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Implementing the HotStuff consensus algorithm

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Declaration of Originality

I, Marc Harvey-Hill of Gonville and Caius College, being a candidate for Part II of the Computer Science Tripos, hereby declare that this dissertation and the work described in it are my own work, unaided except as may be specified below, and that the dissertation does not contain material that has already been used to any substantial extent for a comparable purpose. I am content for my dissertation to be made available to the students and staff of the University.

Signed [signature]

Date [date]

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Introduction

The power of blockchains lies in their ability to decentralise applications that were traditionally run in a centralised manner. The implications of this are far reaching: central banks can be replaced by decentralised cryptocurrencies, traditional corporations can be replaced with DAOs that have decentralised non-hierarchical governance, internet infrastructure like servers and DNS servers can be decentralised, and any possible algorithm can be run on a decentralised 'world computer'. The innovative algorithm underlying blockchains is a solution to the byzantine consensus problem, which allows a group of participants to agree on some shared history (such as a transaction ledger), even while some malicious participants try to undermine the process.

Blockchains can be either permissioned, or permissionless. Permissioned blockchains have a previously agreed set of participants in the consensus algorithm, whereas permissionless blockchains allow participants to join and leave freely. Most well known blockchains such as Bitcoin and Ethereum are of the permissionless variety. Permissionless blockchains can be seen as permissioned blockchains with an additional layer of security which can be proof of work, proof of stake, or some other similar mechanism. These aim to prevent a 'Sybil attack' where a permissioned blockchain can be overrun by a large number of malicous nodes; proof of work, for example, adds a requirement for a proof of computational work in order to participate in consensus, making Sybil attacks economically and computationally infeasible. Permissioned blockchains are of interest for blockchain applications within a group or organisation, such as a company, where the participating machines are known in advance; but they can also be used in a permissionless context with the addition of a proof of work / stake mechanism.

HotStuff is a byzantine consensus algorithm that was notably used by Meta's Libra project, a cancelled permissioned blockchain-based payments system. The algorithm is relevant because of various performance advantages over comparable algorithms like PBFT, DLS, Tendermint, and Casper.

Building practical, well-performing implementations of consensus algorithms is highly non-trivial. These algorithms are usually specified in short pieces of pseudocode that may not be specified precisely, and require much more code to implement in practice. Such software has a wide range of failure modes mostly due to their parallel nature, including deadlocks, resource starvation, and bugs in implementation. [2]

The main contributions of this dissertation are:

- Providing a reference implementation of HotStuff in OCaml based on a paper by Yin et. al [3].
- Outlining key practical challenges and considerations of implementation.
- Adapting the pacemaker mechanism presented, giving a full specification that works in asynchronous environments without synchronised clocks

Preparation

2.1 Starting point

I had some experience of using OCaml from the IA course, but had never used it in a project. I also had some background in distributed systems from the IB course, which briefly covered Raft, a non-byzantine consensus algorithm. I had some understanding of byzantine consensus from my own reading into Nakamoto consensus and from developing a wallet application for Ethereum; neither of these were directly useful to implementing HotStuff, but they gave me some wider context of the field.

2.2 HotStuff algorithm

2.2.1 Problem statement

HotStuff is a Byzantine fault-tolerant consensus algorithm. It allows a group of parties to agree on some piece of information under adverse conditions where some messages can be lost and some parties are controlled by a malicious adversary. The main application of HotStuff is in permissioned blockchains. For example, one could create a cryptocurrency by using HotStuff to reach consensus on a log of transactions like “Account X transfers account Y £10". By allowing a large amount of devices to reach consensus on a ledger, one can develop a decentralised payment system. In general blockchains can be applied in situations where there is some central authority (a bank, DNS server, government, etc.) to instead create a distributed and decentralised system that requires less trust from participants.

The protocol can be viewed as a solution to the Byzantine generals problem. In this problem a group of generals must all agree to siege a castle at the same time, as a single army would be defeated on its own. The problem is that the generals can only communicate via messengers that take some time to arrive and can be captured en-route. Additionally up to a third of the generals may be malicious, and try to prevent the other generals from reaching consensus on a time to attack. By following the HotStuff protocol the generals can ensure that they all attack together. In contrast to the generals problem, HotStuff is able to agree multiple values instead of just one, so instead of deciding a single value like “Attack at dawn", HotStuff can agree on a log of multiple values. The key is that once a value is decided and appended to the log it can never be modified or erased, the log can only ever be extended.

These conditions are described by the system model: partially synchronous, Byzantine, with reliable, authenticated, point-to-point delivery. This means that messages sent by one party will always be delivered to another within some bounded amount of time after GST has been reached and a message source cannot be spoofed. The Byzantine assumption means a maximum of n faulty nodes may be controlled by an adversary that is actively trying to prevent us from correctly reaching consensus, where $n = 3f + 1$ and n is the total number of parties.

2.2.2 Non-Byzantine case

We will start by describing an algorithm to solve the simpler problem of reaching consensus with the stronger assumption of a crash-stop model instead of a Byzantine one. Examples of similar algorithms include Raft and multi-shot Paxos. Such an algorithm must ensure that once a value is decided and appended to the log, it cannot be modified or erased. In each view a leader proposes some log which it aims to decide upon with a group of replicas. Our implementation assigns leaders to views using a round robin system.

Each view can be broken into two phases; phase 1 allows the leader to learn of previously decided values, and in phase 2 it decides on a value. Phase 1 is initiated by the leader broadcasting the current view number to the replicas, that respond by sending their longest accepted log (the one with the highest corresponding view number). Once the leader has a quorum of responses, it initiates phase 2: it selects the longest log that has been sent to it and broadcasts it to the replicas. The leader may also extend the log at this point with its own values, or create a new log if it did not receive anything. Finally the replica updates its log to the value sent by the leader, and sends an acknowledgement. Once the leader receives a quorum of acknowledgements it can commit the new log. This algorithm satisfies our requirement that a committed log can only be extended.

We will refer to phase 1 as the *new-view* stage, and phase 2 as the *commit* phase to use the terminology of the HotStuff paper. These are followed by the *decide* phase, when the leader sends a decide message to replicas. Once they receive this message they can consider the log decided, and execute the new commands which have been added to the log.

1. `New-view:`

   (a) leader → replicas: “view = 3"

   (b) replicas → leader: “view = 2, log = [‘hello’, ‘world’]", ...

2. Commit:

   (a) leader → replicas: “log = [‘hello’, ‘world’, ‘!’]"

   (b) replicas → leader: “ack", ...

3. Decide:

   (a) leader → replicas: “decide"

   (b) replicas: execute log

2.2.3 Byzantine case

In order to extend our algorithm to achieve consensus under a Byzantine threat model we must handle three threats that we will deal with in turn. In order to do this we must first introduce the concept of a ‘threshold signature’. Acknowledgements from replicas all contain a signature over the message to prove that they were actually sent by the correct node. The leader can create a threshold signature by combining $n - f$ ack messages’ signatures to prove that they really received a quorum of acknowledgements. A collection of a quorum of votes with a threshold signature is known as a ‘quorum certificate’.

1. Threat: equivocation - a faulty leader broadcasts one value to some replicas and a different value to others. For example in the case of a cryptocurrency, this could result in a malicious actor (controlling account X) carrying out a double spend attack, sending “Account X transfers account Y £10" to some nodes, and “Account X transfers account Z £10" to other nodes, even if Account X only contains £10.

   Solution: Add a new stage *prepare* which happens just before the *commit* phase. In this phase the leader again chooses the longest log it received in the *new-view* phase to send in the *prepare* phase, this may also be extended with the leader’s new values. Once we receive a quorum of acknowledgements, we begin the *commit* phase. The difference is that this time we include a quorum certificate over the quorum of *prepare* acks, which proves that we preproposed the value to at least $n - f$ nodes and received their acks, so we are not proposing one value to some node and another value to others.

2. Threat: A faulty leader sends in the *prepare* phase a log that conflicts with one that has already been committed.

   Solution: Replicas must lock on a value once it is committed and not accept a *prepare* from a leader that contradicts that. They will store the quorum certificate that they receive during the *commit* phase and will only accept a new *prepare* if it extends from the node stored in the certificate.

3. Threat: A faulty replica sends a the leader a fake log in its new-view message that was never actually proposed. Note that this does not break safety as a pre-proposal for the fake log would not be accepted since the protocol is safe. However, this breaks the liveness property of the protocol as a non-faulty leader could be prevented from making progress by faulty replicas sending fake logs.

   Solution: Add to the *new-view* message a qc over *prepare* acks from the original view in which the message was proposed as a proof that it was indeed proposed.

1. New-view:
   (a) leader → replicas: ‘‘view = 3"
   (b) replicas → leader: ‘‘log = [‘hello’, ‘world’], qc = (prepare acks from view 2)", ...
2. Prepare:
   (a) leader → replicas: ‘‘log = [‘hello’, ‘world’, ‘!’]"
   (b) replicas verify the proposal is safe
   (c) replicas → leader: ‘‘log = [‘hello’, ‘world’, ‘!’] ack", ...
3. Commit (lock):
   (a) leader → replicas: ‘‘log = [‘hello’, ‘world’, ‘!’]", qc = (prepare acks from previous stage)"
   (b) replicas ‘lock’ on proposed log
   (c) replicas → leader: ‘‘ack", ...
4. Decide:
   (a) leader → replicas: ‘‘decide, qc = (commit acks from previous stage)"
   (b) replicas: execute log

2.2.4 Optimistic responsiveness

Consider again the *prepare* phase. In this phase the leader selects the quorum certificate from the highest view that it hears about from the *new-view* messages. However, it is possible that there is some honest replica that we do not hear from (perhaps their message was lost) that is locked on a higher view proposal than the one that we choose to propose. When this replica receives our pre-proposal, it will reject it as it is locked on a higher view proposal. This means that we could be prevented from making progress in this view by missing one replica in the *new-view* phase.

This means that our system doesn't have the 'liveness' property, which means that it will make progress under synchronous conditions when a non-faulty leader is elected. It is possible that we repeatedly don't hear from this one honest replica and fail to make progress indefinitely. One solution to this problem is to introduce a timeout Δ that we must wait for before progressing to ensure that we have allowed sufficient time to pass such that we have heard from the honest replica. This solution has the disadvantage that our system is not 'responsive', which means that the system can make progress as fast as network conditions allow when we have a non-faulty leader, and does not depend on Δ.

In order to achieve responsiveness we can modify our algorithm by adding a *pre-commit* phase in between *prepare* and *commit*. This ensures that if some honest replica becomes locked on a value in the *commit* phase, then there are at least $f + 1$ honest nodes that have a 'key' for that value from the *pre-commit* phase. More specifically, they have a 'key-proof' composed of a quorum certificate over *pre-commit* acks. Replicas then send this key-proof with their *new-view* message when a new view begins. The new leader selects the key-proof with the highest view proposal, and sends the key-proof along with their *prepare* message. Even if the leader does not receive a *new-view* message from the replica which is locked on the highest view proposal, they must have received the key to this proposal from some honest replica, so their *prepare* message will make progress.

1. `New view:`
   - (a) `leader → replicas: “view = 3"`
   - (b) `replicas → leader: “qc = (key-proof for view = 2, log = [‘hello’, ‘world’])", ...`
2. `Prepare:`
   - (a) `leader → replicas: “log = [‘hello’, ‘world’, ‘!’], qc = (key-proof for view = 2, log = [‘hello’, ‘world’])"`
   - (b) `replicas → leader: “log = [‘hello’, ‘world’, ‘!’] ack", ...`
3. `Pre-commit (key):`
   - (a) `leader → replicas: “qc = (prepare acks / key-proof from previous stage)"`
   - (b) `replicas store qc as a ‘key’`
   - (c) `replicas → leader: “log = [‘hello’, ‘world’, ‘!’] ack", ...`
4. `Commit (lock):`
   - (a) `leader → replicas: “qc = (pre-commit acks from previous stage)"`

```
   (b) replicas ‘lock’ on proposed log
   (c) replicas → leader: “ack", ...
5. Decide:
   (a) leader → replicas: “decide, qc = (commit acks from previous
       stage)"
   (b) replicas: execute log
```

2.2.5 View changes

If a leader fails to make progress within some timeout a view change takes place and the next view begins. The HotStuff paper does not go into detail on how view changes take place, so this explanation is based on a talk by one of its authors, Ittai Abraham [1].

Once the view times out, nodes send a *complain* message to the next leader and start a new timeout for the next view. Once the next leader achieves a quorum of *complain* messages it collects them into a QC known as a *view-change proof.* This leader can then send a *view-change* message containing the *view-change proof* to all replicas, who will respond by transitioning to the next view and sending a *new-view* message to the new leader. The inclusion of the *view-change proof* prevents liveness attacks by byzantine nodes that could otherwise attack the system by constantly causing view changes to take place, and preventing non-faulty leaders from making progress.

The use of timeouts in this way is a type of failure detector, which is a system that facilitates the detection of failed nodes. [add more information on failure detectors ***]

2.2.6 Chaining

The unchained algorithm presented goes through three very similar phases in order to commit a proposal, these phases involve collecting votes from replicas to form a QC that then serves in later phases. Instead of having different phases as before, we can have a single *generic* phase that collects votes, creates a *generic QC*, and sends it to the next leader; now we change view on each phase and a QC can serve in multiple phases concurrently.

In each view some leader sends a proposal to all replicas, who send their replies to the next leader who can form a QC to add to their proposal. The replicas also send *new-view* messages to the next leader as before, to allow them to select a node to propose.

*diagram of n-chain ****

The above diagram shows a chain of nodes connected by 'parent' links; every time we propose a value we extend the chain with a new node. Some node *b* also contains a link to another node within *b.justify.node*, this is the previous generic QC in the chain. In the event of a view change a dummy node will be inserted in the chain which will cause a 'gap' where some *b.justify.node* pointer jumps over a dummy node. If some node *b* has a QC that points to its direct parent with no 'gap' in-between, then we say that it forms a *one-chain*. In general if we have n such direct links without 'gaps' we refer to this an *n-chain* as shown above.

To reproduce the behaviour of the unchained algorithm, on receiving a proposal for node b^* a replica must look at b^**.justify.node* to see if it points to an *n-chain*. It must take different actions depending on the length of the chain:

- one-chain: this is equivalent to the *pre-commit* phase from the unchained version, so the replica should store a 'key'
- two-chain: this is equivalent to the *commit* phase, so the replica should 'lock' on the value proposed at the start of the two-chain
- three-chain: this is equivalent to the *decide* phase, so the replica can execute the commands from the node at the start of the three-chain

2.3 Tools & Libraries

2.3.1 OCaml

I chose OCaml for this project due to its high-level nature, static type system, ability to blend functional and imperative paradigms, and good library support. OCaml's multi-paradigm nature is suitable for implementing HotStuff, as the core state machine can be elegantly expressed in a functional way, whereas interacting with the RPC library to send messages is better suited to an imperative paradigm. Additionally the Tezos cryptocurrency is written in OCaml, and contains a cryptography library that provides the functionality needed by HotStuff.

The performance bottlenecks for distributed byzantine algorithms are generally cryptography, message serialisation and network delays. This means that it is more important to chose a language with suitable features to aid implementation, rather than picking a 'high-performance' language like C++.

OCaml has a powerful module system that facilitates writing highly reusable code. The module system was only briefly touched upon in the tripos (in Concepts in programming languages from IB), so I spent time learning about these features. Modules provide an elegant interface for the core state machine to interact with the imperative parts of the program that actually send messages over the network.

There is no existing reference implementation of HotStuff in OCaml, so my project contributes to the growing OCaml ecosystem. This ecosystem is home to an active community, and interesting projects such as MirageOS unikernels. Because my project is implemented purely in OCaml, it could potentially be deployed on a MirageOS unikernel [???].

2.3.2 Lwt

Lwt is a concurrent programming library for OCaml. It allows the creation of promises, which are values that will become determined in the future; these promises may spawn threads that perform computation and I/O in parallel. In order to use Lwt I had to learn about monads, which are ways of sequencing effects in functional languages and are used by asynchronous promises in Lwt. Lwt is useful to this project as promises provide a way to dispatch messages over the network and wait for their responses in different threads. Promises are cheap to create in Lwt, so one can create many lightweight threads with good performance [citation needed***].

2.3.3 Cap'n Proto

Cap'n Proto is an RPC framework that includes a library for sending and receiving RPCs, and a schema language for designing the format of RPCs that can be sent. Benchmarks for the library are presented in 4.2.1.

2.3.4 Tezos cryptography

The Tezos cryptography library provides aggregate signatures using the BLS12-381 elliptic curve construction. It provides functions to sign some data using a private key, to aggregate several signatures into a single one, and to check an aggregate signature is valid. The only difference from the threshold signatures needed by HotStuff is that each individual signature in an aggregate signature can sign different data, whereas with threshold signatures each individual signature is over the same data. It is trivial to implement threshold signatures using this library by checking that the data is the same for all signatures inside the aggregate signature.Benchmarks for the library are presented in 4.2.2.

2.4 Requirements analysis

- Correctness - The consensus algorithm is implemented as it is described in the paper. This can be established by testing of the program trace for compliance with the algorithm specification [do I need to mention change from proposal here ***].
- Evaluation - Analysis of system throughput and latency carried out on a simulated network of 8 replicas. Evaluation will be carried out by testing the program locally, analysing the trace, and testing in an emulator.
- Improve transaction throughput and reduce latency. This can be achieved through architectural decisions, tuning the scheduler, and ensuring cryptographic libraries are being used efficiently.

These requirements are similar to those presented in my proposal (Appendix X***) with a few differences. The first difference is to evaluate on 8 nodes rather than 32. Benchmarking of Cap'n Proto has revealed its limitations when sending large messages. Once batching of requests is implemented the internal messages sent between nodes will be large and could cause a performance bottleneck for the state machine progressing. Because of this it may not be feasible to get reasonable performance with more nodes, as more nodes result in more internal messages being sent. In practice permissioned blockchains are often run with a small number of nodes, so this limitation may not be important [citation needed *** honeybadger BFT?].

2.5 Software engineering practices

2.5.1 Development methodology

[should i say i used something like RAD and productivity tracking tools???]

2.5.2 Testing & debugging methodology

Unit testing was carried out using 'expect tests', which compare the outputted program trace to the correct output. I wrote a testing suite that verifies that the program behaves as specified in the HotStuff paper.

The Memtrace library and viewer were used to profile the memory usage of the program. One can generate a flame graph of memory allocations to see which parts of the program are using the most memory.

Due to the distributed nature of the program, normal debugging tools and profilers are not useful to debugging deadlocks and performance issues. This is because the cause of deadlocks and performance issues is often some process waiting or a backlog of work forming, but this cannot be detected by tools that just track things like CPU usage. Instead I had to rely on manual inspection of the program trace and commands that measure the real time taken for some part of the program to run.

[implement CI???]

2.5.3 Source code management

I used Git for version control, and regularly pushed my local changes to a private GitHub repository.

Implementation

3.1 Overview

The core implementation of the HotStuff algorithm (which we will give in 3.2) is implemented in the *consensus* module. The main function provided by this module is *advance*, which delivers some event (such as an incoming message, client request or timeout) to the consensus state machine and returns an updated state and a list of actions (such as sending a message to another node or responding to a client request) to be carried out. This architecture is inspired by the OCons project[1], which was developed by my project supervisor. The *consensus* module contains both a chained and unchained implementation that share a common signature, so can be interchanged. The module uses the Tezos cryptography library (see 2.3.4) for signing messages, aggregating signatures, and checking quorum certificates.

Each node operates as a server waiting for messages from other nodes or requests from a client (or in the case of our experiments a load generator, which is described in 3.4). The format of RPCs is specified in a Cap'n Proto schema, in their custom markdown language. A received RPC must be decoded, and the Cap'n Proto types converted into the internal types of the consensus state machine. Messages and requests are added to separate streams[2] when they are received. When a request is received the callback function to respond to the request is stored in a hash table, so that it can be accessed when the command has been committed and the client request can be responded to.

The main loop takes events from the message and request streams, prioritising internal messages over client requests [expand on why, maybe cite something ***]. It takes these events and delivers them to the *advance* function of the consensus state machine. This architecture was chosen so that the *advance* function is never run in parallel on different messages / requests, as this could lead to race conditions. The *advance* function then returns a new state which is stored, and a list of actions to carry out.

[1] At the time I began implementation the OCons project was still under development, so I was unable to use the code in my project.

[2] A stream is thread-safe implementation of a queue in Lwt.

The actions that can be carried out are sending a message to other nodes, responding to a client request, and resetting a timer. In order to send a message we must convert the consensus state machine internal types into Cap'n Proto types, and construct an RPC that matches the schema. Messages are dispatched asynchronously in a new thread. The node maintains TCP connections with all other nodes that are reused every time a message is sent, and in the event of the connection breaking the node repeatedly attempts to reconnect with binary exponential back-off. When responding to a client request the committed command's unique identifier is used to lookup the callback to respond to the client, which is then called, sending a response to the client. The timer is implemented with Lwt promises, a new thread is created which waits for a timeout to elapse and then adds a TIMEOUT message to the message stream.

3.2 Pacemaker Specification

We present the pseudocode for the unchained algorithm given in the HotStuff paper (see algorithm 4 & 5) with our additions coloured in green and modifications in pink. We have only shown the main changes and not the other features and performance improvements we have made (such as batching), we will present these in 3.3. In order to better map to the original pseudocode we break the algorithm into HotStuff and the pacemaker, although in our modified algorithm these two sections are not cleanly separable.

- CREATELEAF (Algorithm 1): This function has been modified so that 'dummy' nodes are inserted to maintain the invariant that the height of the chain is always one greater than the current view number. This is very similar to the CREATELEAF function given for the chained version of the protocol in the HotStuff paper; it is unclear as to why this was changed for the unchained version [work out why it changed?].

- ONRECEIVEPROPOSAL (Algorithm 1): The change on line 22 ensures that we only respond to a proposal from our current view, this is important for safety but was not explicitly included in the original pseudocode. The change on line 27 is that we send a NEWVIEW message to the next leader once we receive a proposal. This is in contrast to the original pseudocode where we send a NEWVIEW inside ONNEXTSYNCVIEW on receiving some unspecified interrupt. Finally, once we have received a proposal we can transition into the next view *unless* we are the next leader, in which case we must wait to collect the VOTEMSGs before transitioning.

- ONRECEIVEVOTE (Algorithm 1): The change on line 31 ensures we ignore messages from earlier views, which is important for liveness, and that we are the correct destination for the vote message. Another change we make is dividing V into different sets for messages from different views; this prevents votes

Algorithm 1 Modified HotStuff

```
function CREATELEAF($parent$, $cmd$, $qc$)
    $b.parent \leftarrow$ branch extending with dummy nodes from $parent$ to height $curView$
    $b.height \leftarrow curView + 1$
    $b.cmd \leftarrow cmd$
    $b.justify \leftarrow qc$
    return b
procedure UPDATE($b^*$)
    $b" \leftarrow b^*.justify.node$
    $b' \leftarrow b".justify.node$
    $b \leftarrow b^*.justify.node$
    UPDATEQCHIGH($b^*.justify$)
    if $b'.height > b_{lock}.height$ then
        $b_{lock} \leftarrow b'$
    if $(b".parent = b') \wedge (b'.parent = b)$ then
        ONCOMMIT($b$)
        $b_{exec} \leftarrow b$
procedure ONCOMMIT($b$)
    if $b_{exec}.height < b.height$ then
        ONCOMMIT($b.parent$)
        EXECUTE($b.cmd$)
procedure ONRECEIVEPROPOSAL($\text{MSG}_v(\text{GENERIC}, b_{new}, \bot)$)
    if $m.view = curView$ then
        if $b_{new}.height > vheight \wedge (b_{new}$ extends $b_{lock} \vee n.height > b_{lock}.height)$ then
            $vheight \leftarrow b_{new}.height$
            SEND(GETLEADER(), $\text{VOTEMSG}_u(\text{GENERIC}, b_{new}, \bot)$)
        UPDATE($b_{new}$)
        SEND(GETNEXTLEADER(), $\text{MSG}_u(\text{NEWVIEW}, \bot, qc_{high})$)
        if not ISNEXTLEADER() then
            ONNEXTSYNCVIEW($curview + 1$)
procedure ONRECEIVEVOTE($\text{VOTEMSG}_v(\text{GENERICACK}, b, \bot)$)
    if ISLEADER($m.view + 1$) $\wedge\ m.view \geq curView$ then
        if $\exists (v, \sigma') \in V_{\text{m.view}}[b]$ then
            return
        $V[b] \leftarrow V_{\text{m.view}}[b] \cup \{(v, m.partialSig)\}$
        if $|V_{\text{m.view}}[b]| \geq n - f$ then
            $qc \leftarrow QC(\{\sigma | (v', \sigma) \in V_{\text{m.view}}[b]\})$
            UPDATEQCHIGH($qc$)
            ONNEXTSYNCVIEW($m.view + 1$)
function ONPROPOSE($b_{leaf}$, $cmd$, $qc_{high}$)
    $b_{\text{new}} \leftarrow$ CREATELEAF($b_{leaf}$, $cmd$, $qc_{high}$, $b_{leaf}.height + 1$)
    BROADCAST($\text{MSG}_v(\text{GENERIC}, b_{new}, \bot)$)
    return $\text{b}_{\text{new}}$
```

Algorithm 2 Modified Pacemaker

```
function GETLEADER
    return $curView$ mod $nodeCount$
procedure UPDATEQCHIGH($qc'_{high}$)
    if $qc'_{high}.node.height > qc_{high}$ then
        $qc'_{high} \leftarrow qc_{high}$
        $b_{leaf} \leftarrow qc'_{high}.node$
procedure ONBEAT($cmd$)
    if $u =$ GETLEADER() then
        $b_{leaf} \leftarrow$ ONPROPOSE($b_{leaf}, cmd, qc_{high}$)
procedure ONNEXTSYNCVIEW($view$)
    $curView \leftarrow view$
    ONBEAT($cmds.take()$)
    RESETTIMER($curView$)
procedure ONRECEIVENEWVIEW(MSG$_u$(NEWVIEW, $\perp$, $qc'_{high}$))
    UPDATEQCHIGH($qc'_{high}$)
procedure ONRECIEVECLIENTREQUEST(REQ($cmd$))
    $cmds.add(cmd)$
procedure ONTIMEOUT($view$)
    SEND(GETNEXTLEADER(), MSG(COMPLAIN, $\perp$, $\perp$))
    RESETTIMER($view + 1$)
procedure ONRECIEVECOMPLAIN($m =$ MSG(COMPLAIN, $\perp$, $\perp$))
    if ISLEADER($m.view + 1$) $\wedge$ $m.view \geq curView$ then
        if $\exists(v, \sigma') \in C_{\text{m.view}}[b]$ then
            return
        $C_{\text{m.view}}[b] \leftarrow C[b] \cup \{(v, m.partialSig)\}$
        if $|C_{\text{m.view}}[b]| = n - f$ then
            $qc \leftarrow QC(\{\sigma | (v', \sigma) \in C_{\text{m.view}}[b]\})$
            BROADCAST(MSG(NEXTVIEW, $\perp$, $qc$))
procedure ONRECEIVENEXTVIEW($m =$ MSG($*$, $\perp$, $qc$))
    if $qc.view \geq curView$ then
        ONNEXTSYNCVIEW($qc.view + 1$)
```

from different views being used to form a quorum. The change on line 38 means that once a leader has received a quorum of messages, it can transition to the next view (which it starts by sending a proposal in ONNEXTSYNCVIEW)

- GETLEADER (Algorithm 2): The original pseudocode states that this function is application specific. We have chosen to use a round-robin system to assign leaders to views.

- ONNEXTSYNCVIEW (Algorithm 2): This function previously just sent a NEWVIEW message to the next leader. Our modified function updates the *curView*, and resets the timer for the new view. Additionally it calls ONBEAT which causes a leader to propose a new value.

- ONRECIEVECLIENTREQUEST (Algorithm 2): On receiving a client request we simply add it to a queue of commands waiting to be proposed.
- ONTIMEOUT (Algorithm 2): When our current view times out we send a COMPLAIN to the next leader, and reset our timer for the next view. This behaviour is explained in 2.2.5.
- ONRECIEVECOMPLAIN (Algorithm 2): This function is very similar to ONRECEIVEVOTE, except that we are collecting a quorum of COMPLAIN messages rather than votes. On achieving a quorum we can broadcast a NEXTVIEW message to get the replicas to transition to the next state, including the quorum of COMPLAINs as proof. This behaviour is explained in 2.2.5.
- ONRECEIVENEXTVIEW (Algorithm 2): N.B. the message type given is the wildcard operator, so this procedure is run on every message we receive. The function checks if the quorum received is from a greater view than *curView*, if so then it is safe to transition to that view in order to catch up.

3.2.1 Correctness Proof

3.2.2 Liveness Proof

3.3 Performance improvements

3.3.1 Batching

One way to improve goodput (number of requests committed) is to 'batch' requests, meaning a node may contain many commands instead of just one. This can dramatically increase goodput as now a single view can result in many commands being committed and executed instead of just one.

In order to implement this change in algorithm 2 one simply has to modify the ONNEXTSYNCVIEW procedure. Instead of taking a single element from the queue of commands waiting to be proposed, the whole queue will be 'batched' into a single proposal.

In theory this change should result in much higher goodput without any significant increase in latency. Analysis of timing data after implementing this feature revealed that latency had increased substantially. We now present some of the analysis and experimentation that was carried out to diagnose this issue. As mentioned in 2.5.2, the nature of the project meant that debugging had to be carried out by manual inspection of the program trace and timing sections of the program. To overcome this I carried out tests in a scientific manner, constructing a hypothesis for why the program was slow based on crawling through logs, then attempting to test my hypothesis while controlling other variables, and finally implementing a solution.

Probing effects

I added timing and logging statements to key parts of the program, allowing me to better diagnose the source of the poor performance. Running a live test with print statements has the advantage that one can quickly see when progress is not being made, or when there are pauses, as the print statements stop. However, this benefit is outweighed by the sheer volume of logs and times being printed, it becomes difficult to manage when logs are in the millions of lines long.

Moreover, debugging by print statements had a bigger problem of inconsistencies between runs making it difficult to diagnose any problem. This was because the large volume of print statements being executed had a significant effect on the performance of the program. This is known as a *probing effect*, the behaviour of a system is altered by the act of measuring it. I initially assumed that because the print statements were not in-between the timing statements they would not affect my measurements, but they are an expensive operation that can cause delays to happen in execution where one would not expect.

This problem was overcome through the development of a simple logging framework. Key parts of the program such as the state machine advancing, actions being carried out, messages being sent, are all timed and added to lists. Critically, this list is stored and not printed out until the node is killed, so the printing cannot interfere with execution. Relevant statistics such as mean, standard deviation, and total time for each interval measured are outputted, providing crucial insight into the performance of the program.

Design principle: Minimise probing effects by carrying out the minimum possible amount of work in critical areas of the program, storing data, and moving work (such as outputting statistics) to less critical areas of the program.

Architectural experimentation

Initial analysis of timing data showed a large amount of time was spent by requests queuing on the node before they are delivered to the state machine. At this time the architecture presented in 3.1 had not been fully developed, specifically, there was only one stream that contained both internal messages and client requests.

Hypothesis: Internal messages are being starved by client requests. At large throughputs the stream quickly becomes filled with client requests, which may prevent internal messages being handled. Internal messages represent a backlog of work that we have not yet finished, so this should be handled before accepting more requests.

Design principle: The system should prioritise clearing its backlog of work before accepting new work.

Potential solution: Split the stream into two separate streams for messages and requests, and always pick from the message stream over the request stream until the message stream is empty.

Implementing the potential solution did not lead to a significant improvement in performance, so this was not the issue. However, this architectural model was retained as it is theoretically better and may lead to observable improvements later. Another potential cause of the issue was that internal messages are starving client requests, rather than the other way around. The modified implementation of the pacemaker (3.2) immediately begins a new view as soon as the previous one is finished, which causes a large amount of internal messages.

Hypothesis: Because the state machine is constantly advancing at the fastest possible rate, the volume of internal messages may prevent new requests from being handled.

Experimentation: One approach is to 'balance' the starvation by mostly prioritising internal messages, but every x iterations picking a request instead of an internal message. Varying x led to significant changes in performance, but no value for x led to a good level of performance. This approach did not appear workable, as the nature of starvation seemed to be a complicated, with messages starving requests, and the other way around.

Potential solution: Instead of starting the next view whenever possible, deliver the $BEAT$ and $ONNEXTSYNCVIEW$ interrupts at a steady rate.

This potential solution was implemented; it involved checking a timer on every iteration of the main loop to see if some Δ had elapsed and it was time to deliver the next interrupt, as well as significant modifications to be made to the consensus state machine. Implementing this feature actually led to a significant performance hit, so the changes were discarded. However, experimenting with this feature showed an interesting phenomenon; the main loop, which is infinitely recursive, was running at a much slower rate than one would expect. Further timing statements revealed that the command which removed an element from the stream would sometimes take a very long time (on the order of seconds). Reading more into the Lwt documentation led to the following idea:

Potential solution: Use a different method to take elements from the queue that is synchronous. This should stop the main loop blocking waiting for the queue to fill up.

This change improved performance by reducing the amount of time spent waiting in the main loop, the difference is shown in an ablation study [reference evaluation ***]. However, performance was still not at a satisfactory level.

Send-to-all and deduplication

Further problems were uncovered by the implementation of send-to-all in the load generator (see 3.4). Without this feature a client would randomly chose a node to send their request to, and the command would not be proposed until the round-robin system reached that node to become the leader. Send-to-all can reduce latency by instead broadcasting a request to all nodes, meaning that the next leader can propose the command without having to wait for the round-robin system to reach them.

Instead of reducing latency as expected the implementation of this feature actually led to dramatically increased latency, and some requests not being fulfilled at all within the time that the experiment ran. Heatmaps that show the distribution of latencies over the course of an experiment showed that latency increased exponentially over the course of the test.

Hypothesis: Send-to-all overloads the system by increasing the number of requests by a factor of n, where n is the number of nodes.

Experimentation: If our hypothesis is true, then we should see similar behaviour in a send-to-one system run at n times the throughput. This is indeed the case; running a send-to-one system with such high throughput led to exponential growth in latencies.

Further evidence: Implementing this feature resulted in Lwt errors being thrown. As described in 3.1, in order to respond to client requests, the node maintains a hash table of promises that can be awakened when the consensus state machine has successfully committed and can respond to the client. The Lwt error was caused by promises stored in the hash table being awoken multiple times, which causes an error. This provides further evidence that the same commands are being committed and responded to multiple times, which is wasted work.

Potential solution: Deduplicate commands in the system to minimise the number of commands that are proposed by multiple nodes.

Design principle: Attempt to minimise the amount of redundant work that the system carries out by screening incoming work to check that it actually needs to be done.

One way to implement this is to maintain a set of commands that have been committed, and before proposing or committing a node, calculating the set difference with the *commited* set to filter out redundant commands. Instead of nodes containing a list of commands, as in our naive implementation of batching, we now use a set to enable efficient computation of the difference.

Algorithm 3 Deduplication implementation #1

```
 1: procedure OnCommit($b$)
 2:     if $b_{exec}.height < b.height$ then
 3:         $toCommit \leftarrow b.cmds \setminus commited$
 4:         $commited \leftarrow commited \cup toCommit$
 5:         OnCommit($b.parent$)
 6:         EXECUTE($toCommit$)
 7: procedure OnNextSyncView($view$)
 8:     $curView \leftarrow view$
 9:     $toPropose \leftarrow cmds \setminus commited$
10:     $cmds = \{\}$
11:     OnBeat($toPropose$)
12:     ResetTimer($curView$)
13: procedure OnRecieveClientRequest(REQ($cmd$))
14:     $cmds \leftarrow cmds \cup \{cmd\}$
```

This implementation did not lead to the desired reduction in latency, the system still exhibited the exponential growth previously observed. The lack of change in observable behaviour implied that the filtering of requests was not happening successfully.

Hypothesis: A command (x) is committed four views after it is proposed. In these four views our deduplication is ineffective in screening x from being proposed again. Notably another node will not try to commit x for a second time, but the data indicates that being proposed multiple times is a more important factor affecting performance.

Experimentation: Count the elements that are filtered out from being proposed by measuring the value $|cmds| - |toPropose|$. This value was small, which supports the hypothesis that filtering was largely ineffective.

Potential solution: Screen incoming commands more aggressively by instead maintaining a set of any commands that we have *seen*. Only filter out commands when they are being proposed rather than when thet are being committed, as there is no evidence that screening at commit time gives any tangible benefit.

Note the small optimisation on line 8, we can safely empty the *seen* set as we will not try to propose this value again; this reduces the amount of computation

Algorithm 4 Deduplication implementation #2

```
procedure OnReceiveProposal($\text{MSG}_v(\text{GENERIC}, b_{new}, \bot)$)
    if $m.view = curView$ then
        $seen \leftarrow seen \cup b_{new}.cmds$
        // ... the rest is as before, see algorithm 1
procedure OnNextSyncView($view$)
    $curView \leftarrow view$
    $toPropose \leftarrow cmds \setminus seen$
    $cmds = seen = \{\}$
    OnBeat($toPropose$)
    ResetTimer($curView$)
procedure OnRecieveClientRequest(REQ($cmd$))
    $cmds \leftarrow cmds \cup \{cmd\}$
```

required to calculate the set difference next time. The second implementation was much more effective at screening requests and resulted in a significant reduction in latency. The effectiveness of this change led to further insight into what the bottlenecks for performance were.

Batch sizes

The effectiveness of filtering out commands to make proposals smaller implies that the size of a proposal has a significant impact on the performance of the program. Furthermore, analysis of the time data for sending messages showed that there was a high variance in time taken, with some messages taking on the order of milliseconds [verify ***] to be sent.

Hypothesis: Large batches of commands cause messages to become larger, which causes them to be sent slowly due to limitations in the RPC framework. **Potential mitigation:** Limit the size of batches to prevent messages becoming too large.

Implementing this feature required minimal changes to algorithm 4, one simply has to take a subset of *cmds* to propose instead of the whole set. This feature gave a significant performance improvement, and largely eliminated the exponential growth in latency that had been seen previously. There is a trade-off between sending larger batches (more commands are committed, but messages take longer to send) and sending smaller batches (less commands are committed, but messages are sent faster). This trade-off is explored more in our analysis of the system with different batch sizes [refer to evaluation***]. Fundamentally, there are limitations in the RPC framework that give an upper bound on the performance we can hope to achieve, these limitations are evident in our benchmarking of the Cap'n Proto framework (see 4.2.1).

3.3.2 Encoding of nodes

This section concerns the difficulties of encoding nodes in the Cap'n Proto schema, and designing a suitable type to represent them. As mentioned in 3.1, Cap'n Proto requires a schema to define the format of RPCs that can be sent, and one must convert between consensus state machine internal types and Cap'n Proto types in order to send messages.

Recursive types

The unchained HotStuff algorithm has a genesis node b_0 that starts the chain of nodes. Each node contains the fields *parent*, *cmd*, *justify*, and *height*, with the *justify* field containing a quorum certificate with a 'pointer' to another node in the chain (see 2.2.6). The genesis node b_0 contains a hardcoded link to itself, so $b_0.justify.node = b_0$.

This recursion poses a problem when carrying out the conversion between Cap'n Proto types and OCaml types. It is perfectly possible to define a recursive type in OCaml, so one can represent b_0 inside the consensus state machine. However, the naive implementation of a function to convert this node into a Cap'n Proto type will not terminate, as it will infinitely recurse into the field b_0*.justify.node*. A simple solution to this problem is to add a flag to the Cap'n Proto schema *is_ b_0*; when this flag is enabled then the node is assumed to be equal to b_0. This prevents b_0 from ever having to be converted into a Cap'n Proto type or being sent over the network, it can instead be reconstructed as a recursive type in the consensus state machine of the receiver.

Node offsets

During live testing, memory usage increased rapidly, the nodes quickly exceeded the amount of memory the system had, and their processes were killed by the OS. Profiling with Memtrace (see 2.5.2) showed that the functions for converting between Cap'n Proto types and internal consensus types was responsible for the rapidly increasing memory usage.

This problem arose from the internal types that were designed to represent nodes. As well as the *node* type, there was a *qc* type that was used as the type of the *node.justify* field since it is a quorum certificate. By defining types in this way the 'pointers' in *node.justify.node* were not actually pointers, they contained a significant part of the chain. In this way the chain contained many redundant copies of parts of it, resulting in a very bloated *node* object that was very expensive to convert.

A solution to this problem is to store an offset to a node inside the *justify* field rather than the node itself - making it more like a real pointer. This offset represents how many *parent* links away the node is, and so can be used to reconstruct all of the original information. To implement this another type *node_justify* was added, that is identical to *qc*, but with the field *node* replaced with *node_ offset.* One must then convert between the *node_justify* and *qc* types to reconstruct the original data and follow the *node.justify.node* link.

Node comparison

Memtrace profiling also showed that the equality function for nodes was memory intensive. The solution to this was including a *digest* field in the node that is a hash over all of the other fields. Notably we can compute this hash over the *digest* field of the parent node rather than recursing through the whole chain. This means that we can also compare two nodes by comparing their digests without having to recurse through each chain; the digests being equal cryptographically guarantees that the whole chains are equal.

Node trunction

Implement TCP style nodes??

3.3.3 Connections

3.4 Implementing for evaluation

3.4.1 Load generator

3.4.2 Experiment scripts

3.5 Repository Overview

```
src
    consensus
        consensus_chained_impl.ml
        consensus_chained_test.ml
        consensus_impl.ml
        consensus_test.ml
        consensus.ml
        consensus.mli
        crypto.ml
        types.ml
        types.mli
        util.ml
    experiments
        data
        graphs
        plot.py
        plotDummy.py
    hs_api.capnp
    api_wrapper.ml
    api_wrapper.mli
    dummy.ml
    hs.ml
    live_test.ml
    main.ml
```

The *src* directory contains the main server loop and the code for interacting with Cap'n Proto to send messages. It also contains code for the load generator, and Python scripts to run experiments and benchmarks. Inside the *consensus* folder is the implementation of the consensus library based on the pseudocode presented in 3.2. The *experiments* folder is where the data from running experiments is outputted to, and it contains scripts for plotting graphs.

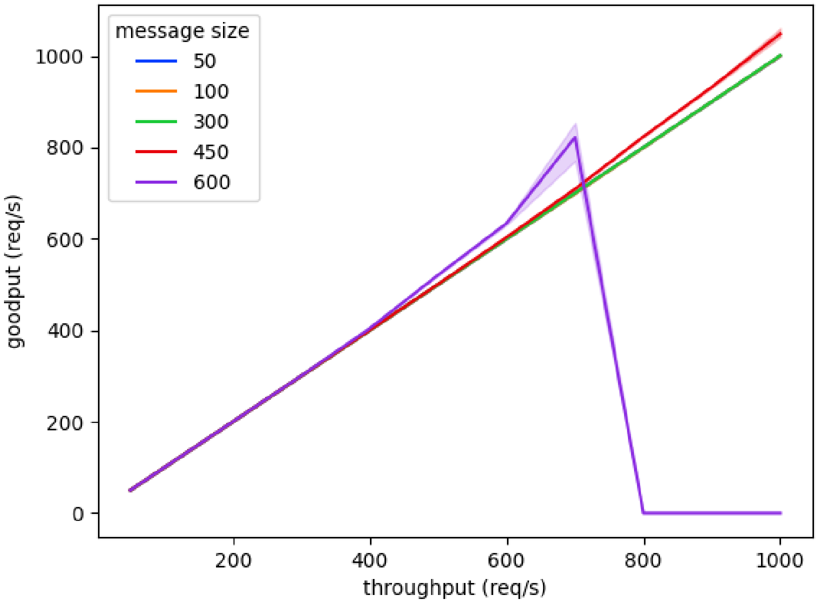
Evaluation

4.1 Setup

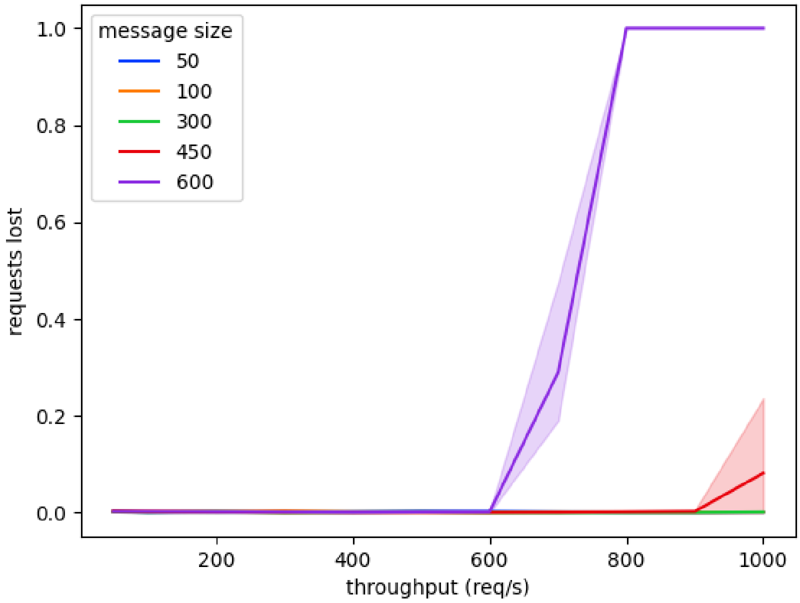
4.2 Library benchmarks

4.2.1 Cap'n Proto

I benchmarked the message sending functionality of Cap'n Proto. I did this with an open-loop load generator (which is described in 3.4), running each experiment for 10s. I varied the size of messages sent in different tests to replicate the behaviour of the algorithm when sending 'batches' of many commands, so a message size of 600 means that the message size is approximately that of a message containing 600 commands.



The figures demonstrate that the framework has a severe drop in performance



when sending large messages. For a message size of 600 the goodput goes to zero as the throughput increases, meaning that no messages are being responded to.

[maybe add goodput / latency graph once it is in a reasonable state! ***]

4.2.2 Tezos Cryptography

I profiled the important functions of the library with Jane Street's Core_bench module. Core_bench is a micro-benchmarking library used to estimate the cost of operations in OCaml, it runs the operation many times and uses linear regression to try to reduce the effect of high variance between runs.

Function	Time (µs)
Sign	427.87
Check	1,171.77
Aggregate (4 sigs)	302.90
Aggregate check (4 sigs)	1,179.25
Aggregate (8 sigs)	605.38
Aggregate check (8 sigs)	1,180.61

Conclusion

5.1 Future Work

Further work would port to using the async library which is known to have better performance. It was out of scope to rewrite in async or use it in the first place due to poor documentation (although now I could look at the type signatures and understand the documentation). Hopefully reimplementing would give better performance and avoid the bugs of capnrpc. I would carry out more extensive tests on the message sending capabilities before diving into implementation. I would be more aware beforehand of the whole algorithm (including the pacemaker code) and implement based on the new pseudocode we have presented and proven correct. This would allow for better structuring of the code.

We have presented a potential path for implementing verifiable anonymous identities and reconfiguration using our permissioned blockchain, future work could consist of a practical implementation of this.

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