFT requires $> 2n/3$ votes (in two phases) and tolerates $f < n/3$ Byzantine replicas. A simple quorum intersection argument shows that two conflicting blocks cannot both receive $> 2n/3$ votes. Thus, partially synchronous protocols commit as soon as these quorums of votes are obtained, so the latency does not depend on Δ.

Pass and Shi [5] use the term *responsive* to capture the above latency distinction. A protocol is said to be *responsive* if the latency only depends on the actual network δ but not the maximum network delay Δ. A protocol is said to be *optimistically responsive* if it achieves responsiveness when some additional constraints are met.

Since Sync HotStuff aims to tolerate up to minority corruption, similar to Thunderella [5], to achieve responsiveness we will use a quorum size of $> 3n/4$. This will give Sync HotStuff *a responsive mode* that is optimistically responsive when messages from $> 3n/4$ replicas reach within time δ. This happens if the *actual* number of faults is less than $n/4$. Put differently, if more than $n/4$ replicas are faulty, they can prevent responsiveness but they cannot cause a safety violation. In that case, since there is no responsive progress, we fall back to the *synchronous mode* in Section III.

A. Protocol

Since the responsive mode requires a larger quorum, we introduce the notion of a *strong certificate*. A block B is said to have a strong certificate if it has $> 3n/4$ votes. In the responsive mode, an honest replica only votes for a proposal that includes a strong certificate for its predecessor block.

Figure 4 describes the augmented protocol with the responsive mode, augmented from the protocol in Section III. Below, we highlight the modifications and how to switch between the modes. Each view starts in the synchronous mode. If the leader recognizes that $> 3n/4$ replicas are voting, it signals a switch to the responsive mode by including a strong certificate. Once a replica votes for a block containing a strong certificate, it switches to the responsive mode. Once in the responsive mode, a replica will only vote for blocks containing strong certificates for the rest of the view. But crucially, the responsive commit rule (based on two strong certificates) does not immediately kick in. The replicas need to commit one more block using the synchronous commit rule (i.e., pre-commit after the 2Δ wait) after switching to the responsive mode. This essentially "commits the switch" using the regular synchronous commit rule, and thus ensures that most replicas have switched to the responsive mode.

Whenever the responsive mode fails due to an equivocating leader or lack of progress, replicas engage in a view-change (Figure 5) to move to the next view and start in its synchronous mode. Guarding the safety of responsive commits turns out to be non-trivial. We provide an intuition of the concern and our approach to solving it. This is discussed formally in Lemma 7 and 8.

In the protocol without responsiveness, a commit implied that a certificate was broadcast by some prompt replica 2Δ time earlier. This ensured that a certificate was obtained at a majority of honest replicas at or before they quit the view. Since we do not wait for 2Δ time in the responsive mode, this does not hold any more. First, some of the sluggish replicas (sluggish even before the responsive honest commit) can quit the old view and enter the new view without informing a majority of honest replicas. In the new view, they may vote for an unsafe block. By itself, this does not create a safety violation since $d + b < f$. However, in combination with the sluggish mobility that is introduced right after a prompt responsive honest commit, another set of d replicas may not learn a certificate before switching to the next view leading to a safety violation.

The solution is to increase the wait between quitting the old view and entering the new view to 2Δ (Step iii in Figure 5). The delay is introduced at a replica *after* learning that a majority of replicas have quit the view giving it sufficient time for the certificates to be sent across to a majority of prompt replicas before they enter and subsequently vote in the next view. These changes will be utilized in the proof of Lemma 7.

B. Safety with Optimistic Responsiveness

We now amend the proofs for safety to account for the responsive mode. Lemma 6 and 2 apply to directly committed blocks in the synchronous mode. We need to augment the lemmas to directly committed blocks in the responsive mode.

Lemma 7. *If an honest replica directly commits a block B_l in the responsive mode of view v, then $f + 1$ honest replicas receive $\mathcal{C}_v(B_l)$ (which can be either a normal or a strong certificate) before entering the next view.*

The proof is very similar to that of Lemma 6. The only difference is, as mentioned earlier, that the 2Δ in pre-commit now happens during view-change.

Proof. If an honest replica directly commits B_l in view v, then $d + 1$ honest replicas pre-commit B_l in view v. Denote the set of these $d + 1$ replicas by R. Let the earliest pre-commit among R be performed by replica r at time t. At least $d + 1$ honest replicas voted for B_{l+1} before time t. At least one of them is honest and prompt at time t. Its vote B_{l+1} reaches the set of honest replicas that are prompt time $t + \Delta$. Denote this set by R'. R' has size at least $f + 1$ and satisfies the lemma. The only scenario in which this is not true is if one of the replicas in R', say replica r' has entered view $v + 1$ before time $t + \Delta$. In that case, due to the 2Δ wait during view-change, r' has received $f + 1$ blame2 messages before time $t - \Delta$. Thus, $f + 1$ replicas have sent blame2 and quit view v before time $t - \Delta$. One of them is honest and prompt at time $t - \Delta$. Thus, at least one replica in the pre-committing

Let v be the current view number and replica L be the leader of the current view. The steady state protocol runs the following steps in iterations.

1) **Propose.** The leader L broadcasts $\langle \mathsf{propose}, B_k, v, \mathcal{C}_{v'}(B_{k-1})\rangle_L$ where $B_k = (b_k, h_{k-1})$, b_k is a batch of new client requests, $v' \leq v$, and B_{k-1} is the highest certified block known to L.
2) **Vote.** Upon receiving the first valid proposal for a height-k block B_k from L or through a vote by some other replica, if it is extending a highest certified block, forward the proposal to all other replicas and broadcast a vote in the form of $\langle \mathsf{vote}, B_k, v\rangle$. If B_k contains a strong certificate for its predecessor, switch to the responsive mode and only vote for blocks with strong certificates for the rest of this view.
3) **(Non-blocking) Pre-commit.** If at least one block has been committed in the responsive mode of this view, pre-commit B_{k-2} and broadcast $\langle \mathsf{commit}, B_{k-2}, v\rangle$ right away. Else, set $\mathsf{pre\text{-}commit\text{-}timer}_{k-2}$ to 2Δ and start counting down. When $\mathsf{pre\text{-}commit\text{-}timer}_{k-2}$ reaches 0, pre-commit B_{k-2} and broadcast $\langle \mathsf{commit}, B_{k-2}, v\rangle$.
4) **(Non-blocking) Commit.** On receiving $\langle \mathsf{commit}, B_k, v\rangle$ from $f+1$ distinct replicas with the same v, commit B_k and all its ancestors.

Fig. 4: The steady state protocol augmented with the responsive mode.

Let L and L' be the leader of view v and $v+1$, respectively.

i **Blame.** If less than p blocks are received from L in $(2p+1)\Delta$ time in view v, broadcast $\langle \mathsf{blame}, v\rangle$. If a leader L proposes equivocating blocks, broadcast $\langle \mathsf{blame}, v\rangle$ and the two equivocating blocks.

ii **Quit old view.** Upon gathering $f+1$ $\langle \mathsf{blame}, v\rangle$ messages, broadcast them along with $\langle \mathsf{blame2}, v\rangle$, and quits view v (abort all $\mathsf{pre\text{-}commit\text{-}timer}$(s) and stop voting in view v).

iii **Status.** Upon gathering $f+1$ $\langle \mathsf{blame2}, v\rangle$ messages, Wait for 2Δ time and enter view $v+1$. Upon entering view $v+1$, send a highest certified block to L' and transition back to steady state.

Fig. 5: The view-change protocol to support the responsive mode.

set R would be prompt and would have received this blame2 message by time t. It would have prevented the pre-commit of B_l at that replica, a contradiction. □

Lemma 8 (Unique Extensibility, restated from Lemma 2)**.** *If an honest replica directly commits a block B_l in view v, then (i) there does not exist a certificate for block $B'_l \neq B_l$ in the same view, (ii) all certified blocks with higher ranks (i.e., higher views and/or higher heights) extend B_l.*

Proof. If the direct commit happens in the synchronous mode, the proof of Lemma 2 (invoking Lemma 6) applies, so we only need to focus on the responsive mode here.

If an honest replica directly commits B_l in the responsive mode of view v, it implies two things: (1) that replica has performed a synchronous commit on a proper ancestor of B_l, say B_k, that contains a strong certificate to its predecessor, (2) a set R of $d+1$ honest replicas pre-commit B_l in the responsive mode of view v. The commit of B_k implies that at most d honest replicas are left behind in the synchronous mode, not enough to produce an equivocating certificate in the synchronous mode. Meanwhile, a strong certificate for a block that equivocates B_l requires at least $(3n/4+1)+(3n/4+1)-n = n/2+1 = f+1$ replicas to vote for equivocating blocks in the current view, which cannot happen.

This proves part (i), and also forms the base case of the inductive proof for part (ii). For the inductive step, observe that by Lemma 7, every honest replica receives $\mathcal{C}_v(B_l)$ (which can be either a normal or a strong certificate) before entering the next view. Due to the inductive hypothesis, any certified block with a higher rank from view v to v' extends B_l. Thus, before the start of view $v'+1$, all honest replicas lock on a block that extends B_l. Thus, any certified block from view $v'+1$ extends B_l. □

Safety. The safety proof remains identical to that of Theorem 3.

V. Evaluation

In this section, we first evaluate the throughput and latency of Sync HotStuff under different parameters and conditions (batch size, payload, and client command load). We then evaluate the impact of Δ on throughput and latency and show it is insignificant as expected. Lastly, we compare Sync HotStuff with HotStuff [12] and Dfinity [3].

A. Implementation Details and Methodology

We implement the Sync HotStuff protocol under the standard synchrony model. Our implementation is an adaptation of the open-source implementation of HotStuff [12]. We modify the HotStuff code to primarily replace the core protocol logic while reusing some of its utility modules, such as its event queue and network library.

In our implementation, replicas reach consensus on a sequence of blocks. Each block contains a batch of commands

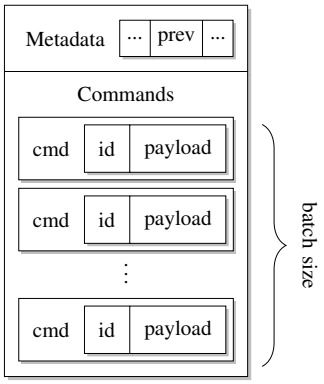


Fig. 6: Structure of a block in the implementation.

sent by clients. A command consists of a unique identifier and an associated payload. We refer to the maximum number of commands in a block as the batch size. A conceptual representation of a block is shown in Figure 6.

All throughput and latency results were measured from clients which are separate processes running on machines different from those for replicas. Each client generates a number of outstanding commands and broadcasts them to every replica. Replicas only use the unique identifier (e.g. hash) of a command to represent it in proposals and votes. To execute the commands for the replicated state machine, a replica either has the command content received from the client's initial broadcast, or fetches it from the leader if the client crashes before it finishes the broadcast. We use four machines, each running four client processes, to inject commands into the system. Each client process can maintain a configurable number of outstanding commands at any time. We ensure that the performance of replicas will not be limited by lack of client commands.

Experimental setup. All our experiments were conducted over Amazon EC2 where each replica was executed on a `c5.4xlarge` instance. Each instance had 16 vCPUs supported by Intel Xeon Platinum 8000 processors. All cores sustained a Turbo CPU clock speed up to 3.4GHz. The maximum TCP bandwidth measured by `iperf` is around 4.9 Gbps, i.e., 0.6 Gigabytes per second. We did not throttle the bandwidth in any run. The network latency between two machines is measured to be less than 1 ms. We used `secp256k1` for digital signatures in votes and a certificate consists of a compact array of `secp256k1` signatures.

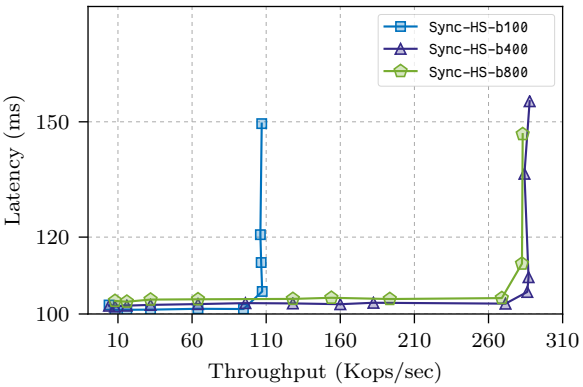
Baselines. We compare with two baselines: (i) HotStuff, a partially synchronous protocol, and (ii) Dfinity , a synchronous protocol. We use HotStuff as a baseline because (i) Sync HotStuff shares the same code base as HotStuff enabling a fair comparison, and (ii) HotStuff achieves comparable (sometimes even better) performance to the state-of-the-art partially synchronous BFT implementation [12]. We pick Dfinity as our other baseline because it is the state-of-the-art synchronous BFT protocol. We did not find an implementation of Dfinity's cousensus protocol in its Github repository, so we implemented our own version of Dfinity with our codebase (which should also help ensure a fair comparison). While implementing and evaluating Dfinity, we made several simplifications that are favorable to Dfinity. For instance, it was shown that a malicious leader can exploit a flaw in the original Dfinity design to force unbounded communication complexity [14]. We did not implement the suggested fix to this flaw (and of course we did not exploit this flaw). We assume all leaders in Dfinity are honest, which will improve Dfinity's theoretical latency from 9Δ to 7Δ. We simulate their Verifiable Random Functions (VRF), by essentially assuming VRF generation takes negligible time in Dfinity. Implementing these extra fixes and actual mechanisms will only further hurt Dfinity's performance.

B. Basic Performance

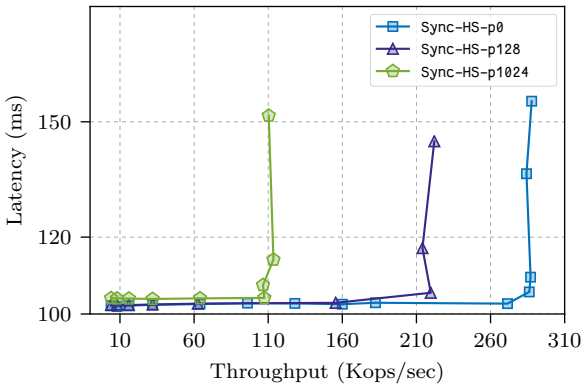
We first evaluate the basic performance of Sync HotStuff to tolerate $f = 1$ fault for a synchrony bound of $\Delta = 50$ ms. We measure the observed throughput (in number of commands committed/sec, or ops/sec) and end-to-end latency for clients (in ms). We conduct two experiments. The first one fixes the payload and varies batch size (Figure 7a) while the other fixes a batch size and varies the payload size (Figure 7b).

In Figure 7a, each command has a zero-byte payload to demonstrate the overhead induced solely by consensus and state machine replication. We consider three different batch sizes, 100, 400 and 800, represented by the three lines in the throughput-latency graph. In the graph, each point represents the measured throughput and latency for one run with a given "load" by the clients. More specifically, a client process maintains a fixed number of outstanding commands at any moment. When an outstanding command is committed, a new command is immediately issued to keep up with the specified number. We vary the size of the outstanding command pool to simulate different loads. The points at the lower left represent the state when the system is not saturated by client commands. As the load increases, the throughput initially increases without incurring a loss in latency. Finally, after the load saturates the bandwidth, the throughput remains unchanged (or slightly degrades) when clients inject more commands, while the latency goes up. The latency increases because the commands stay in the command pool longer before they are proposed in a block for consensus. For a batch size of 400, we observe that the throughput is saturated at around 280 Kops/sec. There is no further throughput gain when batch size increases from 400 to 800. So in all of our following experiments, we fix our batch size to 400.

We also test how payload size of a command affects performance. Figure 7b shows the performance with three client request/reply payload sizes (in bytes) 0/0, 128/128 and 1024/1024, denoted by "`p0`", "`p128`", and "`p1024`". In addition to the actual payload, each command also contains an 8 byte counter to differentiate the commands. For example, the actual command size for 0/0 is 8 bytes.

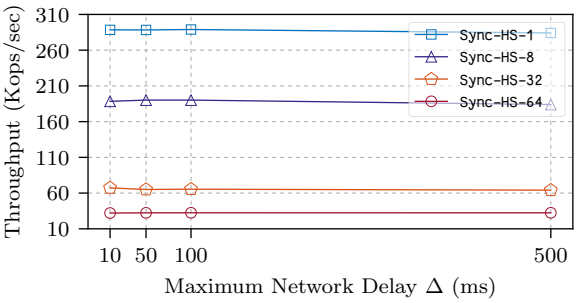


(a) Varying batch sizes.

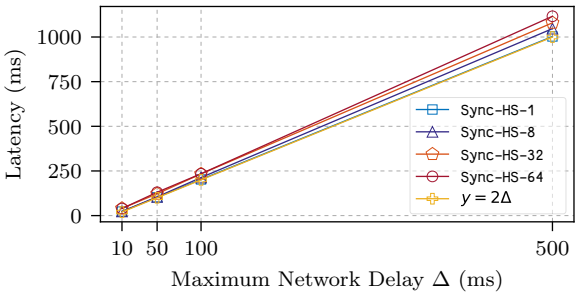


(b) Varying payload.

Fig. 7: Throughput vs. latency of Sync HotStuff at varying batch sizes and payloads at $\Delta = 50$ ms and $f = 1$.



(a) Δ vs. throughput.



(b) Δ vs. latency.

Fig. 8: Performance of Sync HotStuff at varying Δ and f at batch size = 400 and 0/0 payload.

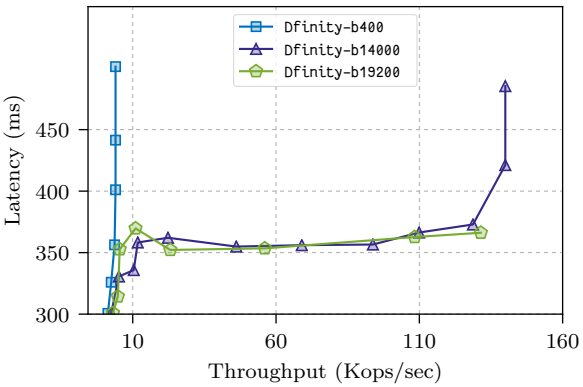


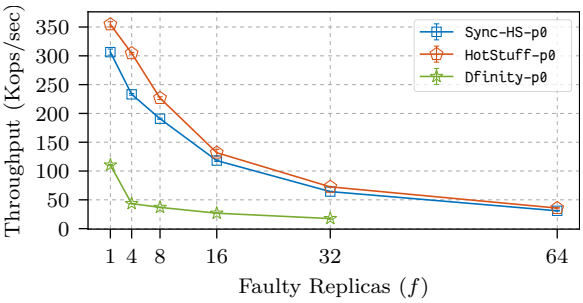
Fig. 9: Throughput vs. latency of Sync HotStuff at varying batch sizes at $\Delta = 50$ ms and $f = 1$ for Dfinity [14].

C. The Impact of Δ on Performance

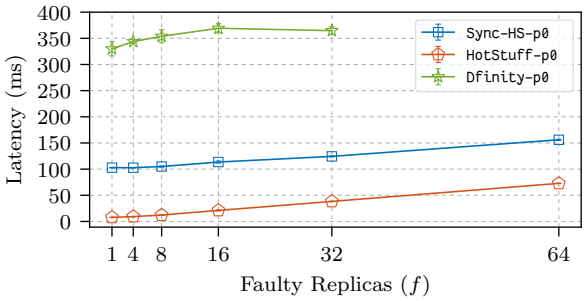
In the steady state of Sync HotStuff, replicas advance to the next step as soon as previous messages arrive, without waiting for any conservative Δ bound. Thus, although each block still incurs 2Δ latency to be committed, the system is able to move on after a single round-trip time, process new blocks in pipeline, and saturate available network bandwidth. Figures 8a and 8b study the effect of varying Δ on throughput and latency. Each line represents a choice of f, denoted by "1", "8", "32", "64". As expected, we observe that the saturated throughput remains unaffected by different choices of Δ, whereas the latency deviates little from the theoretical 2Δ line. We do note that the latency remains unaffected *only* when the Δ bound is conservative, because that is when the time for certifying a block (the $O(\delta)$ terms in our theoretical analysis) is overshadowed by the 2Δ wait. When tolerating a larger number of faults or when deployed on slower network conditions (e.g., consortium blockchains), Δ should be set appropriately to ensure safety.

D. Scalability and Comparison with Prior Work

We perform an experiment to understand how Sync HotStuff scales as the number of replicas increases. We also compare this with HotStuff and Dfinity. In our baseline, clients issue zero-byte payload commands and saturate the system, without overloading the replicas. We then vary the choice of f. Each experiment is repeated five times with the same setup to average out fluctuations. A data point shows the average value, capped by error bars indicating the standard deviation. Since synchronous protocols tolerate one-half faults as against one-third in case of partial synchrony, for the same f, the actual number of replicas is $2f + 1$ for Sync HotStuff and Dfinity, whereas it is $3f + 1$ for HotStuff. We would also like to point out that such comparison is not entirely fair since HotStuff does not assume synchrony. Nevertheless, it is

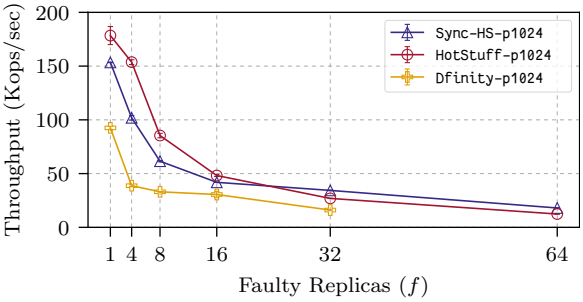


(a) f vs. throughput.

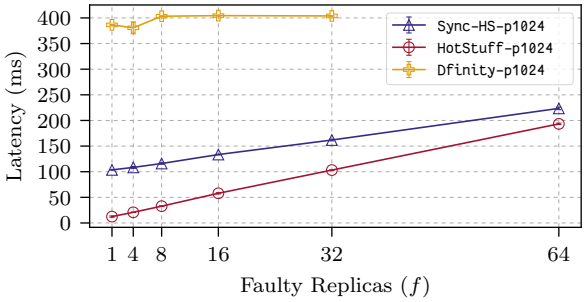


(b) f vs. latency.

Fig. 10: Performance as function of faults at Δ = 50 ms, optimal batch size, and 0/0 payload.



(a) f vs. throughput.



(b) f vs. latency.

Fig. 11: Performance as function of faults at Δ = 50 ms, optimal batch size, and 1024/1024 payload.

still helpful to understand the performance of Sync HotStuff by comparing it to a state-of-the-art partially synchronous protocol like HotStuff.

Comparison with HotStuff. Figures 10 and 11 show the throughput and latency for two different payload configurations, 0/0 and 1024/1024. We use a batch size of 400 for Sync HotStuff and HotStuff. Generally, the throughput of Sync HotStuff tends to be slightly worse than HotStuff. But at more faults, the throughput of Sync HotStuff gets closer to HotStuff and in the 1024/1024 case, eventually surpasses HotStuff. This is because in both cases the system is bottlenecked by a leader communicating with all other replicas and since Sync HotStuff requires fewer replicas to tolerate f faults, its performance scales better than HotStuff.

Comparison with Dfinity. For Dfinity, we first perform an experiment to determine good batch sizes that maximize its throughput. The results are shown in Figure 9. We observe that Dfinity requires a batch size of 14000 to reach its peak throughput of $\sim$130 Kops/sec. The reason why Dfinity requires a much larger batch size is because proposals are made much less frequently at every 2Δ time. In contrast, in Sync HotStuff, a new proposal is made every $2\delta \ll 2\Delta$ time, i.e., as soon as the previous proposal has been "processed" (i.e., certified). This allows Sync HotStuff to fully utilize available network bandwidth with much smaller batch sizes.

Figures 10b and 11b show the latency for two different payload configurations, 0/0 and 1024/1024. As can be seen in the figures, the latency of Dfinity varies between 330ms and 400ms. This is much higher than that for Sync HotStuff and is consistent with the expected theoretical average latency as described in Section V-A. We also observe that at $f = 64$, the large batch size we choose for Dfinity violates our $\Delta = 50$ms synchrony bound, leading to safety violations. Hence, our evaluation does not include that data point.

VI. Related Work

Several decades of research on the Byzantine agreement problem [15] brought a myriad of solutions. Dolev and Strong gave a deterministic protocol for its related problem, Byzantine broadcast, with the tolerance of $f < n-1$ [16]. Their protocol achieves $f+1$ round complexity and $O(n^2 f)$ communication complexity. The $f+1$ round complexity matches the lower bound for deterministic protocols [17], [16]. To further improve round complexity, randomized protocols have been introduced [18], [19], [20], [21], [22], [23]. We review the most recent and closely related works below.

Some key design goals of Sync HotStuff are inspired by recent related works. In particular, elimination of lock-step synchrony is first explored by Hanke et al. [3] and the mobile sluggish model is introduced by Guo et al. [6]. Compared to these works, Sync HotStuff uses techniques that are significantly simpler and more efficient to achieve the same goals.

Dfinity. The Dfinity Consensus protocol described in [3] is a replication protocol in the synchrony model that tolerates $f < n/2$ Byzantine faults. It makes a key observation that a synchronous replication protocol can start processing the next

client request without waiting for the previous one to commit. While a standard practice in partially synchronous protocols, this was not obvious for synchronous protocols.

However, it was recently discovered [14] that the presentation in [3] allowed unbounded communication complexity due to an exploitable requirement that honest replicas vote for all leader proposals, including conflicting ones. Another inefficient design in Hanke et al. is that each leader makes only one proposal before getting replaced by a new random leader. This hurts latency because (i) up to half of the proposal opportunities are wasted on Byzantine leaders, (ii) their leader election step is on the critical path and is blocking (in this sense, Hanke et al. did not fully remove lock-step execution.) This results in a large latency of $9\Delta + O(\delta)$.

In comparison, Sync HotStuff removes lock-step execution using a simple and natural protocol (replicas do not vote for conflicting proposals). Sync HotStuff embraces the stable leader approach, common in the partial synchrony SMR (e.g., PBFT [24], Paxos [25]), that uses a steady state leader to drive many decisions. These techniques allow Sync HotStuff to achieve $2\Delta + O(\delta)$ latency and expected quadratic communication complexity.

Guo et al. and PiLi. Guo et al. [6] introduced the mobile sluggish model (called weak synchronous model in that work). This model better reflects reality compared to the standard synchrony model and we adopt it. PiLi [4] presents a BFT SMR protocol in the mobile sluggish model. Its solution is theoretical and highly involved. It assumes lock-step execution in "epochs". Each epoch lasts for 5Δ time and the protocol commits five blocks after 13 consecutive epochs if certain conditions are met. Thus, PiLi requires a latency of at least 40Δ-65Δ (65Δ for the earliest and 40Δ for the latest of the five blocks). In contrast, we observe that simple techniques suffice to handle mobile sluggish faults at almost no extra overhead.

Thunderella. The notion of optimistic responsiveness (under one-quarter faults) was introduced in Thunderella [5]. Thunderella observes that it is safe to commit a decision in $O(\delta)$ time if $> 3n/4$ votes are received. In this paper, we adopt the key idea of Thunderella to achieve optimistic responsiveness. But we also make two changes. First, when more than $3n/4$ replicas are correct, Thunderella commits a decision after a single round of voting. However, the decision cannot be conveyed to external clients, hence it does not support SMR. Sync HotStuff uses two-phase commit in the responsive path, and hence provides safety for SMR. Second, Thunderella uses the synchronous mode to monitor the progress of the responsive mode and, if it does not make progress quickly, falls back to the synchronous mode. The fallback mechanism is presented in a black-box fashion but it is unclear how it works in a non-Nakamoto-style protocol. In Sync HotStuff, we take the conventional approach of having replicas monitor the progress of the responsive mode, and using the view-change protocol to perform the fallback.

XFT. A different type of protocol with optimistic responsiveness is XFT [26]. XFT guarantees responsiveness when a group of $f + 1$ honest replicas is determined. Thus, when the actual number of faults is t, it may take $\binom{n}{f+1}/\binom{n-t}{f+1}$ view-changes for an honest group of $f + 1$ to emerge; after that, the protocol is responsive. Such a solution is practical when t is a small constant but for $t = \Theta(n)$, it requires an exponential number of view-changes to find an honest group. In comparison, Sync HotStuff and Thunderella are responsive under $t < n/4$ faults after at most t view-changes.

VII. Conclusion

In this work, we introduce Sync HotStuff, a simple and practical synchronous BFT SMR protocol. Sync HotStuff does not require lock-step execution, tolerates mobile sluggish faults, and offers practical performance. As we mentioned, the mobile sluggish fault model captures short network glitches but is not ideal for replicas going offline for too long. It remains interesting future work to come up an even weaker and more realistic synchronous model as well as practical solutions in that model that still tolerate up to minority faults.

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