Que Sera Consensus:
Simple Asynchronous Agreement with Private Coins and Threshold Logical Clocks

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Abstract

It is commonly held that asynchronous consensus is much more complex, difficult, and costly than partially-synchronous algorithms, especially without using common coins. This paper challenges that conventional wisdom with que sera consensus (QSC), an approach to consensus that cleanly decomposes the agreement problem from that of network asynchrony. QSC uses only private coins and reaches consensus in $O(1)$ expected communication rounds. It relies on “lock-step” synchronous broadcast, but can run atop a threshold logical clock (TLC) algorithm to time and pace partially-reliable communication atop an underlying asynchronous network. This combination is arguably simpler than partially-synchronous consensus approaches like (Multi-)Paxos or Raft with leader election, and is more robust to slow leaders or targeted network denial-of-service attacks. The simplest formulations of QSC atop TLC incur expected $O(n^2)$ messages and $O(n^4)$ bits per agreement, or $O(n^3)$ bits with straightforward optimizations. An on-demand implementation, in which clients act as “natural leaders” to execute the protocol atop stateful servers that merely implement passive key-value stores, can achieve $O(n^2)$ expected communication bits per client-driven agreement.

1 Introduction

Most consensus protocols deployed in practice are derived from Paxos [58, 59], which relies on leader election and failure detection via timeouts. Despite decades of refinements and reformulations [11, 24, 47, 59, 75, 85], consensus protocols remain complex, bug-prone [1, 61] even with formal verification [24], and generally difficult to understand or implement correctly. Because they rely on network synchrony assumptions for liveness, their performance is vulnerable to slow leaders or targeted network denial-of-service attacks [4, 26].

Fully-asynchronous consensus algorithms [8, 14, 17, 21, 42, 70, 81] address these performance vulnerabilities in principle, but are even more complex, often slow and inefficient in other respects, and rarely implemented in practical systems. The most practical asynchronous consensus algorithms in particular rely on common coins [3, 9, 17, 18, 21, 27, 28, 35, 42, 69, 71, 81, 96], which in turn require even-more-complex distributed setup protocols [15, 52, 56, 103].

This paper makes no attempt to break any complexity-theoretic records, but instead challenges the conventional wisdom that fully-asynchronous consensus is inherently more complex, difficult, or inefficient than partially-synchronous leader-based approaches. To this end we introduce que sera consensus (QSC), a randomized con-
sensus algorithm that relies only on private coins and is expressible in 13 lines of pseudocode (Algorithm 1 on page 7). This algorithm relies on neither leader-election nor view-change nor common-coin setup protocols to be usable in practice. QSC does assume private, in-order delivery between pairs of nodes, but this requirement is trivially satisfied in practice by communicating over TLS-encrypted TCP connections, for example [86, 98].

QSC also relies on a new threshold synchronous broadcast (TSB) communication abstraction, in which coordinating nodes operate logically in lock-step, but only a subset of nodes’ broadcasts in each step may arrive. TSB provides each node an operation Broadcast(m) → (R, B), which attempts to broadcast message m, then waits exactly one (logical) time step. Broadcast then returns a receive set R and a broadcast set B, each consisting solely of messages sent by nodes in the same time step.

The level of reliability a particular TSB primitive guarantees is defined by three parameters: a receive threshold \( t_r \), a broadcast threshold \( b \), and a spread threshold \( s \). A TSB(\( t_r \), \( b \), \( s \)) primitive guarantees that on return from Broadcast on any node, the returned receive set R contains the messages sent by at least \( t_r \) nodes. Further, the returned broadcast set B contains messages broadcast by at least \( b \) nodes and reliably delivered to (i.e., appearing in the returned R sets of) at least \( s \) nodes in the same time step, provided the receiving nodes have not (yet) failed during the time step.

To implement this synchronous broadcast abstraction atop asynchronous networks, we introduce a class of protocols we call threshold logical clocks (TLC). Like Lamport clocks [57, 82], TLC assigns integer numbers to communication events independently of wall-clock time. Also like Lamport clocks but unlike vector clocks [36, 37, 39, 62, 67, 82] or matrix clocks [34, 82, 87, 88, 102], all communicating nodes share a common logical time. Unlike Lamport clocks, which merely label arbitrary communication timelines, TLC not only labels but also actively paces communication so that all nodes progress through logical time in “lock-step” – although different nodes may reach a logical time step at vastly different real (wall-clock) times. Nodes that fail (crash) may be conceptually viewed as reaching some logical time-steps only after an infinite real-time delay.

TLCR, a simple receive-threshold logical clock algorithm, implements TSB(\( t_r \), 0, 0) communication for a configurable threshold \( t_r \), in 11 lines of pseudocode (Algorithm 2 on page 10). TLCB, a broadcast-threshold logical clock algorithm, builds on TLCR to implement TSB(\( t_r \), \( b \), \( n \)) full-spread broadcast communication in five lines of pseudocode (Algorithm 3 on page 12). Full-spread broadcast ensures that at least \( b \) nodes’ messages in each round reach all nodes that have not failed by the end of the round, as required by QSC. In a configuration with \( n \geq 3f \) nodes where at most \( f \) nodes can fail, QSC atop TLCB (atop TLCR) ensures that each consensus round enjoys at least a 1/3 probability of successful commitment, yielding three expected consensus rounds per agreement.

This combination represents a complete asynchronous consensus algorithm, expressible in less than 30 lines of pseudocode total, and requiring no leader election or common coin setup or other dependencies apart from standard network protocols like TCP and TLS. To confirm that the pseudocode representation is not hiding too much complexity, Appendix B presents a fully-working model implementation of QSC, TLCB, and TLCR in only 37 lines of Erlang, not including test code.

QSC over TLCB is usable in \( n = 2f + 1 \) configurations only in the special (but common in practice) case of \( f = 1 \) and \( n = 3 \). Alleviating this restriction, TLCW (Algorithm 4 on page 13) directly implements TSB(\( b \), \( b \), \( b \)) communication for configurable \( b \) and \( b \), by proactively confirming the delivery of \( b \) messages to \( b \) nodes each, similar to signed echo broadcast [84] or witness cosigning [97] as used in other recent consensus protocols [3, 16, 55]. TLCF (Algorithm 5 on page 14), in turn, implements full-spread TSB(\( t_r \), \( b \), \( n \)) communication atop TLCW provided \( t_r + s > n \). QSC atop TLCF (atop TLCW) supports minimal \( n = 2f + 1 \) configurations for any \( f \geq 0 \), and ensures that each consensus round succeeds with at least 1/2 probability, for two expected consensus rounds per successful agreement.

QSC incurs only \( O(n) \) bits of communication per round if messages are constant-size. The TLC algorithms incur \( O(n^4) \) bits per round if implemented naively, but this is easily reduced to \( O(n^3) \) with simple optimizations.

Further efficiency improvements are feasible with QSCOD, an on-demand approach to implementing QSC. In QSCOD, clients wishing to commit transactions are responsible for driving communication and protocol progress, and the stateful consensus nodes merely im-
implement passive key-value stores. QSCOD clients effectively serve as “natural leaders” to drive communication and consensus efficiency using only $O(n^2)$ expected bits per client-driven agreement. In transactional applications, contention can require some clients to retry when other clients’ proposed transactions “win.” The communication costs of retries may be mitigated using classic techniques such as by exponential backoff as in CSMA/CD [49], or by committing batches of gossiped transactions together in blocks as in Bitcoin [72].

In summary, this paper’s main contributions are (a) threshold synchronous broadcast (TSB), a lock-step broadcast communication abstraction with parameterized delivery thresholds; (b) que sera consensus (QSC), a simple consensus protocol that builds atop the synchronous TSB abstraction but requires neither leader election, view changes, nor common coins; and (c) threshold logical clocks (TLC), a framework for timing and pacing group communication that builds lock-step TSB communication primitives atop asynchronous underlying networks.

This paper extends and formalizes ideas first proposed informally in an earlier preprint outlining the principles underlying threshold logical clocks [41].

2 Background

There are many different formulations of consensus and closely-related problems such as atomic broadcast [19]. QSC’s aim is to provide a practical asynchronous consensus protocol functionally equivalent to Paxos [58, 59] or Raft [75]. In particular, QSC provides the equivalent of Multi-Decree Paxos [58] or Multi-Paxos [23], where the goal is to agree on not just one value, but to commit a sequence of proposed values progressively to form a total order.

Because deterministic algorithms cannot solve asynchronous consensus [38], QSC relies on randomness for symmetry-breaking [6]. Like Ben-Or’s early exponential-time randomized protocol [8] but unlike the vast majority of more efficient successors, QSC relies only on private randomness: coins that each node flips independently of others, as provided by the random number generators standard in modern processors and operating systems. A key goal in particular is not to rely on common coins, where all nodes choose the same random values. While protocols based on secret sharing [91, 92, 94] can produce common coins or public randomness efficiently [17, 21, 96], robust asynchronous setup of common coins is essentially as difficult as asynchronous consensus itself [15, 41, 52, 56, 103].

2.1 System model and threat model assumptions

We assume as usual a group of $n$ nodes communicating over a network by sending and receiving messages. A node broadcasts a message to the group by sending $n$ identical messages, one to each member including itself.

We assume nodes follow the protocols faithfully as specified. Nodes can fail, but only by crashing cleanly and permanently, producing no more messages after the crash. While it appears readily feasible to extend QSC and TLC to account for Byzantine node behavior [41], we leave this goal for future work.

We make the standard asynchronous model assumption that the network eventually delivers every message, but only after an arbitrary finite delay of a (network) adversary’s choosing. For simplicity, QSC also assumes that messages are delivered in-order between pairs of nodes. Both assumptions are satisfied in practice if nodes communicate over TCP [98] or another reliable, ordered transport [40, 95, 100]. We assume ordered connections never fail unless one endpoint fails: e.g., timeouts are disabled and connections are protected against reset attacks [64].

QSC further assumes that nodes communicate over private channels (e.g., encrypted with TLS [86]), or that the network adversary is content-oblivious [6] or unable to look into the content of messages or process memory. Given the prevalence of deep-packet inspection technologies that intelligent network adversaries can readily employ, the use of encrypted channels seems safer than obliviousness assumptions in today’s Internet.

3 Threshold Synchronous Broadcast (TSB)

Before describing QSC, we first introduce a conceptually simple collective communication abstraction we call threshold synchronous broadcast (TSB). TSB presumes
that a group of \( n \) communicating nodes conceptually operates not in the asynchronous model above but in lock-step synchronous rounds, which we will call time steps or just steps. In each step, each node in the group that has not (yet) failed broadcasts a message to the others, then receives some subset of all messages sent in that round. Messages sent are tied to and received only in the same time-step: any messages a node does not receive in a given time-step are simply “lost” to that node forever and are never delivered late.

For now we treat threshold broadcast as a primitive API that a (slightly unrealistic) underlying network might conceivably provide. We will later develop algorithms to implement this abstraction atop asynchronous networks.

A TSB primitive does not in general offer perfect communication reliability. TSB instead guarantees reliability only so as to meet certain threshold parameters, hence the name. We say that a broadcast primitive provides TSB\((t_r, t_b, t_s)\) reliability if: (a) it guarantees that each node receives the messages broadcast by at least \( t_r \) nodes in the same time-step, and (b) it guarantees that the messages sent by at least \( t_b \) nodes are each reliably delivered or spread to at least \( t_s \) nodes each. A perfectly-reliable TSB primitive would be TSB\((n, n, n)\), guaranteeing that every message sent in each round reaches every node. A completely-unreliable TSB primitive would be TSB\((0, 0, 0)\), which makes no message delivery guarantees at all and hence might not be very useful, although it might sometimes deliver some messages.

For simplicity, we assume time is measured in integer units, as if each broadcast were a “real-time” operation taking exactly one unit of time. That is, each node broadcasts exactly one message at time-step 0 intended to be received at time-step 1, at step 1 each node broadcasts exactly one message to be received at step 2, and so on.

We represent the threshold broadcast primitive as a single API function, Broadcast\((m) \rightarrow (R, B)\). When a node \( i \) calls Broadcast\((m)\) at time-step \( s \), the network broadcasts message \( m \), waits exactly one time-step, then returns two message sets \( R \) and \( B \) to the caller. Returned set \( R \) is a set of messages that node \( i \) received during time-step \( s \). Returned set \( B \) indicates a set of messages that were each broadcast reliably to at least \( t_s \) nodes each.

When a message is reliably broadcast to a node \( j \), this means that node \( j \) will receive the message in the same time-step \( s \), unless \( j \) fails before time-step \( s \) completes. Thus, a failed node \( j \) may count toward towards a broadcast message’s spread threshold \( t_s \), provided we can be certain that node \( j \) would receive the message had it not failed.\(^1\)

A TSB\((t_r, t_b, t_s)\) primitive guarantees on return from Broadcast that \( R \) contains the messages broadcast by at least \( t_r \) nodes in step \( s \), and that \( B \) contains messages broadcast by at least \( t_b \) nodes, each of which is reliably delivered to at least \( t_s \) nodes during step \( s \). TSB makes no other delivery guarantees, however. For example, TSB makes no guarantee even that \( i \)’s own message \( m \) is within the sets \( R \) or \( B \) returned to \( i \). Further, two nodes \( i \) and \( j \) may see different received sets \( R_i \neq R_j \) and/or different broadcast sets \( B_i \neq B_j \) returned from their respective Broadcast calls in the same step. And two messages \( m_1 \in B_i \) and \( m_2 \in B_i \), both returned from the same node \( i \)’s Broadcast call, may have been broadcast to different node sets \( N_1 \) and \( N_2 \) respectively: TSB guarantees only that \(|N_1| \geq t_s\) and \(|N_2| \geq t_s\) and not that \( N_1 = N_2 \).

**Definition 3.1.** A network offers a TSB\((t_r, t_b, t_s)\) primitive provided:

- **Lock-step synchrony:** A call to Broadcast\((m)\) at any integer time-step \( s \) completes and returns at time-step \( s + 1 \), unless the node fails before reaching time-step \( s + 1 \).
- **Receive threshold:** If a node \( i \)’s call to Broadcast\((m)\) at step \( s \) returns \((R, B)\), then there is a node set \( N_R \subseteq \{1, \ldots, n\} \) such that \(|N_R| \geq t_r\), and \( R \) contains exactly the set of messages \( m_j \) broadcast by nodes \( j \in N_R \) during step \( s \).
- **Broadcast threshold:** If a node \( i \)’s call to Broadcast\((m)\) at step \( s \) returns \((R, B)\), then there is a node set \( N_B \subseteq \{1, \ldots, n\} \) such that \(|N_B| \geq t_b\), and \( B \) contains exactly the set of messages \( m_j \) broadcast by nodes \( j \in N_B \) during step \( s \).

\(^1\) An alternative, perhaps mathematically cleaner conception of node failures is to presume that all nodes always “eventually” reach all time steps. But when a node \( j \) “fails” before step \( s \), this simply means that \( j \) reaches \( s \) after an infinite real-time delay, i.e., \( j \) reaches \( s \) at wall-clock time \( \infty \). Adopting this viewpoint, TSB’s promise that a message \( m \in B \) will “eventually” reach at least \( t_s \) nodes becomes unconditional, independent of node failure. This is because any failed node \( j \) in that set conceptually does reach step \( s \) and receive \( m \), only at real-time \( \infty \).
• Spread threshold: If a node \( i \) called Broadcast\((m)\) at step \( s \), which returned \((R, B)\) such that a message \( m' \in B \), then there are at least \( ts \) nodes whose message sets \( R \) to be returned from Broadcast in step \( s \) will include message \( m' \).

The special case of TSB\((t_r, t_b, n)\), where \( t_s = n \), represents a particularly-useful full-spread broadcast primitive. Such a primitive guarantees that in each time-step \( s \), each reliably-broadcast message returned in any node’s \( B \) set is delivered to all \( n \) nodes during time-step \( s \), apart from any nodes that fail before step \( s \) completes. A full-spread TSB has the useful property that the \( B \) set returned to any node is a subset of the \( R \) set returned to any (other) node.

**Lemma 3.1.** In a network of \( n \) nodes offering a full-spread TSB\((t_r, t_b, n)\) primitive, if a node \( i \)’s call to Broadcast\((m_1)\) at step \( s \) returns \((R_1, B_1)\), and a node \( j \)’s call to Broadcast\((m_2)\) at the same step \( s \) returns \((R_2, B_2)\), then \( B_1 \subseteq R_2 \).

**Proof.** This property directly follows from the broadcast spread property.

## 4 Que Sera Consensus

In this section we describe the *que sera consensus* (QSC) protocol, then analyze its correctness and complexity.

### 4.1 Building consensus atop TSB

First we will define more precisely what properties we desire from a consensus protocol built atop a TSB communication abstraction. We will focus on implementing a multi-consensus protocol, functionally analogous to Multi-Paxos [23, 58], where nodes agree on a sequence of values (a log) instead of just one value as in (single-decree) Paxos [58]. For simplicity, the rest of this paper refers to multi-consensus simply as consensus.

We represent a consensus protocol \( \mathcal{P} \) as a process that runs concurrently and indefinitely (or until it fails) on each of a set of \( n \) nodes communicating via threshold reliable broadcast. We consider \( \mathcal{P} \) to be an algorithm parameterized by four functions: ChooseMessage, Deliver, RandomValue, and Broadcast. The first two function parameters represent an “upcall-style” interface to the application or higher-level protocol. \( \mathcal{P} \) invokes ChooseMessage to ask the application to choose the next message that the application wishes to commit. \( \mathcal{P} \) invokes Deliver to deliver committed messages up to the application. \( \mathcal{P} \)’s remaining two function parameters represent its lower-level interface to the network and operating system. \( \mathcal{P} \) calls RandomValue to choose a numeric value using node-private randomness, and \( \mathcal{P} \) calls Broadcast to broadcast a message using the TSB primitive and obtain the broadcast’s results one time-step later.

For simplicity of presentation and reasoning, our formulation of consensus protocols will deliver not just individual messages but entire *histories*, ordered lists cumulatively representing all messages committed and delivered so far. An easy and efficient standard practice is to represent a history as the typically constant-size head of a tamper-evident log [30, 90] or blockchain [72], each log entry containing a hash-link to its predecessor. Thus, the fact that histories conceptually grow without bound is not a significant practical concern.

Intuitively, the key properties we want from \( \mathcal{P} \) are liveness, validity, and consistency. Liveness means that \( \mathcal{P} \) regularly keeps advancing time and delivering progressively-longer histories via Deliver, forever or until the node fails. Validity means that any message delivered by any node is one that the application recently returned via ChooseMessage on some (potentially different) node. Finally, consistency means that a history delivered in an earlier time-step is a prefix of any history delivered in a later time-step, both on the same node and across distinct nodes.

**Definition 4.1.** A multi-consensus protocol \( \mathcal{P} \) is a potentially-randomized algorithm that takes function parameters \((\text{ChooseMessage, Deliver, RandomValue, Broadcast})\) and behaves as follows:

- **Liveness:** If \( h \) is the longest history \( \mathcal{P} \) has delivered by time-step \( s \) on some non-failing node \( i \), or \( h = \\{\} \) if \( \mathcal{P} \) has not yet invoked Deliver by step \( s \), then there is some future time-step \( s' > s \) at which \( \mathcal{P} \) invokes Deliver\((h')\) with some history \( h' \) strictly longer than \( h \) (i.e., \( |h'| > |h| \)).

- **Validity:** For some constant \( \delta \geq 0 \), if \( \mathcal{P} \) invokes Deliver\((h' \parallel \{p\})\) at time-step \( s' \) on node \( j \), then \( p \)
is a proposal \(\langle \text{proposal } i, m, r \rangle\) that node \(i\) returned from an invocation of \text{ChooseMessage} at some time-step \(s \leq s',\) where \(s' - s \leq \delta\).

- Consistency: if \(\mathcal{P}\) invokes \text{Deliver}(h) at time-step \(s\) on node \(i\), and \(\mathcal{P}\) later invokes \text{Deliver}(h') at time-step \(s' \geq s\) on node \(j\) (either the same or a different node), then \(h\) is a prefix of \(h'\).

\(\mathcal{P}\)'s behavior above is contingent on its function parameters satisfying operational specifications described below.

The application upcall function parameters \text{ChooseMessage} and \text{Deliver} may behave in arbitrary application-specific fashions, provided they do not interfere with the operation of \text{QSC} or the lower layers it depends on (e.g., by corrupting memory, or de-synchronizing the nodes via unexpected calls to \text{Broadcast}). \text{ChooseMessage} always returns some message, which may be an empty message if the application has nothing useful to broadcast in a given time-step. The \text{RandomValue} function must return a numeric value (integer or real) from a nontrivial random distribution, containing at least two values each occurring with nonzero probability, and from the same random distribution on every node. The \text{Broadcast} function must correctly implement a threshold reliable broadcast primitive as described above in Section 3.

### 4.2 Que Sera Consensus (QSC) algorithm

Algorithm 1 concisely summarizes que sera consensus (QSC), a simple multi-consensus algorithm satisfying the above specification. The QSC algorithm is a process that runs on each of the \(n\) nodes forever or until the node fails. Each iteration of the main loop implements a single consensus round, which may or may not deliver a new history \(h\).

The QSC algorithm depends on \text{Broadcast} providing a full-spread \text{TSB}\((t_r, t_h, n)\) abstraction (Section 3). Each consensus round invokes this \text{Broadcast} primitive twice, thus taking exactly two broadcast time-steps per consensus round.

Each node maintains its own view of history, denoted by \(h\), which increases in size by one entry per round. Each node does not build strictly on its own prior history in each round, however, but can discard its own prior history in favor of adopting one built by another node. In this way QSC’s behavior is analogous to Bitcoin [72], in which the “longest chain” rule may cause a miner to abandon its own prior chain in favor of a longer chain on a competing fork. QSC replaces Bitcoin’s “longest chain” rule with a “highest priority” rule, however.

At the start of a round, each node \(i\) invokes \text{ChooseMessage} to choose an arbitrary message \(m\) to propose in this round, possibly empty if \(i\) has nothing to commit. This message typically represents a transaction or block of transactions that the application running on \(i\) wishes to commit, on behalf of itself or clients it is serving. Node \(i\) also uses \text{RandomValue} to choose a random numeric priority \(r\) using node-private (not shared) randomness. Based on this information, node \(i\) then appends a new proposal \(\langle \text{proposal } i, m, r \rangle\) to the prior round’s history and broadcasts this new proposed history \(h'\) using \text{Broadcast}, which returns sets \(R'\) and \(B'\).

From the set \(B'\) of messages that \text{TSB} promises were reliably broadcast to all non-failed nodes in this first \text{Broadcast} call, node \(i\) picks any history \(h''\) (not necessarily unique) having priority at least as high as any other history in \(B'\), and broadcasts \(h''\). This second \text{Broadcast} call returns history sets \(B''\) and \(R''\) in turn. Finally, \(i\) picks from the resulting set \(R''\) any best history (Definition 4.2), i.e., any history (again not necessarily unique) with priority at least as high as any other in \(R''\), as the resulting history for this round and the initial history for the next round from node \(i\)’s perspective. We define the priority of a nonempty history as the priority of the last proposal it contains. Thus, a history \(h = [\ldots, \langle \text{proposal } i, m, r \rangle]\) has priority \(r\).

**Definition 4.2.** A history \(h\) is best in a set \(H\) if \(h \in H\) and no history \(h'\in H\) has priority strictly greater than \(h\).

The resulting history each node arrives at in a round may be either tentative or final. Each node decides separately whether to consider its history tentative or final, and nodes may make different decisions on finality in the same round. Each node \(i\) then delivers the resulting history \(h\) to the application built atop QSC, via a call to \text{Deliver}, only if node \(i\) determined \(h\) to be final. If \(i\) decides that \(h\) is tentative, it simply proceeds to the next round, leaving
 Algorithm 1: Que Sera Consensus (QSC)

**Input:** configuration parameters $n, t_r, t_b, t_s$, where $t_r > 0, t_b > 0$, and $t_s = n$

Run the following concurrently on each communicating process $i \in \{1, \ldots, n\}$:

$h \leftarrow []$  // consensus history is initially empty

forever

$m \leftarrow $ ChooseMessage ()  
$r \leftarrow $ RandomValue ()

$h' \leftarrow h || [(\text{proposal } i, m, r)]$

$(R', B') \leftarrow $ Broadcast($h'$)

$h'' \leftarrow$ any best history in $B'$

$(R'', B'') \leftarrow $ Broadcast($h''$)

$h \leftarrow$ any best history in $R''$

if $h \in B''$ and $h$ is uniquely best in $R'$  then

Deliver ($h$)

end

end

its view of history effectively undecided until some future round eventually delivers a final history.

A node $i$ decides that its resulting history $h$ in a round is final if (a) $h$ is in the set $B''$ returned from the second broadcast, and (b) $h$ is the uniquely best history in the set $R'$ returned from the first broadcast.

**Definition 4.3.** A history $h$ is uniquely best in a set $H$ if $h \in H$ and there is no other history $h' \neq h$ such that $h'$ is also in $H$ and has priority greater than or equal to that of $h$.

This pair of finality conditions is sufficient to ensure that all nodes will have chosen exactly the same resulting history $h$ at the end of this round, as explained below — even if other nodes may not necessarily realize that they have agreed on the same history. Since all future consensus rounds must invariably build on this common history $h$ regardless of which nodes' proposals “win” those future rounds, node $i$ can safely consider $h$ final and deliver it to the application, knowing that all other nodes will similarly build on $h$ regardless of their individual finality decisions.

### 4.3 Correctness of QSC

While the QSC algorithm itself is simple, analyzing the correctness of any consensus protocol involves some subtleties, which we examine first intuitively then formally. The main challenges are first, ensuring that the histories it delivers are consistent, and second, that it determines rounds to be final and delivers longer histories “reasonably often.” This section only states key lemmas and the main theorem; the proofs may be found in Appendix A.1.

#### 4.3.1 Safety

Consistency is QSC’s main safety property. We wish to ensure that if at some step $s$ a node $i$ delivers history $h$, and at some later step $s' \geq s$ any node $j$ delivers history $h'$, then $h$ is a prefix of $h'$. That is, every node consistently builds on any history prefix delivered earlier by any other node. To accomplish this, in QSC each node $i$ delivers a history $h$ only if $i$ can determine that all other non-failed nodes must also become aware that $h$ exists and can be chosen in the round, and that no other node could choose any other history in the round.

Each node first chooses some best (highest-priority) eligible history $h''$ from the set $B'$ returned by the first Broadcast call. Set $B'$ includes only confirmed histories: those that $i$ can be certain all non-failed nodes will become aware of during the round. By the full-spread requirement ($t_s = n$) on the TSB primitive, history $h''$ must be included in the $R'$ sets returned on all non-failed
nodes, ensuring this awareness requirement even if other nodes choose different histories.

After the second Broadcast\((h'')\) call, all histories returned in \(R''\) and \(B''\) are confirmed histories. Further, any history \(h \in B''\) is not just confirmed but reconfirmed, meaning that all non-failed nodes will learn during the round not just that \(h\) exists but also the fact that \(h\) was confirmed. Each node \(i\) chooses, as its tentative history to build on in the next round, some (not necessarily unique) highest-priority confirmed proposal among those in \(R''\) that \(i\) learns about.

Finally, \(i\) considers \(h\) committed and actually delivers it to the application only if \(h\) is both reconfirmed (in \(B''\)) and uniquely best within the broader set \(R'\) that includes all confirmed proposals other nodes could choose in this round. These two finality conditions ensure that (a) all nodes know that \(h\) is confirmed and thus can choose it as their tentative history, and (b) all nodes must choose \(h\) because it is uniquely best among all the choices they have. This does not guarantee that other nodes will know that \(h\) is committed, however: another node \(j \neq i\) might not observe the finality conditions, but will nevertheless “obliviously” choose and build on \(h\), learning only in some later round that \(h\) is final.

A key first step is showing that any node’s (tentative or final) history at any round builds on some node’s (tentative or final) history at any prior round.

**Lemma 4.1.** History preservation: If a consensus round starting at time-step \(s\) has initial history \(h_s\), on node \(i\), then at any earlier consensus round starting at step \(s' < s\), there exists some node \(j\) whose initial history \(h_{s'j}\) in that round is a strict prefix of \(h_s\).

Consistency, QSC’s main safety property, relies on the fact that each node \(i\) delivers a resulting history in a consensus round only when \(i\) is sure that all nodes will choose the same resulting history in that round. The above lemma in turn guarantees that the histories all nodes build on and potentially deliver in the future must build on this common history.

**Lemma 4.2.** Agreement on delivery: If QSC delivers history \(h_{(s+2)}\), on node \(i\) at the end of a consensus round starting at time-step \(s\), then the resulting history \(h_{(s+2)}\), of every node \(j\) in the same round is identical to \(h_{(s+2)}\).

### 4.3.2 Liveness

The other main prerequisite to QSC’s correctness is ensuring its liveness, i.e., that it makes progress. Unlike safety, liveness is probabilistic: QSC guarantees only that each node has a “reasonable” nonzero chance of delivering a committed history in each round. This ensures in turn that for each node \(i\), after any time-step \(s\), with probability \(1\), there exists some future time-step \(s' \geq s\) at which \(i\) delivers some (next) final history.

QSC’s liveness depends on the network scheduling message delivery independently of the contents of proposals. More precisely, QSC assumes that the network underlying TSB primitive chooses the sets \(N_R\) and \(N_B\), determining which messages each node receives and learns were reliably broadcast (Definition 3.1), independently of the random priority values contained in the proposals. As mentioned before, in practice, we can satisfy this assumption either by assuming a content-oblivious network scheduler [6], or by using private channels (e.g., encrypted by TLS [86]).

**Lemma 4.3.** If the network delivery schedule is independent of proposal priorities and \(p_t\) is the probability that two nodes tie for highest priority, then each node delivers a history in each round independently with probability at least \(t_b/n - p_t\).

### 4.3.3 Overall correctness of QSC

The above lemmas in combination ensure that QSC correctly implements consensus.

**Theorem 4.1.** QSC implements multi-consensus on \(n\) nodes (Definition 4.1) atop a full-spread TSB primitive TSB\((t_r, t_b, n)\) where \(t_r > 0\) and \(t_b > 0\).

Notice that QSC’s correctness theorem makes no direct assumptions about the number of failing nodes, in particular not mentioning the standard majority requirement \(n > 2f\). This is because the number of failing nodes affects only liveness, and QSC depends for liveness on the underlying TSB primitive’s unconditional promise to return from each Broadcast call in exactly one time-step. It will prove impossible to implement a TSB\((t_r, t_b, t_s)\) primitive where \(t_r > n - f\) or \(t_b > n - f\), because the primitive would have to collect messages from failed
nodes in order to return $R$ and $B$ sets of the promised size after each step. But that is TSB’s problem, not QSC’s.

### 4.4 Asymptotic complexity of QSC

Implementing QSC naively, the histories broadcast in each round would grow linearly with time. We can make QSC efficient, however, by adopting the standard practice of representing histories as tamper-evident logs or blockchains [30, 72, 90]. Each broadcast needs to contain only the latest proposal or head of the history, which refers to its predecessor (and transitively to the entire history) via a cryptographic hash. Since QSC does not need anything but the head in each round, this is sufficient. With this optimization, the two messages each node broadcasts in each round are $O(1)$ size. The total message and communication complexity of QSC is therefore $O(n^2)$ per round across the $n$ nodes, assuming each broadcast requires $n$ unicast transmissions (efficient broadcast would eliminate a factor of $n$). Provided $t_b/n$ is constant, it takes a constant expected number of rounds (namely $n/t_b$) to commit and deliver a new consensus result, so each consensus progress event likewise incurs $O(n^2)$ expected communication complexity. This analysis neglects the cost of implementing the underlying TSB abstraction that QSC builds on, of course, an issue we address later.

### 5 Threshold Logical Clocks

Since the TSB abstraction seems somewhat tailor-made for implementing consensus, it would not be particularly useful if it were almost as difficult to implement TSB as to implement consensus directly. Fortunately, there are multiple clean and simple ways to implement the TSB primitive atop realistic, fully-asynchronous networks.

For this purpose we develop several variants of a lower-level abstraction we call threshold logical clocks (TLC). The main purpose of TLC is to provide the illusion of lock-step synchrony that the TSB abstraction presents and that QSC relies on, despite the underlying network being asynchronous. Secondary, a TLC also conveniently provides the communication patterns needed to implement the threshold reliability that the TSB abstraction promises.

The rest of this section is organized as follows: In Section 5.1 we introduce TLCR, a simple receive-threshold logical clock algorithm realizing TSB$(t_r, 0, 0)$. Afterwards, in Section 5.2, we discuss TLCB a broadcast-threshold logical clock algorithm building on top of TLCR to provide full-spread broadcast communication TSB$(t_r, t_b, n)$. In Section 5.3 we then present TLCW, a witnessed-threshold logical clock algorithm implementing TSB$(t_r, t_b, t_s)$ communication, amending some of the restrictions of TCB. Finally, in Section 5.4, we describe TLF, which builds full-spread witness broadcast communication TSB$(t_r, t_b, n)$ on top of TLCR and TLCW. Proofs for theorems in this section are in Appendix A.2.

#### 5.1 TLCR: receive-threshold synchrony on asynchronous networks

Algorithm 2 implements a TSB$(t_r, 0, 0)$ abstraction atop an asynchronous network, ensuring that each node receives messages from at least $t_r$ nodes during each logical time-step. Although TLCR tolerates messages being scheduled and delayed arbitrarily, it makes the standard assumption that a message broadcast by any node is eventually delivered to every other non-failing node. For simplicity, TLCR also assumes messages are delivered in order between any pair of nodes, e.g., via any sequenced point-to-point transport such as TCP.

In TLCR, each node broadcasts a message at the beginning of each step $s$, then waits to receive at least $t_r$ messages from step $s$. TLCR internally logs the receive-set it returns from each step in $R$, whose length tracks the current time-step.

Each node’s broadcast in each step also includes the receive-set with which it completed the previous step. If a node receives any message from step $s + 1$ before collecting a threshold of messages from $s$, it immediately completes step $s$ using the previous receive-set it just obtained. Because of the above assumption messages are pairwise-ordered (e.g., by TCP), a node never receives a message for step $s + 2$ or later before receiving a message for step $s + 1$ from the same node, and thus never needs to “catch up” more than one step at a time.

As an alternative to including the previous step’s receive set in each broadcast, TLCR could simply defer the processing of messages for future steps until the receive
Theorem 5.1. Delayed messages arriving early from future time steps cannot virally help delayed nodes make progress. This approach eliminates the pairwise ordered-delivery assumption, at the potential cost of slightly slower progress in practice because messages arriving early from future time steps cannot virally help delayed nodes make progress.

Theorem 5.1. TLCR (Algorithm 2) implements a TSB($t_r, 0, 0$) communication primitive with receive threshold $0 \leq t_r \leq n$, provided at most $f \leq n - t_r$ nodes fail.

5.1.1 Asymptotic complexity of TLCR

Since each node broadcasts exactly one message per time-step, TLCR incurs a total message complexity of $O(n^2)$ per round across the $n$ nodes, assuming each broadcast requires $n$ unicasts.

If the messages passed to TLCR are constant size, then TLCR incurs a communication complexity of $O(n^3)$ per round because of TLCR’s inclusion of the previous round’s receive-set in each broadcast. Implementing TLCR naively, if the messages passed to TLCR are $O(n)$ size, then total communication complexity is therefore $O(n^4)$ per round, and so on.

A simple way to reduce this communication cost, however, is simply to defer the processing of messages for future time steps that arrive early, as discussed above. This way, broadcasts need not include the prior round’s receive-set, so communication complexity is only $O(n^2)$ per round when application messages are constant size.

Another approach is to replace the application messages themselves with constant-size references (e.g., cryptographic hashes) to out-of-line blocks, and to use a classic IHAVE/SENDME protocol as in USENET [46] on the point-to-point links between nodes to transmit only messages that the receiver has not yet obtained from another source. In brief, on each point-to-point message transmission the sender first transmits the summary message containing only references; the sender then waits for the receiver to indicate for which references the receiver does not yet have the corresponding content; and finally the sender transmits only the content of the requested references. With this standard practice in gossip protocols, each node typically receives each content block only once (unless the node simultaneously downloads the same block from multiple sources to minimize latency at the cost of bandwidth).

| Algorithm 2: TLCR($m$), using a threshold logical clock to implement receive-threshold synchronous broadcast |
|---|
| Configuration : node number $i$, number of nodes $n$, receive threshold $t_r \leq n$ |
| Configuration : functions Receive, Broadcast representing underlying asynchronous network API |
| Persistent state : receive message-set $\bar{R}$, initialized to the singleton list $\{\}$ |
| Function input : message $m$ to broadcast in this time-step |
| Function output: sets $(\bar{R}, B)$ of messages received in this time-step, and reliably broadcast (always empty) |
| $\bar{R} \leftarrow \bar{R} \cup \{\}$ | // start a new logical time-step with an empty receive message-set |
| Broadcast$(\langle i, m, |\bar{R}|, \bar{R}_{t_{\bar{R}}-1} \rangle)$ | // broadcast our message, current time-step, and last message-set |
| while $|\bar{R}| < t_r$ do |
| $\langle j, m', s', B' \rangle \leftarrow$ Receive() | // loop until we reach receive threshold $t_r$ to advance logical time |
| if $s' = |\bar{R}|$ then | // await and receive next message $m'$ from any node $j$
| $\bar{R}_{t_{\bar{R}}} \leftarrow \bar{R}_{t_{\bar{R}}} \cup \{j, m'\}$
| else if $s' > |\bar{R}|$ then | // message $m'$ was sent in our current time-step |
| $\bar{R}_{t_{\bar{R}}} \leftarrow \bar{R}_{t_{\bar{R}}} \cup R'$ | // collect messages received in this time-step |
| end |
| end |
| return $(\{m' | \langle j, m' \rangle \in \bar{R}_{t_{\bar{R}}}, \}, \})$ | // virally adopt message-set $R'$ that $j$ used to advance history |
|  |
| return the received message set and an empty broadcast set |
Implementing TLCR in this way, each round incurs a communication complexity of \(O(n^3)\) per round even if the messages passed to TLCR are \(O(n)\) size. This is because each node proposes only one new message \(m\) per round and each node receives its content only once, even if the prior round receive-set in each round’s proposal refers to \(O(n)\) messages from the prior round via constant-size references (e.g., hashes).

5.2 TLCB: broadcast-threshold synchrony atop TLCR

Although TLCR provides no broadcast threshold guarantees, in suitable network configurations, TLCB (Algorithm 3) does so by simply using two successive TLCR rounds per (TLCB) time-step. In brief, TLCB uses its second TLCR invocation to broadcast and gather information about which messages sent in the first TLCR invocation were received by enough \((t_s)\) nodes. Simple “pigeon-hole principle” counting arguments ensure that enough \((t_b)\) such first-round messages are so identified, provided the configuration parameters \(n, t_r, t_s,\) and \(t_b\) satisfy certain constraints. These constraints are specified in the following theorem, whose detailed underlying reasoning may be found in Appendix A.2.

**Theorem 5.2.** If \(0 < t_r \leq n - f, 0 < t_s \leq t_r, 0 < t_b \leq n - f_b\) where \(f_b = t_r(n - t_r)/(t_r - t_s + 1)\), and at most \(f\) nodes fail, then TLCB (Algorithm 3) implements a TSB\((t_r, t_b, t_s)\) partial-spread broadcast abstraction.

Suppose we desire a configuration tolerating up to \(f\) node failures, and we set \(n = 3f, t_r = 2f, t_b = f,\) and \(t_s = f + 1\). Then \(f_b = 2f(3f - 2f)/(2f - (f + 1) + 1) = 2f\), so \(t_b \leq n - f_b\) as required. This TLCB configuration therefore reliably delivers at least \(t_b = n/3\) nodes’ messages to at least \(t_s = n/3 + 1\) nodes each in every step.

5.2.1 Full-spread reliable broadcast using TLCB

If we configure TLCB above to satisfy the additional constraint that \(t_r + t_s > n\), then it actually implements full-spread reliable broadcast or TSB\((t_r, t_b, n)\). This constraint reduces to the classic majority rule, \(n > 2f\), in the case \(t_r = t_s = f\) where at most \(f\) nodes fail.

Under this constraint, each of the (at least \(t_b\)) messages in the set \(B\) returned from TLCB on any node is guaranteed to appear in the set \(R\) returned from the same TLCB round on every node that has not yet failed at that point. Intuitively, this is because the first TLCR call propagates each message in \(B\) to at least \(t_s\) nodes, every node collects \(R'\) sets from at least \(t_r\) nodes during the second TLCR call, and since \(t_r + t_s > n\) these spread and receive sets must overlap.

**Theorem 5.3.** If \(0 < t_r \leq n - f, 0 < t_s \leq t_r, t_r + t_s > n, 0 < t_b \leq n - f_b\) where \(f_b = t_r(n - t_r)/(t_r - t_s + 1)\), and at most \(f\) nodes fail, then TLCB (Algorithm 3) implements a TSB\((t_r, t_b, n)\) full-spread broadcast abstraction.

Under these configuration constraints, therefore, TLCB provides a TSB abstraction sufficient to support QSC (Section 4). This consensus algorithm supports the optimal \(2f + 1\) node count for the special case of \(f = 1\) and \(n = 3\), which in practice is an extremely common and important configuration. For larger \(f\), however, running QSC on TLCB requires \(n\) to grow faster than \(2f + 1\). This limitation motivates witnessed TLC, described next, which is slightly more complex but allows QSC to support an optimal \(n = 2f + 1\) configuration for any \(f \geq 0\).

5.2.2 Asymptotic complexity of TLCB

Implementing TLCB naïvely on naïvely-implemented TLCR yields a total communication complexity of \(O(n^4)\) per round if the messages passed to TLCB are of size \(O(1)\).

As discussed above in Section 5.1.1, however, this cost may be reduced by delaying the processing of messages for future time steps, or by using hash-references and an IHAVE/SENDME protocol on the point-to-point links. In this case, TLCR incurs a communication complexity of \(O(n^3)\) per round with \(O(1)\)-size messages, because the set \(R'\) in the second broadcast is not actually an \(O(n)\)-length list of \(O(n)\)-size messages, but is rather an \(O(n)\)-size list of \(O(1)\)-size hash-references to messages whose content each node receives only once.
5.3 TLCW: witnessed threshold logical clocks

TLCW (Algorithm 4) in essence extends TLCR (Section 5.1) so that each node $i$ works proactively in each round to ensure that at least $t_s$ nodes’ messages are received by at least $t_r$ nodes each, and waits until $i$ can confirm this fact before advancing to the next logical time-step.

TLCW accomplishes this goal by having each node run an echo broadcast protocol [19] in parallel, to confirm that its own message has been received by at least $t_s$ nodes, before its message is considered threshold witnessed and hence “counts” toward a goal of $t_s$ such messages. Variants of this technique have been used in other recent consensus protocols such as ByzCoin [55] and VABA [3]. As in TLCR, a slow node can also catch up to another node at a later timestep by reusing the set of threshold-witnessed messages that the latter node already used to advance logical time.

**Theorem 5.4.** If $0 < t_b \leq n - f$, $0 < t_s \leq n - f$, and at most $f$ nodes fail, then TLCW (Algorithm 4) implements a TSB($t_b, t_s, t_r$) partial-spread broadcast abstraction.

5.4 TLCF: full-spread threshold synchronous broadcast with TLCW and TLR

While TLCW directly implements only partial-spread threshold synchronous broadcast, similar to TLCB above we can “bootstrap” it to full-spread synchronous broadcast in configurations satisfying $t_r + t_s > n$. TLCF, shown in Algorithm 5, simply follows a TLCW round with a TLR round. By exactly the same logic as in TLCB, this ensures that each message in the broadcast set $B$ returned from TLCW propagates to every node that has not failed by the end of the subsequent TLR round, because all the broadcast-spread sets in TLCW overlap with all the receive-sets in the subsequent TLR.

**Theorem 5.5.** If $0 < t_r \leq n - f$, $0 < t_b \leq n - f$, $0 < t_s \leq n - f$, $t_r + t_s > n$, and at most $f$ nodes fail, then TLCF (Algorithm 5) implements a TSB($t_r, t_b, n$) full-spread broadcast abstraction.

6 On-demand client-driven implementation of TLC and QSC

In appropriate network configurations, QSC (Section 4.2) may be implemented atop either full-spread TLCB (Section 5.2.1) or TLR (Section 5.4). Supporting a fully-asynchronous underlying network, these combinations progress and commit consensus decisions continuously as quickly as network connectivity permits. Using the optimizations described in Sections 4.4, 5.1.1, and 5.2.2, these implementations incur expected total communication costs of $O(n^3)$ bits per successful consensus decision and QSC history delivery.

We would like to address two remaining efficiency challenges, however. First, in many practical situations we want consensus to happen not continuously but only on demand, leaving the network idle and consuming no bandwidth when there is no work to be done (i.e., no transactions to commit). Second, it would be nice if QSC could
atomic merely passive servers that implement only a locally-representing the actual consensus group members are above by implementing QSC and TLC in a client-driven architecture. In this instantiation, the n stateful nodes representing the actual consensus group members are merely passive servers that implement only a locally-atomic write-once key-value store.

With certain caveats, we can achieve both efficiency goals by implementing QSC and TLC in a client-driven architecture. In this instantiation, the n stateful nodes representing the actual consensus group members are merely passive servers that implement only a locally-atomic write-once key-value store.

Algorithm 4: TLCW(m), a witnessed threshold logical clock implementing TSB(t_b, t_b, t_s) synchronous broadcast

| Configuration | node number i, number of nodes n, broadcast threshold t_b ≤ n, spread threshold t_s ≤ n |
|--------------|----------------------------------------------------------------------------------|
| Configuration | functions Receive, Broadcast, Unicast representing underlying asynchronous network API |
| Persistent state | message receive log R and broadcast log B, each initialized to a singleton list [{}, {}] |
| Function input | message m to broadcast in this time step |
| Function output: | sets (R, B) of messages received, and reliably broadcast, in this time-step |

```
(\bar{R}, \bar{B}) \leftarrow (\bar{R} \cup \{\}) \cup (\bar{B} \cup \{}\})  
\text{// start a new logical time-step with empty receive and broadcast sets} 
N_A \leftarrow \{\}  
\text{// initially empty witness acknowledgment set for our message m} 
Broadcast((\text{req}, i, m, |\bar{R}|, |\bar{R}|_{-1}, |\bar{B}|_{-1}))  
\text{// broadcast our request, current time-step, and last message-sets} 
\text{while } |\bar{B}|_i < t_b \text{ do}  
\text{// loop until we reach broadcast threshold t_b to advance logical time} 
\text{switch Receive() do}  
\text{// receive the next message from any node} 
\text{case } \langle \text{req}, j, m', |\bar{R}|, \langle \rangle \rangle \text{ do}  
\text{// request message m' from node j in the same time-step} 
\text{Unicast}(j, \langle \text{ack}, i, m', |\bar{R}|, |\bar{R}|_{-1}, |\bar{B}|_{-1} \rangle)  
\text{// acknowledge node j’s request as a witness} 
\text{case } \langle \text{ack}, j, m, |\bar{R}|, \langle \rangle \rangle \text{ do}  
\text{// acknowledgment of our request m from node j} 
N_A \leftarrow N_A \cup \{j\}  
\text{// collect acknowledgments of our request message} 
\text{if } |N_A| = t_s \text{ then}  
\text{// our message has satisfied the spread threshold t_s} 
\text{Broadcast((\text{wit}, i, m, |\bar{R}|, |\bar{R}|_{-1}, |\bar{B}|_{-1} \rangle))}  
\text{// announce our message m as fully witnessed} 
\text{case } \langle \text{wit}, j, m', |\bar{R}|, \langle \rangle \rangle \text{ do}  
\text{// announcement that j’s message m' was witnessed by t_s nodes} 
\text{Unicast}(j, \langle \text{ack}, i, m', |\bar{R}|, |\bar{R}|_{-1}, |\bar{B}|_{-1} \rangle)  
\text{// collect fully-witnessed messages received in this time-step} 
\text{case } \langle \langle \rangle \rangle \text{ do}  
\text{// message m' is from the next step due to in-order channels} 
\text{(\bar{R} | \bar{R}|, \bar{B} | \bar{B}|) \leftarrow (\bar{R} | \bar{R}| \cup R', \bar{B} | \bar{B}| \cup B')}  
\text{// virally adopt message-sets that j used to advance history} 
\text{end} 
\text{end} 
\text{return } (\{m' | \langle j, m' \rangle \in |\bar{R}|_i\}, \{m' | \langle j, m' \rangle \in |\bar{B}|_i\})  
\text{// return the final message sets for this time-step} 
```

achieve the optimal lower bound of $O(n^2)$ communication complexity, at least in common-case scenarios.

6.1 Consensus over key-value stores

With certain caveats, we can achieve both efficiency goals above by implementing QSC and TLC in a client-driven architecture. In this instantiation, the n stateful nodes representing the actual consensus group members are merely passive servers that implement only a locally-atomic write-once key-value store.

Definition 6.1. A write-once store serves Write and Read requests from clients. A Write($K, V$) operation atomically writes value $V$ under key $K$ provided no value exists yet in the store under key $K$, and otherwise does nothing. A Read($K$) → $V$ operation returns the the value written under key $K$, or empty if no value has been written yet under key $K$.

In practice the $n$ servers can be any of innumerable distributed key/value stores supporting locally-atomic writes [51, 73, 80, 83]. The $n$ servers might even be standard Unix/POSIX file systems, mounted on clients via
To read past consensus history from each server \( i \) and
catch up to its current state, a client’s local state-machine
simulation thread \( i \) simply reads keys from \( i \) in their
well-defined sequence, replaying the QSC/TLC state machine
defined by their values at each transition, until the client
encounters the first key not yet written on the server.
Clients that are freshly started or long out-of-date can
catch up more efficiently using optimizations discussed later
in Section 6.4.

6.2 Representing the QSC/TLC state machine

Each of the \( n \) key/value stores implicitly represents the
current state of that node’s QSC/TLC state machine, as
simulated by the clients. One transition in each server’s
state machine is represented by exactly one atomic key/value
Write. All keys ever used on a node inhabit a well-
defined total order across both TLC time-steps and state
transitions within each step. To track the \( n \) servers’ con-
sensus states and drive them forward, each client locally runs \( n \) concurrent threads or processes, each simulating
the state machine of one of the servers.

To advance consensus state, each client’s \( n \) simulation
threads coordinate locally to decide on and (attempt to)
write new key-value pairs to the servers, representing non-
deterministic but valid state transitions on those servers.
Each of these writes may succeed or fail due to races with
other clients’ write attempts. In either case, the client ad-
vances its local simulation of a given server’s state ma-
chine only after a read to the appropriate key, i.e., accord-
ing to the state transition defined by whichever client won
the race to write that key.

We outline only the general technique here. QSCOD,
Algorithm 6 in Appendix C, presents pseudocode for a
specific example of a client-driven on-demand implemen-
tation of QSC over TLCB.

6.3 Complexity analysis

Implementing QSC over TLCB in this way in QSCOD, a
client that is already caught up to the servers’ states in-
curs \( O(n^2) \) expected communication bits to propose and
reach agreement on a transaction. This is because the
client reads and writes only a constant number of \( O(n) \)-
size messages to the \( O(n) \) servers per consensus round,
and QSC requires a constant expected number of rounds
to reach agreement.

The client eliminates the need for broadcasts by effec-
tively serving in a “natural leader” role, analogous to an
elected leader in Paxos — but without Paxos’s practical
risk of multiple leaders interfering with each other to halt
progress entirely. When multiple QSCOD clients “race”

---

\(^2\) A standard way to implement Write atomically on a POSIX file
system is first to write the contents of \( V \) to a temporary file (ensuring that
a partially-written file never exists under name \( K \)), attempt to hard-link the
temporary file to a filename for the target name \( K \) (the POSIX \texttt{link}
operation fails if the target name already exists), and finally unlink the
temporary filename (which deletes the file if the \texttt{link} operation failed).
to drive the servers’ implicit QSC/TLC state machines concurrently in one consensus round, the round still progresses normally and completes with the same (constant) success probability as with only one client.

Under such a time of contention, only one’s client’s proposal can “win” and be committed in the round, of course. Since in transactional applications clients whose proposals did not win may need to retry, this contention can increase communication costs and server load, even though the system is making progress. Since each consensus round is essentially a shared-access medium, one simple way to mitigate the costs of contention is using random exponential backoff, as in the classic CSMA/CD algorithm for coaxial Ethernet [49]. Another approach is for clients to submit transactions to a gossip network that includes a set of intermediating back-end proxies, each of which collect many clients’ transactions into blocks to propose and commit in batches, as in Bitcoin [72]. This way, it does not matter to clients which proxy’s proposal wins a given round provided some proxy includes the client’s transaction in a block.

### 6.4 Implementation optimizations

The above complexity analysis assumes that a client is already “caught up” to the servers’ recent state. Implementations can enable a freshly-started or out-of-date client can catch up efficiently, with effort and communication logarithmic rather than linear in the history size, by writing summaries or “forward pointers” to the key-value stores enabling future clients to skip forward over exponentially-increasing distances, as in skip lists [79] or skipchains [74].

### 7 Limitations and Future Work

The QSC and TLC protocols developed here have many limitations, most notably tolerating only crash-stop node failures [19,89]. It appears readily feasible to extend QSC and TLC to tolerate Byzantine node failures along lines already proposed informally [41]. Further, it seems promising to generalize the principles of QSC and TLC to support quorum systems [20,45,63,65], whose threat models assume that not just a simple threshold of nodes, but more complex subsets, might fail or be compromised. Full formal development and analysis of Byzantine versions of QSC and TLC, however, remains for future work.

Most efficient asynchronous Byzantine consensus protocols rely on threshold secret sharing schemes [91,92,94] to provide shared randomness [17,21,96] and/or efficient threshold signing [3,97]. Setting up these schemes asynchronously without a trusted dealer, however, requires distributed key generation or DKG [15,52,56,103]. The TLC framework appears applicable to efficient DKG as well [41], but detailed development and analysis of this application of TLC is again left for future work.

While this paper focuses on implementing consensus in a fashion functionally-equivalent to [Multi-]Paxos or Raft, it remains to be determined how best to implement closely-related primitives such as atomic broadcast [19,29,33,66] in the TLC framework. For example, QSC as formulated here guarantees only that each round has a reasonable chance of committing some node’s proposal in that round – but does not guarantee that any particular node’s proposals have a “fair” chance, or even are ever, included in the final total order. Indeed, a node that is consistently much slower than the others will never see its proposals chosen for commitment. An atomic broadcast protocol, in contrast, should guarantee that all messages submitted by any correct node are eventually included in the final total order. The “fairness” or “eventual-inclusion” guarantees required for atomic broadcast are also closely-related to properties like chain quality recently explored in the context of blockchains [10,76,77].

While the algorithms described above and their fundamental complexity-theoretic characteristics suggest that QSC and TLC should yield simple and efficient protocols in practice, these properties remain to be confirmed empirically with fully-functional prototypes and rigorous experimental evaluation. In particular, we would like to see systematic user studies of the difficulty of implementing QSC/TLC in comparison with traditional alternatives, similar to the studies that have been done on Raft [47,75]. In addition, while we have decades of experience optimizing implementations of Paxos for maximum performance and efficiency in deployment environments, it will take time and experimentation to determine how these lessons do or don’t translate, or must be adapted, to apply to practical implementations of QSC/TLC.
8 Related Work

This section summarizes related work, focusing first on TLC in relation to classic logical clocks, and then on QSC in relation to other consensus protocols, first asynchronous and then those specifically designed with simplicity in mind.

Logical clocks and virtual time TLC is inspired by classic notions of logical time, such as Lamport clocks [57, 82], vector clocks [36, 37, 39, 62, 67] and matrix clocks [34, 82, 87, 88, 102]. Prior work has used logical clocks and virtual time for purposes such as discrete event simulation and rollback [50], verifying cache coherence protocols [78], and temporal proofs for digital ledgers [48]. We are not aware of prior work defining a threshold logical clock abstraction or using it to build asynchronous consensus, however.

Conceptually analogous to TLC, Awerbuch’s synchronizers [7] are intended to simplify the design of distributed algorithms by presenting a synchronous abstraction atop an asynchronous network. Awerbuch’s synchronizers assume a fully-reliable system, however, tolerating no failures in participating nodes. TLC’s purpose might therefore be reasonably described as building fault-tolerant synchronizers.

The basic threshold communication patterns TLC employs have appeared in numerous protocols in various forms, such as classic reliable broadcast [13, 14, 84]. Witnessed TLC is inspired by threshold signature schemes [12, 93], signed echo broadcast [3, 16, 84], and witness cosigning protocols [74, 97]. We are not aware of prior work to develop or use a form of logical clock based on these threshold primitives, however, or to use them for purposes such as asynchronous consensus.

Asynchronous consensus protocols The FLP theorem [38] implies that consensus protocols must sacrifice one of safety, liveness, asynchronous, or determinism. QSC sacrifices determinism and implements a probabilistic approach to consensus. Randomness has been used in consensus protocols in various ways: Some use private coins that nodes flip independently but require time exponential in group size [8, 13, 70], assume that the network embodies randomness in the form of a fair scheduler [14], or rely on shared coins [3, 9, 16–18, 21, 27, 28, 35, 42, 69, 71, 81, 96]. Shared coins require complex setup protocols, however, a problem as hard as asynchronous consensus itself [15, 52, 56, 103]. QSC in contrast requires only private randomness and private communication channels.

QSC’s consensus approach, where each node maintains its own history but adopts those of others so as to converge statistically, is partly inspired by randomized blockchain consensus protocols [2, 43, 53, 72], which rely on synchrony assumptions however. QSC in a sense provides Bitcoin-like consensus using TLC for fully-asynchronous pacing and replacing Bitcoin’s “longest chain” rule with a “highest priority” rule.

Consensus protocols designed for simplicity Consensus protocols, such as the classic (Multi-)Paxos [58], are notoriously difficult to understand, implement, and reason about. This holds especially for those variants that run atop asynchronous networks, can handle Byzantine faults, or try to tackle both [17, 18, 22, 69, 70]. (Multi-)Paxos, despite being commonly taught and used in real-world deployments, required a number of additional attempts to clarify its design and further modifications to adapt it for practical applications [25, 31, 54, 59, 60, 68, 99]. The intermingling of agreement and network synchronization appears to be a source of algorithmic complexity that has not been addressed adequately in past generations of consensus protocols, resulting in complex leader-election and view-change (sub-)protocols and restrictions to partial synchrony [47, 75].

In its aim for simplicity and understandability, QSC is closely related to Raft [75], which however assumes a partially-synchronous network and relies on a leader. QSC appears to be the first practical yet conceptually simple asynchronous consensus protocol that depends on neither leaders nor common coins, making it more robust to slow leaders or network denial-of-service attacks. The presented approach is relatively clean and simple in part due to the decomposition of the agreement problem (via QSC) from that of network asynchrony (via TLC).

9 Conclusion

This paper has presented QSC, the first asynchronous consensus protocol arguably simpler than current partially-
synchronous workhorses like Paxos and Raft. QSC requires neither leader election, view changes, nor common coins, and cleanly decomposes the consensus problem itself from that of handling network asynchrony. With appropriate implementation optimizations, QSC completes in $O(1)$ expected rounds per agreement, incurring $O(n^3)$ communication bits in a broadcast-based group, or $O(n^2)$ bits per client-driven transaction in an on-demand implementation approach.

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Appendix

A Correctness Proofs

This appendix contains the proofs for the theorems in the main paper.

A.1 Que Sera Consensus (QSC)

This section contains correctness proofs for the QSC consensus algorithm (Section 4).

Lemma 4.1. History preservation: If a consensus round starting at time-step $s$ has initial history $h_s$, on node $i$, then at any earlier consensus round starting at step $s' < s$, there exists some node $j$ whose initial history $h_{s,j}$ in that round is a strict prefix of $h_s$.

Proof. Let $h_s$ be the initial history of the round starting at step $s$ on node $i$, let $h_{s,i}'$ and $h_{s,i}''$ be $i$’s proposed and intermediate histories in that round, respectively, and let $(B_{s,i}', B_{s,i}''')$ and $(B_{s,i}', R_{s,i}'')$ be the sets returned by the round’s two Broadcast calls. By QSC’s requirement that $t_b > 0$ and the TSB’s broadcast threshold property (Section 3), the sets $B_{s,i}'$ and $B_{s,i}''$ returned by the round’s two Broadcast calls are nonempty. By TSB’s receive threshold property, the returned sets $R_{s,i}'$ and $R_{s,i}''$ are nonempty as well. By message propagation through these Broadcast calls, these sets consist solely of histories $h_{s,j}'$ each proposed in the same round by some node $j$, and each of which builds on $j$’s initial history $h_{s,j}$. By induction over consensus rounds, therefore, at step $s$ each node $i$’s history $h_s$, builds on some node $j$’s history $h_{s,j}$ at each earlier step $s' < s$. That is, $h_{s,j}$ is a strict prefix of $h_s$.

Lemma 4.2. Agreement on delivery: If QSC delivers history $h_{(s+2),i}$, on node $i$ at the end of a consensus round starting at time-step $s$, then the resulting history $h_{(s+2),j}$ of every node $j$ in the same round is identical to $h_{(s+2),i}$.

Proof. Agreement can be violated only if some node $j$ arrives at a different resulting history $h_{(s+2),j}$.

Because $h_{(s+2),i} \in B_{s,i}''$, and $B_{s,i}'' \subseteq R_{s,i}'$, by TSB’s broadcast spread property, $i$’s delivered history $h_{(s+2),i}$ is also among the set of histories from which $j$ chooses its resulting (but not necessarily delivered) history $h_{(s+2),j}$. Because $j$ chooses some best history from set $R_{s,j}'$, $h_{(s+2),j}$ cannot have strictly lower priority than $h_{(s+2),i}$, otherwise $j$ would instead choose $h_{(s+2),i}$. So we can subsequently assume that the priority of $h_{(s+2),j}$ is greater than or equal to that of $h_{(s+2),i}$.

Every history occurring in $j$’s set $R_{s,j}'$, however, is a proposal derived (via the round’s second Broadcast call) from a member of some set $B_{s,k}'$ that the first Broadcast call returned to some node $k$. Because $B_{s,k}' \subseteq R_{s,j}'$, by TSB's broadcast spread property, both $h_{(s+2),i}$ and $h_{(s+2),j}$ must therefore also appear in $R_{s,j}'$. But then $h_{(s+2),j}$ cannot be uniquely best in $R_{s,j}'$, satisfying the second condition on $i$ delivering $h_{(s+2),j}$, unless $h_{(s+2),j} = h_{(s+2),i}$.

Lemma 4.3. If the network delivery schedule is independent of proposal priorities and $p_t$ is the probability that two nodes tie for highest priority, then each node delivers a history in each round independently with probability at least $t_b/n - p_t$.

Proof. We will show that in the absence of a tie for best priority, node $i$’s probability of successfully finalizing a round is $t_b/n$. Since a round without a tie thus fails with probability at most $1 - t_b/n$, by the Union Bound, the overall probability of round failure is at most $1 - t_b/n + p_t$.

Let $N_{B_{s,i}'}$, $N_{R_{s,i}'}$, $N_{B_{s,i}''}$, $N_{R_{s,i}''}$ each be the subsets of nodes $\{1, \ldots, n\}$ whose messages $i$’s broadcast calls returned in its respective sets $B_{s,i}'$, $R_{s,i}'$, $B_{s,i}''$, $R_{s,i}''$ (Definition 3.1). By the above independence assumption, the network adversary’s choices of these sets does not depend on the content of messages or their priority values.
If the set $B''_{s_j}$ returned from $i$’s second broadcast contains the round’s unique globally-best history $\hat{h}_s$, which exists due to our exclusion of ties above, then $i$ will necessarily choose $\hat{h}_s$ and deliver it. This is because $\hat{h}_s$ must also be in $R''_s$, and in $R''_s$, and no other proposal exists in either set with priority greater than or equal to that of $\hat{h}_s$.

This desirable event that $\hat{h}_s \in B''_{s_j}$ occurs if at least one node $j \in N_{B''_{s_j}}$ chose $\hat{h}_s$ as its intermediate history $h''_s$, and broadcast it in $j$’s second call to Broadcast. Since the probability of this event occurring for any specific node $j \in N_{B''_{s_j}}$, we now conservatively focus on analyzing this probability of any specific such node $j \in N_{B''_{s_j}}$ choosing $\hat{h}_s$.

If the set $B'_{s_j}$ returned from $j$’s first broadcast contains the round’s unique globally-best history $\hat{h}_s$, then $j$ will necessarily choose $h''_s = \hat{h}_s$ and broadcast it in $j$’s second Broadcast call. This desirable event occurs in turn if $N_{B'_{s_j}}$ includes the node $k$ that proposed the unique globally-best history $\hat{h}_s$ in this round. Since all nodes choose their priorities from the same random distribution, each node has an equal chance of proposing the globally-best history $\hat{h}_s$. Since $|N_{B'_{s_j}}| \geq t_b$, node $j$ therefore sees $\hat{h}_s$ in its set $B'_{s_j}$ with a probability of at least $t_b/n$.

Node $i$ therefore sees $\hat{h}_s$ in its set $B''_{s_j}$ and delivers a history in this round with a probability of at least $t_b/n$.

**Theorem 4.1.** QSC implements multi-consensus on $n$ nodes (Definition 4.1) atop a full-spread TSB primitive $TSB(t_r, t_b, n)$ where $t_r > 0$ and $t_b > 0$.

**Proof.** Liveness: QSC regularly advances time forever on non-failing nodes by calling Broadcast twice each time through an infinite loop, at each step delivering a history with some independent nonzero probability (Lemma 4.3). These delivered histories grow in length by one message each time through the loop. Therefore, if $h$ is the longest history delivered by time-step $s$ on a non-failing node $i$, then with probability 1 there is eventually some future time-step $s' > s$ at which node $i$ delivers a longer history $h' (|h'| > |h|)$, thereby satisfying liveness.

Validity: If QSC invokes Deliver($h'||p)$ at step $s'$ on node $j$, then by the TSB receive threshold property $p$ is a proposal (proposal $i, m, r$) that some node $i$ appended to its internal history $h_i$ and broadcast at step $s = s' - 2$, at the beginning of the same QSC round (main loop iteration).

Consistency: If QSC delivers $h$ at step $s$ on node $i$, then delivers $h'$ at step $s' \geq s$ on node $j$, then by induction over $s' - s$, using Lemma 4.2 as the base case, and using Lemma 4.1 in the inductive step, $h$ must be a prefix of $h'$.

### A.2 Threshold Logical Clocks (TLC)

This section contains correctness proofs for the threshold logical clock algorithms in Section 5.

**Theorem 5.1.** TLCR (Algorithm 2) implements a TSB($t_r, t_b, n$) communication primitive with receive threshold $0 \leq t_r \leq n$, provided at most $f \leq n - t_r$ nodes fail.

**Proof.** Provided TLCR terminates, it satisfies the TSB’s lock-step synchrony property (Definition 3.1), because $|\hat{R}|$ represents the current time-step at each invocation counting from 2, and each TLCR call adds exactly one element to $\hat{R}$. Because at most $f \leq n - t_r$ nodes can fail, each non-failed node eventually receives a threshold $t_r$ of messages from the $t_r$ non-failed nodes at each time-step, ensuring that each TLCR call eventually terminates and successfully advances logical time.

TLCR satisfies the TSB receive threshold property by construction, i.e., by not returning until it accumulates and returns a receive-set $R$ of size at least $t_r$ or until it obtains such a set $R$ all at once by catching up to another node via a message from a future time-step. Because messages are pairwise-ordered between nodes, the condition $s' > |\hat{R}|$ implies $s' = s + 1$. Because the internal receive-sets consist of pairs $(j, m')$ representing the sending node $j$ and message $m'$ that node $j$ broadcast in the same time-step, the returned set $R$ contains messages sent by at least $t_r$ nodes even if multiple nodes send the same message, ensuring that the required node-set $N_R$ exists (Definition 3.1).

TLCR trivially satisfies the broadcast threshold and broadcast spread properties in Definition 3.1 by always returning an empty broadcast set $B$, thereby making no broadcast threshold promises to be fulfilled.
Theorem 5.2. If $0 < t_r \leq n - f$, $0 < t_s \leq t_r$, $0 < t_b \leq n - f_b$ where $f_b = t_r(n - t_r)/(t_r - t_s + 1)$, and at most $f$ nodes fail, then TSB (Algorithm 3) implements a TSB$(t_r, t_b, t_s)$ partial-spread broadcast abstraction.

Proof. In the second TLCR call, each node $i$ collects at least $t_r$ nodes’ receive-sets from the first TLR call, each of which contains at least $t_r$ nodes’ first-round messages. We represent node $i$’s observations as a view matrix with $t_r$ rows (one per-receive set) and $n$ columns (one per node), such that each cell $j, k$ contains 1 if $i$’s receive-set $j$ indicates receipt of node $k$’s message from the first TLCR round, and 0 otherwise.

Node $i$’s $t_r \times n$ view matrix contains at least $t_r^2$ one bits, and hence at most $t_r(n - t_r)$ zero bits. To prevent $t_b$ nodes’ messages from reaching at least $t_s$ nodes each in $i$’s view, the network must schedule the deliveries seen by $i$ so that at least $n - t_b + 1$ columns of $i$’s view matrix each fail to contain at least $t_s$ one bits. Each such failing column must contain at least $t_r - t_s + 1$ zero bits. Since there are at most $t_r(n - t_r)$ zero bits total, there can be at most $f_b = t_r(n - t_r)/(t_r - t_s + 1)$ failing columns. The matrix must therefore have at least $n - f_b$ non-failing columns representing reliable broadcasts to at least $t_s$ nodes each. TLCB therefore satisfies the required broadcast threshold $t_b$ since $t_b \leq n - f_b$.

Theorem 5.3. If $0 < t_r \leq n - f$, $0 < t_s \leq t_r$, $t_r + t_s > n$, $0 < t_b \leq n - f_b$ where $f_b = t_r(n - t_r)/(t_r - t_s + 1)$, and at most $f$ nodes fail, then TLCB (Algorithm 3) implements a TSB$(t_r, t_b, t_s)$ full-spread broadcast abstraction.

Proof. By Theorem 5.2, the returned broadcast set $B$ contains the messages sent by at least $t_b$ nodes in the first TLCR step. Consider any such message $m \in B$ and any node $i$ that completes this TLCB step without failing.

By construction, node $i$’s set $B$ contains only messages $i$ knows have been received by at least $t_s$ nodes. Therefore, there is some set $N_s \subseteq \{1, \ldots, n\}$ of nodes such that $|N_s| \geq t_s$, and for each node $j \in N_s$, the intermediate receive set $R'_j$ on node $j$ contains $m$.

Further, due to the receive threshold $t_r$ enforced by TLCR, the message set $R''$ returned on node $i$ must contain the intermediate message sets $R'$ that were returned on at least $t_r$ nodes. That is, there is some set $N_r \subseteq \{1, \ldots, n\}$ of nodes such that $|N_r| \geq t_r$, and for each node $j \in N_r$, the intermediate receive set $R'_j$ returned on node $j$ is a subset of $R''$ on node $i$.

Because $t_r + t_s > n$, the sets $N_s$ and $N_r$ must therefore overlap by at least one node $k$. Node $k$ therefore received message $m$ in its intermediate set $R''$, and thus in turn must have passed $m$ on to $i$ via the second TLCR step. Therefore, message $m$ must be in the receive set finally returned by TLCB on node $i$. Since this applies to all messages $m \in B$ and all nodes $i$, TLCB therefore implements TSB$(t_r, t_b, t_s)$ full-spread synchronous broadcast.

Theorem 5.4. If $0 < t_b \leq n - f$, $0 < t_s \leq n - f$, and at most $f$ nodes fail, then TLCW (Algorithm 4) implements a TSB$(t_b, t_b, t_s)$ partial-spread broadcast abstraction.

Proof. TLCW satisfies the TSB’s lock step synchrony because (a) each call to TLCW only adds one element to $R$, which represents the current time-step, if it terminates, and (b) because at most $f \leq n - t_b$ nodes can fail, each non-failed node eventually receives at least $t_b$ messages from the non-failed nodes guaranteeing that each call to TLCW eventually terminates and advances the logical time. TLCW satisfies both the broadcast threshold and broadcast spread properties by construction. Specifically, TLCW does not return until it accumulates and returns a broadcast-set $B$ of size at least $t_b$ or until it obtains such a set by catching up to another node via a message from a future time step. The returned broadcast-set $B$ consists of at least $t_b$ fully witnessed messages $m'$ (satisfying broadcast threshold), where each node $j$ announces that $(j, m')$ has been fully witnessed only after its message $m'$ was acknowledged by $t_s$ nodes (satisfying broadcast spread), given that at most $f$ nodes can fail and $t_b, t_s \leq n - f$. Finally, since $B \subseteq R$, we get $t_r \geq t_b$.

Theorem 5.5. If $0 < t_r \leq n - f$, $0 < t_b \leq n - f$, $0 < t_s \leq n - f$, $t_r + t_s > n$, and at most $f$ nodes fail, then TLCF (Algorithm 5) implements a TSB$(t_r, t_b, t_s)$ full-spread broadcast abstraction.

Proof. The proof is identical in essence to that of Theorem 5.3.

B QSC model in Erlang

To illustrate QSC more concretely, this section lists a full working model implementation of QSC atop TLCB and
Erlang is particularly well-suited to modeling QSC, being a distributed functional programming language with a concise syntax. As a result, the actual working Erlang code is not much longer in line count than the pseudocode in Algorithms 1.2, and 3 that it implements.

Erlang’s selective receive capability [101], in particular, simplifies implementation of TLCR. Selective receive allows TLCR to receive messages for the current step-time and discard messages arriving late for past time-steps, while saving messages arriving early for future time-steps in the process’s mailbox for later processing.

B.1 qsc.erl: Erlang code listing

```erlang
-module(qsc).
-export([qsc/1, test/1]).

% Node configuration is a tuple defined as a record.
-record(node, {tr, ts, Pids, choose, random, deliver}).

% A history is a record representing the most recent in a chain.
-record(history, {nop, step, Bids}).
% save as list of messages received.
-compile(export_def, test).

% qsc(C) -> test (C) (see tlc).
qsc(C) -> qsc(C, !node, msg).% start at step 0 with placeholder msg.

% Launch a process representing each of the N nodes.
launch = fun(# config {node0, choose0, random0, deliver0}) ->
    Choose = fun(# config {node0}),% Function to choose message for node I to propose at TLC time
    Random = fun(C, S, H) ->
        % Receive a config record C and run QSC with that configuration.
        RunQSC = fun()
        % Launch a process representing each of the N nodes.
        Run = spawn(fun(options), C, [], {node, node0}).
        {node, Run}.
    end.
    Deliver = fun(C, S, H) ->
        % Return N the number of lists in list 
        %tlcr(C, S, M).
        {node, Result} = C, S, H.
        spawn_scenario. % Spawn a tester process
        % Stop testing if a node sends 
        %tlcr(C, S, M).
        when length(Result) == N.
        {node, Result}.
    end.
    Tester = fun(F)
    ->
        % Run QSC and TLC through a test suite.
        test(F, Parent, Steps) ->
            % Generate a random valid configuration from number of failures F.
            F = random:uniform(F).
            % Function to choose message for node I to propose at TLC time
            % Choose a random value to attach to a proposal in time.
            Random = fun(C, S, H) ->
                % Receive a config record C and run QSC with that configuration.
                RunQSC = fun()
                % Launch a process representing each of the N nodes.
                Run = spawn(fun(options), C, [], {node, node0}).
                {node, Run}.
            end.
            Deliver = fun(C, S, H) ->
                % Return N the number of lists in list 
                %tlcr(C, S, M).
                {node, Result} = C, S, H.
                spawn_scenario. % Spawn a tester process
                % Stop testing if a node sends 
                %tlcr(C, S, M).
                when length(Result) == N.
                {node, Result}.
            end.
            Tester = fun(F)
            ->
                % Run QSC and TLC through a test suite.
                test(F, Parent, Steps, # hist []) ->
                    % Wait for a test to finish and consistency—check the results: it connects
                    %tlcr(C, S, M).
                    {node, Result} = C, S, M.
                    % Save main process ‘s PID
                    spawn_scenario. % Spawn a tester process
                    % Stop testing if a node sends 
                    %tlcr(C, S, M).
                    when length(Result) == N.
                    {node, Result}.
            end.
        end.
    Tester.
end.
```

## C QSCOD: Client-driven on-demand QSC with TLCB

To provide a concrete illustration of the on-demand approach to implementing QSC and TLC outlined in Section 6, Algorithm 6 shows pseudocode for client-driven QSC built atop full-spread TLCB (Section 5.2.1).

In QSCOD, each client node wishing to submit pro-

### TLKR in Erlang [5]. The model implements nodes as Erlang processes interacting via message passing, in less than 73 code lines as counted by cloc [32]. Of these, only 37 code lines comprise the consensus algorithm itself, the rest representing test framework code.
Algorithm 6: QSCOD: client-driven execution of QSC over TLCB

Configuration: node number $i$ this thread drives, number of nodes $n$, thresholds $t_r \leq n$ and $t_s \leq n$

Configuration: functions RandomValue, Write$_i$, Read$_i$, Deliver

Global state: client cache $C$ of servers’ state; $C_{j,k}$ holds value stored on server $j$ under key $k$ if known

Synchronization: WaitMessage() → $m$ waits for and returns the next message this client wishes to commit

Synchronization: WaitCache($k$) → $R$ waits to collect and return set $R$ of cached values $C_{j,k}$ from $\geq t_r$ nodes

```
q ← 1
h ← {}
// fictitious initial round
m ← WaitMessage()
// loop until any client thread determines $m$ was committed
forever

q ← q + 1
r ← RandomValue()
// empty history for fictitious initial round
h′ ← $h_c$ ← ⟨Hash($h$), $m$, $r$⟩
Write$_i$(⟨$q$, 1⟩, ⟨$h$, $h'$⟩)
⟨$h$, $h'$⟩ ← $C_{i,(q,1)}$ ← Read$_i$(⟨$q$, 1⟩)
$R'_1$ ← WaitCache(⟨$q$, 1⟩)
Write$_i$(⟨$q$, 2⟩, $R'_1$)
$R'_2$ ← $C_{i,(q,2)}$ ← Read$_i$(⟨$q$, 2⟩)
$R'_1$ ← WaitCache(⟨$q$, 2⟩)
$R_1$ ← $\bigcup\{R'_1 \cup R''_1\}$
// tentative receive-set return from first TLCB
$B_1 ← \{m' | \text{at least } t_s \text{ messages-sets in } R''_1 \text{ contain } m'\}$
// tentative broadcast-set return from first TLCB
$h''$ ← any best history in $B_1$
Write$_i$(⟨$q$, 3⟩, ⟨$R_1$, $B_1$, $h''$⟩)
⟨$R_1$, $B_1$, $h''$⟩ ← $C_{i,(q,3)}$ ← Read$_i$(⟨$q$, 3⟩)
$R'_2$ ← WaitCache(⟨$q$, 3⟩)
Write$_i$(⟨$q$, 4⟩, $R'_2$)
$R'_2$ ← $C_{i,(q,4)}$ ← Read$_i$(⟨$q$, 4⟩)
$R'_2$ ← WaitCache(⟨$q$, 4⟩)
$R_2$ ← $\bigcup\{R'_2 \cup R''_2\}$
// tentative receive-set return from second TLCB
$B_2 ← \{m' | \text{at least } t_s \text{ messages-sets in } R''_2 \text{ contain } m'\}$
// tentative broadcast-set return from second TLCB
$h$ ← any best history in $R_2$
if $h = h_c$ and $h$ ∈ $B_2$ and $h$ is uniquely best in $R_2$ then

   $\text{Deliver}(h)$

   // history $h$ has no competition

end

m ← WaitMessage()
// wait for next message this client wishes to commit
```

proposals and drive consensus locally runs $n$ concurrent instances of Algorithm 6, typically in separate threads, one for each of the $n$ servers providing key-value stores. Initially and after each successful commitment of a client’s proposal, the client invokes WaitMessage to wait for the next message to submit as a proposal. The client may be quiescent for arbitrarily long in WaitMessage, during which the client produces no interaction with the servers.
(but other clients can propose messages and drive consensus in the meantime).

When `WaitMessage` returns the next message `m` to be committed, each client thread actively drives the key-value state of its respective server forward – in local cooperation with other client threads driving other servers – to complete as many consensus rounds as necessary to commit the client’s proposed message `m`.

Since each QSC round invokes TCR twice, which in turn invokes TCB twice, each consensus round requires four TCR time-steps. We could model each TCR round as having `t_r` + 1 state transitions: one representing a given node `i`’s initial broadcast, the rest for each of the `t_r` messages subsequently “received” by `i` as its condition to advance logical time. It is possible and more efficient, however, to summarize the effects of all simulated message “receives” in a time-step as part of the Write representing the node’s next broadcast. With this approach, Algorithm 6 requires only four pairs of Write/Read requests to each server per consensus round, one pair for each of the four total TCR invocations.

Coordination between the simulated consensus nodes occurs via the client’s locally-shared cache `C` of key-value pairs that have been read from the `n` servers so far. After attempting to write a value to a key, then reading back that key to learn what value was actually written by the “winning” client, each client thread invokes `WaitCache` to wait until `t_r` total threads also write and read corresponding values for that key. The client thread representing node `i` takes this locally-determined set as a tentative, possible receive-set of size `t_r` for node `i` – but neither this nor anything computed from it may be considered “definite” until the next Write/Read pair commencing the next TCR time-step.

Each client-side instance of QSCOD evaluates the QSC finality conditions, deciding whether the client’s message has been successfully committed, based on tentative information not yet finalized on the corresponding server. This may seem like a problem, but is not. Like any actual broadcast-based server implementation of QSC, a client thread will observe the finality conditions for history `h` only when it is “inevitable” that all servers commit history `h` – regardless of whether or not they know that `h` is committed. A client thread might observe that `h` is final, deliver it to the application, then lose a race to commit that result to the server at the start of the next time-step – but this means only that the simulated server does not “know” that `h` is committed, even though the client in question (correctly) knows this fact.

## C.1 Model QSCOD implementation in Go

To illustrate the operation of QSCOD more concretely, this section finally presents a simple but fully-functional model implementation of QSCOD in the Go language. The model implements nodes as goroutines communicating via shared memory instead of via real network connections, and is only 200 code lines as counted by `cloc` [32] including test infrastructure (less than 125 lines without). Despite its simplicity and limitations, this model implements all the fundamental elements of QSCOD, and can operate in truly distributed fashion by filling in core for remote access to key/value stores representing the consensus nodes, including the marshaling and unmarshaling of stored values. The latest version of this model may be found at https://github.com/dedis/tlc/tree/master/go/model/qscod.

## C.2 clii.go: QSCOD client model

```go
// Package qsc implements a simple model version of the QSC algorithm
// for client-drawn "on-demand" consensus.
package qsc

import "sync"

// Store represents an interface to one of the n key/value stores
// representing the persistent state of each of the n consensus group members.
// A Store's keys are integer TLC time-steps, and its values are Val structures.

type Store interface {
    WriteRead(Step, Val) Val // Write if no value yet, then read
}

// WaitMessage returns the next message m to be committed, each client thread actively drives the key-value state
// of its respective server forward – in local cooperation with other client threads driving other servers –
// to complete as many consensus rounds as necessary to commit the client’s proposed message m.
// Since each QSC round invokes TCR twice, which in turn invokes TCB twice, each consensus round requires four TCR time-steps.
// We could model each TCR round as having t_r + 1 state transitions: one representing a given node i’s initial broadcast,
// the rest for each of the t_r messages subsequently “received” by i as its condition to advance logical time.
// It is possible and more efficient, however, to summarize the effects of all simulated message “receives” in a time-step
// as part of the Write representing the node’s next broadcast. With this approach, Algorithm 6 requires only four pairs of Write/Read requests to each server per consensus round, one pair for each of the four total TCR invocations.

// Coordination between the simulated consensus nodes occurs via the client’s locally-shared cache C of key-value pairs
// that have been read from the n servers so far.
// After attempting to write a value to a key, then reading back that key to learn what value was actually written by
// the “winning” client, each client thread invokes WaitCache to wait until t_r total threads also write and read corresponding
// values for that key.
// The client thread representing node i takes this locally-determined set as a tentative, possible receive-set of size t_r for
// node i – but neither this nor anything computed from it may be considered “definite” until the next Write/Read pair
// commencing the next TCR time-step.

// Each client-side instance of QSCOD evaluates the QSC finality conditions, deciding whether the client’s message
// has been successfully committed, based on tentative information not yet finalized on the corresponding server.
// This may seem like a problem, but is not. Like any actual broadcast-based server implementation of QSC, a client thread
// will observe the finality conditions for history h only when it is “inevitable” that all servers commit history h –
// regardless of whether or not they know that h is committed. A client thread might observe that h is final, deliver it to
// the application, then lose a race to commit that result to the server at the start of the next time-step – but
```

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// Client represents one logical client issuing transactions // to the consensus group and driving the QSCODE state machine forward // asynchronously across the n key/value stores.

// Start starts a Client with given configuration parameters.
// Returns the history that successfully committed msg.
// Defined message msg, repeatedly if necessary,
// by whatever client won this race to initiate TLCB at step s.
// The provided v0 represents a potential next state for this node,
// but other clients may of course race with this one to set the next state.
// The returned Val represents the next—state value successfully registered
// by whatever client won this race to initiate TLCB at step s.
// The returned R and B sets, in contrast, are tentative.
// representing possible threshold receive—set and broadcast—set outcomes
// from this TLCB invocation, computed locally by this client.
// These locally—computed sets cannot be relied on to be definite for this node
// until the values computed from them are committed via Store.WriteRead.

// Create our key/value cache map for step s if not already created
// or return from TLCB:

// Try to write potential value v, then read that of the client who won
// that this Client would like to commit, and invoke TLCB to (try to) issue that proposal on this node.
// We see emerging from the first TLCB instance,
// we invoke TLCR to (try to) record the desired next state value, and register the definite winning value and a tentative receive—set.

// Final TLCB to (try to) record the desired next—state value,
// and record the definite winning value and a tentative receive—set.

// For the provided v0 represents a potential next state for this node,
// but other clients may of course race with this one to set the next state.

// Try to write potential value v, then read that of the client who won
// the provided v0 represents a potential next state for this node,
// but other clients may of course race with this one to set the next state.

// First invoke TLCB to (try to) record the desired next—state value,
// and register the definite winning value and a tentative receive—set.
type testStore struct {
  kv map[Step]Val
  mut sync.Mutex
}

// WriteRead implements the Store interface with a simple intra−process map.
func (ts *testStore) WriteRead(s Step, v Val) Val {
  ts.mut.Lock()
  if v, ok := ts.kv[s]; !ok {
    // no client wrote a value yet for s?
    ts.kv[s] = v // write−once
  }
  v = ts.kv[s] // Read the winning value in any case
  ts.mut.Unlock()
  return v
}

// Object to record the common total order and verify it for consistency

// A test client with particular configuration parameters

// Run a consensus test case with the specified parameters.