

MERCURY: A multilinear Polynomial Commitment Scheme with constant proof size and no prover FFTs

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Abstract

We construct a pairing-based polynomial commitment scheme for multilinear polynomials of size n where constructing an opening proof requires $O(n)$ field operations, and $2n + O(\sqrt{n})$ scalar multiplications. Moreover, the opening proof consists of a constant number of field elements. This is a significant improvement over previous works which would require either

1. $O(n \log n)$ field operations; or
2. $O(\log n)$ size opening proof.

The main technical component is a new method of verifiably folding a witness via univariate polynomial division. As opposed to previous methods, the proof size and prover time remain constant *regardless of the folding factor*.

1 Introduction

1.1 Our results

2 Overview of technique

Remark 2.1. *In this overview, we use some of the notation defined in Sections 3.1 and 3.2.*

Table 1: Comparison of pairing-based ml-PCS. \mathbb{G} denotes a scalar multiplication. All verifiers below additionally require two pairings.

Scheme	Proof size	Prover Work	Verifier Work
univariate-based e.g.[PH23]	$O(1) \mathbb{F}$	$O(n \log n) \mathbb{F}, O(n) \mathbb{G}$	$O(\log n) \mathbb{F}, O(1) \mathbb{G}$
gemini [gem]	$O(\log n) \mathbb{F}$	$O(n) \mathbb{F}, 3n \mathbb{G}$	$O(\log n) \mathbb{F}, O(\log n) \mathbb{G}$
zeromorph [zer]	$O(\log n) \mathbb{F}$	$O(n) \mathbb{F}, 2.5n \mathbb{G}$	$O(\log n) \mathbb{F}, O(\log n) \mathbb{G}$
mercurij (this work)	$O(1) \mathbb{F}$	$O(n) \mathbb{F}, 2n + O(\sqrt{n}) \mathbb{G}$	$O(\log n) \mathbb{F}, O(1) \mathbb{G}$

Our technique is best thought of as an improvement of the ml-PCS from **gemini** [gem]. Let's start by recalling how **gemini** works. **gemini** commits to a multilinear function as a *univariate* KZG commitment[KZG10]. Specifically, fix a vector $f \in \mathbb{F}^n$ describing the function's values on the boolean cube B_s where $s = \log n$. That is, we think of f as representing the multilinear

$$M(X_0, \dots, X_{s-1}) = \sum_{u \in B_s} \text{eq}(u, X_0, \dots, X_{s-1}) f_i.$$

Let $\text{srs} = \{[x^i]_1\}_{i < n}$ be a KZG structured reference string. **gemini** outputs $\text{cm} = \sum_{i < n} f_i [x^i]_1$ as a commitment to M

Now suppose prover **P** wants to convince verifier **V** that $M(u) = v$, for some $u = (u_0, \dots, u_{s-1}) \in \mathbb{F}^s$. In **gemini**, **P** sends commitments $\text{cm}_1, \dots, \text{cm}_s$ to the s incremental restrictions leading to evaluation at u . Namely, to $M_1 = M(u_0, X_1, \dots, X_{s-1})$, $M_2 = M(u_0, u_1, X_2, \dots, X_{s-1})$, \dots , $M_s = M(u_0, \dots, u_{s-1})$. Assuming **P** sent commitments to the correct functions, all that is needed is to check that cm_s is the commitment to the constant v . Of course, the interesting part is proving the commitments *are* to the correct functions!

For this purpose, **gemini** exploits a connection between M and its corresponding univariate $f(X)$: Write $f(X) = f_0(X^2) + X f_1(X^2)$, for $f_0(X), f_1(X)$ of degree $< n/2$. Let $f_{u_1}(X)$ be the univariate corresponding to M_1 defined above. Then, we have

$$f_{u_1}(X) = (1 - u_1) f_0(X) + u_1 f_1(X).$$

Additionally, we can evaluate f_0 and f_1 via f using the equations

$$f_0(X^2) = \frac{f(X) + f(-X)}{2}, f_1(X^2) = \frac{f(X) - f(-X)}{2X}$$

Thus, we can perform consistency checks between each pair $\text{cm}_{i-1}, \text{cm}_i$, via univariate KZG openings at a random challenge, inductively showing cm_i is indeed the commitment to the next desired restriction. Of course, we get $O(s) = O(\log n)$ proof length due to this sequence of restriction commitments.

Here is a first idea on how to reduce proof length. Protocols based on univariate polynomials allow us to do multilinear evaluation in $O(n \log n)$ prover time with constant proof size (e.g. Section 5 of [PH23]). Choose a parameter t and set $b = 2^t$. We can run

only the first t rounds of **gemini**, reaching a restricted multilinear on $n - t$ variables. If $n' = n/b \leq n/\log n$, we can afford to run a univariate protocol with $O(n' \log n') = O(n)$ prover time to evaluate $M_t(u_t, \dots, u_s)$. This still doesn't take us to overall constant proof size - as we need to use a super-constant t to reach such n' . (For us $t = \log n/2$ will be optimal, although $t \geq \log \log n$ suffices here.)

This raises the question - can we “skip” the intermediate **gemini** rounds and send *only* the commitment \mathbf{cm}_t , and directly prove it is consistent with the original \mathbf{cm} ? Extrapolating the **gemini** strategy in the natural way, we get the answer - yes, but not with constant proof size: We can decompose f into b polynomials of degree $< n/b$: $f(X) = \sum_{0 \leq i < b} X^i f_i(X^b)$. As in the $b = 2$ case, one can show the univariate $f'(X)$ corresponding to M_t is a linear combination of the $\{f_i(X)\}$. Moreover, evaluating the f_i using f (for the consistency check) can be done. However, it requires b evaluations of f . Specifically, $f_i(r^b)$ is a linear combination of $\{f(r), f(r\omega), \dots, f(r\omega^{b-1})\}$ where ω is a primitive b 'th root of unity.

Our central innovation is a different way to prove \mathbf{cm}_t is correct *with* constant proof size. The (univariate corresponding to the) correct restricted polynomial is

$$h(X) = \sum_{0 \leq i < b} \mathbf{eq}(i, u_1) f_i(X).$$

Let $g(X) := f(X) \bmod X^b - \alpha$. Calculation shows

$$g(X) = \sum_{0 \leq i < b} X^i f_i(\alpha).$$

The multilinear $\mathbf{ml}(g)$ corresponding to $g(X)$ is

$$\mathbf{ml}(g)(X_0, \dots, X_{t-1}) = \sum_{0 \leq i < b} \mathbf{eq}(i, X_0, \dots, X_{t-1}) f_i(\alpha)$$

In particular, we have

$$\mathbf{ml}(g)(u_1) = \sum_{0 \leq i < b} \mathbf{eq}(i, u_1) f_i(\alpha) = h(\alpha).$$

In words, the evaluation of g at u_1 as a *multilinear* corresponds to the evaluation of h at α as a univariate! We can use standard univariate KZG to open \mathbf{cm}_t at α . And, crucially, we can afford to evaluate $\mathbf{ml}(g)(u_1)$ using the aforementioned univariate protocols as it is of size b rather than n . In summary, we can show a committed polynomial corresponds to the correct restriction. And now, again, we can afford to open h as a multilinear at u_2 using univariate protocols as it has size n/b rather than n .

Comparison to sumcheck It is instructive to see what happens if we try to get a similar result via a modification of the sumcheck protocol [?]. Note first that a multilinear

evaluation can indeed be written as a sum over the function's values on the B_n multiplied by the **eq** function:

$$\hat{f}(u) = \sum_{b \in B_s} \mathbf{eq}(b, u) \hat{f}(b).$$

The classic sumcheck protocol, like **gemi**, works by $\log n$ reductions of the domain size by a factor of two; each round fixing one more variable of the summed function. In the above spirit, we could look at a modified sumcheck protocol, where the first variable ranges over a domain of super-constant size b . The first round univariate P_1 would thus have degree $2b$ in our case. We can send a commitment to it rather than its coefficients as usually done to maintain constant proof size. However, *computing* P_1 would require superlinear time $O(n \log b)$ - as we need to perform a b -size FFT for n/b values of the second variable appearing in the sum.

3 Preliminaries

3.1 Terminology and conventions

We work with integer parameter n that we'll assume throughout the paper is of the form $n = 2^{2t}$ for integer $t > 0$. We'll denote its square root by $b := 2^t = \sqrt{n}$. We index vectors starting from zero. For example, for $g \in \mathbb{F}^b$ we have $g = (g_0, \dots, g_{b-1})$. We associate vectors with univariate polynomials in the following natural way: Given $g \in \mathbb{F}^b$ we denote $g(X) := \sum_{0 \leq i < b} g_i X^i$.

We make the convention that integer ranges in sums begin at zero if not specified otherwise. Thus, we write $g(X) = \sum_{i < b} g_i X^i$.

We assume vectors of size n are indexed by two indices ranging over $\{0, \dots, b-1\}$. Thus, for $f \in \mathbb{F}^n$, we have $f = (f_{0,0}, \dots, f_{0,b-1}, \dots, f_{b-1,0}, \dots, f_{b-1,b-1})$. Accordingly, for $0 \leq i < b$, we denote by f_i the vector $(f_{i,0}, \dots, f_{i,b-1})$.

In particular, for $f \in \mathbb{F}^n$ we have under these notations that

$$f(X) := \sum_{i < b} X^i f_i(X^b) = \sum_{i < b} \sum_{j < b} f_{i,j} X^{i+j \cdot b}$$

We denote by B_t the binary cube $\{0, 1\}^t$ of dimension t .

3.2 Multilinear polynomials

Let $n = 2^{2t}$. We define the well-known **eq** multilinear polynomial in $4t$ variables.

$$\mathbf{eq}(x, y) := \prod_{i=0}^{t-1} (x_i y_i + (1 - x_i)(1 - y_i))$$

We have for $x, y \in B_{2t}$, $\mathbf{eq}(x, y) = 1$ when $x = y$ and $\mathbf{eq}(x, y) = 0$ otherwise.

We use the convention that an integer $0 \leq i < n$ can be used as an input to **eq** by interpreting i as its binary representation. Namely, for $0 \leq i < n$, $u \in \mathbb{F}^t$, $\mathbf{eq}(i, u) := \mathbf{eq}(i_1, \dots, i_t, u)$ where $i = \sum_{j \in [t]} i_j 2^{j-1}$.

For $a \in \mathbb{F}^n$, we define $\text{ml}(f)$ to be the multilinear polynomial obtaining f 's values on the boolean cube. Namely,

$$\text{ml}(a)(X_0, \dots, X_{t-1}) := \sum_{i < n} \mathbf{eq}(i, X_1, \dots, X_t) \cdot a_i.$$

3.3 The algebraic group model

We introduce some terminology from [GWC19] to capture analysis in the Algebraic Group Model of Fuchsbauer, Kiltz and Loss [FKL18].

In our protocols, by an *algebraic adversary* \mathcal{A} in an SRS-based protocol we mean a $\text{poly}(\lambda)$ -time algorithm which satisfies the following.

- For $i \in \{1, 2\}$, whenever \mathcal{A} outputs an element $A \in \mathbb{G}_i$, it also outputs a vector v over \mathbb{F} such that $A = \langle v, \text{srs}_i \rangle$.

First we say our *srs has degree Q* if all elements of srs_i are of the form $[f(x)]_i$ for $f \in \mathbb{F}_{<Q+1}[X]$ and uniform $x \in \mathbb{F}$. In the following discussion let us assume we are executing a protocol with a degree Q SRS, and denote by $f_{i,j}$ the corresponding polynomial for the j 'th element of srs_i .

Denote by a, b the vectors of \mathbb{F} -elements whose encodings in $\mathbb{G}_1, \mathbb{G}_2$ an algebraic adversary \mathcal{A} outputs during a protocol execution; e.g., the j 'th \mathbb{G}_1 element output by \mathcal{A} is $[a_j]_1$.

By a “real pairing check” we mean a check of the form

$$(a \cdot T_1) \cdot (T_2 \cdot b) = 0$$

for some matrices T_1, T_2 over \mathbb{F} . Note that such a check can indeed be done efficiently given the encoded elements and the pairing function $e : \mathbb{G}_1 \times \mathbb{G}_2 \rightarrow \mathbb{G}_t$.

Given such a “real pairing check”, and the adversary \mathcal{A} and protocol execution during which the elements were output, define the corresponding “ideal check” as follows. Since \mathcal{A} is algebraic when he outputs $[a_j]_i$ he also outputs a vector v such that, from linearity, $a_j = \sum v_\ell f_{i,\ell}(x) = R_{i,j}(x)$ for $R_{i,j}(X) := \sum v_\ell f_{i,\ell}(X)$. Denote, for $i \in \{1, 2\}$ the vector of polynomials $R_i = (R_{i,j})_j$. The corresponding ideal check, checks as a polynomial identity whether

$$(R_1 \cdot T_1) \cdot (T_2 \cdot R_2) \equiv 0$$

The following lemma is inspired by [FKL18]'s analysis of [Gro16], and tells us that for soundness analysis against algebraic adversaries it suffices to look at ideal checks. Before stating the lemma we define the Q -DLOG assumption similarly to [FKL18].

Definition 3.1. Fix integer Q . The Q -DLOG assumption for $(\mathbb{G}_1, \mathbb{G}_2)$ states that given

$$[1]_1, [x]_1, \dots, [x^Q]_1, [1]_2, [x]_2, \dots, [x^Q]_2$$

for uniformly chosen $x \in \mathbb{F}$, the probability of an efficient \mathcal{A} outputting x is $\text{negl}(\lambda)$.

Lemma 3.2. *Assume the Q -DLOG for $(\mathbb{G}_1, \mathbb{G}_2)$. Given an algebraic adversary \mathcal{A} participating in a protocol with a degree Q SRS, the probability of any real pairing check passing is larger by at most an additive $\text{negl}(\lambda)$ factor than the probability the corresponding ideal check holds.*

See [GWC19] for the proof.

3.4 Polynomial commitment schemes for multilinear polynomials

Definition 3.3. *Let $n = 2^t$. A multi-linear polynomial commitment scheme (ml-PCS) consists of*

- $\text{gen}(n)$ - a randomized algorithm that outputs an SRS srs .
- $\text{com}(f, \text{srs})$ - that given a polynomial $f \in \mathbb{F}^n$ returns a commitment cm to f .
- A public coin protocol $\text{open}(\text{cm}, n, u, v)$ between parties \mathbf{P} and \mathbf{V} . \mathbf{P} is given $f \in \mathbb{F}^n$. \mathbf{P} and \mathbf{V} are both given integer n , cm - the purported commitment to f , $u \in \mathbb{F}^t$ and $v \in \mathbb{F}$ - the purported value $\text{ml}(f)(u)$.

such that

- **Completeness:** Suppose that $\text{cm} = \text{com}(f, \text{srs})$. Then if open is run correctly with values $n, \text{cm}, u, v = \text{ml}(f)(u)$, \mathbf{V} outputs *accept* with probability one.
- **Knowledge soundness in the algebraic group model:** There exists an efficient E such that for any algebraic adversary \mathcal{A} the probability of \mathcal{A} winning the following game is $\text{negl}(\lambda)$ over the randomness of \mathcal{A} and gen .
 1. Given srs , \mathcal{A} outputs n, cm .
 2. E , given access to the messages of \mathcal{A} during the previous step, outputs $f \in \mathbb{F}^n$.
 3. \mathcal{A} outputs $u \in \mathbb{F}^t$, $v \in \mathbb{F}$.
 4. \mathcal{A} takes the part of \mathbf{P} in the protocol open with inputs n, cm, u, v .
 5. \mathcal{A} wins if
 - \mathbf{V} outputs *accept* at the end of the protocol.
 - $\text{ml}(f)(u) \neq v$.

4 Components

In this section we go over known components (with some new optimizations), that will be used in our main protocol in Section 6. The treatment will be semi-formal, and assume basic familiarity with the KZG polynomial commitment scheme [KZG10]. The formal treatment will be part of the description and knowledge soundness proof of the main protocol in Section 6.

4.1 Inner products in $O(b \log b)$ time.

Fix polynomials $g_1(X) = \sum_{i=0}^{d_1} a_i X^i, g_2 = \sum_{i=0}^{d_2} b_i X^i$ in $\mathbb{F}[X]$. We define $\langle g_1, g_2 \rangle$ to be $\sum_{i=0}^d a_i b_i$ where $d := \min\{d_1, d_2\}$. We present a convenient way to verify inner products $\langle g_1, g_2 \rangle$ similar to [BCC⁺16, MBKM19]. The basic observation is that $\langle g_1, g_2 \rangle$ is the constant coefficient of the rational function $R(X) := g_1(X)g_2(1/X)$. Thus, $\langle g_1, g_2 \rangle = v$ is equivalent to the existence of polynomials S_1, S_2 such that

$$g_1(X)g_2(1/X) = 1/X \cdot S_1(1/X) + v + X \cdot S_2(X).$$

We can thus send commitments to S_1, S_2 as proof of the correctness of v . As an optimization, we observe that we can “symmetrize” R and look instead at the rational function

$$R'(X) := g_1(X)g_2(1/X) + g_1(1/X)g_2(X).$$

The advantage of R' is that the negative and positive coefficients are equal. Thus, $\langle g_1, g_2 \rangle = v$ is equivalent to the existence of $S(X) \in \mathbb{F}[X]$ such that

$$g_1(X)g_2(1/X) + g_1(1/X)g_2(X) = 2v + X \cdot S(X) + (1/X)S(1/X).$$

Claim 4.1. *Suppose $g_1(X), g_2(X) \in \mathbb{F}_{<b}[X]$. Let $S(X)$ be as defined above. Then S can be computed in $O(b \log b)$ \mathbb{F} -operations.*

Proof. When $g_1(X), g_2(X) \in \mathbb{F}_{<b}[X]$ we multiply the equation above by X^{b-1} to get

$$X^{b-1}(g_1(X)g_2(1/X) + g_1(1/X)g_2(X)) = X^{b-1}(2v + X \cdot S(X) + (1/X)S(1/X)).$$

We can use an $O(b \log b)$ time FFT to evaluate the LHS on $2b$ points. We then do an inverse FFT to get the coefficients c_0, \dots, c_{2b-2} of the LHS. Now, we can output $S = (c_b, \dots, c_{2b-2})$. \square

Batching inner product checks Suppose we now have two inner product claims $\langle g_1, g_2 \rangle = v_1$ and $\langle h_1, h_2 \rangle = v_2$. If we choose random $\gamma \in \mathbb{F}$ with high probability

Claim 4.2. *Fix polynomials $g_1, g_2, h_1, h_2 \in \mathbb{F}[X]$ and $v_1, v_2 \in \mathbb{F}$. Suppose $\langle g_1, g_2 \rangle \neq v_1$ or $\langle h_1, h_2 \rangle \neq v_2$. Then, e.w.p $1/|\mathbb{F}|$ over $\gamma \in \mathbb{F}$ there does not exist $S(X) \in \mathbb{F}[X]$ such that*

$$\begin{aligned} g_1(X)g_2(1/X) + g_1(1/X)g_2(X) + \gamma(h_1(X)h_2(1/X) + h_1(1/X)h_2(X)) \\ = 2(v_1 + \gamma v_2) + X \cdot S(X) + (1/X)S(1/X) \end{aligned}$$

Proof. Denote $(v'_1, v'_2) = (\langle g_1, g_2 \rangle, \langle h_1, h_2 \rangle)$. The constant coefficient of the LHS is $v'_1 + \gamma v'_2$. To satisfy the equation in the claim for some S , we need

$$v'_1 + \gamma v'_2 = v_1 + \gamma v_2,$$

which can hold for at most one γ when $(v_1, v_2) \neq (v'_1, v'_2)$. \square

4.2 Multilinear evaluations as inner products of univariate polynomials.

For $u \in B_t$ define the polynomial $P_u(X) := \sum_{i < b} \mathbf{eq}(i, u) X^i$. Note that for $g(X) \in \mathbb{F}_{<b}[X]$, we have

$$\langle P_u, g \rangle = \sum_{i < b} \mathbf{eq}(i, u) g_i = \mathbf{ml}(g)(u).$$

As leveraged in [?], we have the product formula

$$P_u(X) = \prod_{i=0}^{t-1} \left(u_i X^{2^i} + 1 - u_i \right),$$

implying $P_u(X)$ can be evaluated in $O(t)$ \mathbb{F} -operations.

Using the polynomials P_u , multilinear evaluations can be proven in a batched manner based on Claim 4.2: Suppose we want to show given committed univariates $g, h \in \mathbb{F}[X]$, $u_1, u_2 \in \mathbb{F}^t$, $v_1, v_2 \in \mathbb{F}$ that $\mathbf{ml}(g)(u_1) = v_1$ and $\mathbf{ml}(h)(u_2) = v_2$.

1. \mathbf{V} sends random $\gamma \in \mathbb{F}$.
2. \mathbf{P} sends commitment to S such that

$$\begin{aligned} g(X)P_{u_1}(1/X) + g(X)P_{u_1}(1/X) + \gamma(h(X)P_{u_2}(1/X) + h(1/X)P_{u_2}(X)) \\ = 2(v_1 + \gamma v_2) + X \cdot S(X) + (1/X)S(1/X). \end{aligned}$$

3. \mathbf{V} chooses a random $\mathfrak{z} \in \mathbb{F}$.
4. \mathbf{P} sends and proves correctness of the values of g, h and S on $\mathfrak{z}, 1/\mathfrak{z}$.
5. \mathbf{V} evaluates P_{u_1}, P_{u_2} at $\mathfrak{z}, 1/\mathfrak{z}$.
6. \mathbf{V} checks the above equation holds at \mathfrak{z} .

4.3 Degree checks

The idea presented here is from [Tha23]. Suppose \mathbf{P} wants to prove to \mathbf{V} that \mathbf{cm} is a commitment to a polynomial $g \in \mathbb{F}_{<b}[X]$. Let $D(X) := X^{b-1}g(1/X)$. The idea is that D is a polynomial if and only if $g(X)$ has degree $< b$. Thus, assuming our structured reference string doesn't contain negative powers, \mathbf{P} can commit to D if and only if $g(X) \in \mathbb{F}_{<b}[X]$.

This motivates the following protocol.

1. \mathbf{P} sends a commitment \mathbf{d} to $D(X)$.
2. \mathbf{V} chooses random $\mathfrak{z} \in \mathbb{F}$.
3. \mathbf{P} sends $D_{\mathfrak{z}} := D(\mathfrak{z}), \bar{g}_{\mathfrak{z}} := g(1/\mathfrak{z})$, and uses KZG to prove their correctness.
4. \mathbf{V} can now check D 's correctness on \mathfrak{z} , using the equation

$$D_{\mathfrak{z}} \stackrel{?}{=} \mathfrak{z}^{b-1} \bar{g}_{\mathfrak{z}}.$$

5 Univariate division

Our protocol crucially relies on the following simple claim about division by a polynomial of the form $X^b - \alpha$.

Claim 5.1. *Fix integers $b > 0$ and let $n = b^2$. Fix $\alpha \in \mathbb{F}$, and $f(X) \in \mathbb{F}_{<n}[X]$. Let $f_0(X), \dots, f_{b-1}(X) \in \mathbb{F}_{<b}[X]$ be such that $f(X) = \sum_{i<b} X^i f_i(X^b)$. Let $g(X) \in \mathbb{F}_{<b}[X], q(X) \in \mathbb{F}[X]$ be such that*

$$f(X) = (X^b - \alpha) \cdot q(X) + g(X).$$

Then,

1. $g(X) = \sum_{i<b} X^i f_i(\alpha)$.
2. The coefficients of $q(X)$ can be computed in $O(n)$ \mathbb{F} -operations.

Proof. To see the first item, note that reduction mod $X^b - \alpha$ corresponds to substituting α into X^b inside each $f_i(X^b)$ in the expression $\sum_{i<b} X^i f_i(X^b)$. We proceed to the computation of $q(X)$. We compute for each $0 \leq i < b$, the coefficients of the quotient $q_i(X) \in \mathbb{F}[X]$ such that

$$f_i(X) = q_i(X)(X - \alpha) + f_i(\alpha).$$

Using Horner's method for division by the linear polynomial $X - \alpha$ this requires only n multiplications and additions in \mathbb{F} . Now, we have that

$$f(X) = \sum_{i<b} X^i f_i(X^b) = \sum_{i<b} X^i \left(q_i(X^b)(X^b - \alpha) + f_i(\alpha) \right) = q(X)(X^b - \alpha) + g(X),$$

for $q(X) := \sum_{i<b} X^i q_i(X^b)$. Thus, the coefficients of $q(X)$ are simply the interleaving of the coefficients of the $\{q_i(X)\}$. \square

6 Main Construction

MERCURY is the tuple $(\text{gen}, \text{com}, \text{open})$ described next.

gen(n): Choose random $x \in \mathbb{F}$ and outputs $\{[1]_1, [x]_1, \dots, [x^{n-1}]_1, [1]_2, [x]_2\}$

com(n, f, srs): Output $\sum_{i<b} \sum_{j<b} f_{i,j} \cdot [x^{i \cdot b + j}]_1$.

open($n, \text{cm}, u, v; f$):

1. Committing to partial sums:

- (a) \mathbf{P} computes the polynomial to $h(X) := \sum_{i < b} \mathbf{eq}(i, u_1) f_i(X)$. Note that the coefficient of X^j in $h(X)$ is $\sum_{i < b} \mathbf{eq}(i, u_1) f_{i,j}$ - hence we think of it as a commitment to partial sums.
 - (b) \mathbf{P} computes and sends $\mathbf{h} := [h(x)]_1$.
2. Committing to “folded” polynomial g :
- (a) \mathbf{V} sends random $\alpha \in \mathbb{F}$.
 - (b) \mathbf{P} computes polynomials $g(X) \in \mathbb{F}_{<b}[X]$ and $q(X) \in \mathbb{F}[X]$ such that

$$f(X) = (X^b - \alpha) \cdot q(X) + g(X).$$

- (c) \mathbf{P} computes and sends $\mathbf{q} := [q(x)]_1$ and $\mathbf{g} := [g(x)]_1$.
3. Sending proofs of correctness for h and the degree of g :
- (a) \mathbf{V} sends a random batching challenge $\gamma \in \mathbb{F}$.
 - (b) \mathbf{P} computes and sends $\mathbf{s} = [S(x)]_1$ where $S(X) \in \mathbb{F}[X]$ is such that

$$\begin{aligned} g(X)P_{u_1}(1/X) + g(1/X)P_{u_1}(X) + \gamma \cdot (h(X)P_{u_2}(1/X) + h(1/X)P_{u_2}(X)) \\ = h(\alpha) + \gamma \cdot v + X \cdot S(X) + (1/X)S(1/X). \end{aligned}$$

- (c) \mathbf{P} computes and sends $\mathbf{d} := [D(x)]_1$ where

$$D(X) := X^{b-1}g(1/X).$$

4. KZG evaluations:

- (a) \mathbf{V} sends a random evaluation challenge $\mathfrak{z} \in \mathbb{F}$.
- (b) \mathbf{P} sends the values $f_{\mathfrak{z}} := f(\mathfrak{z}), q_{\mathfrak{z}} := q(\mathfrak{z}), g_{\mathfrak{z}} := g(\mathfrak{z}), \bar{g}_{\mathfrak{z}} := g(1/\mathfrak{z}), h_{\mathfrak{z}} := h(\mathfrak{z}), \bar{h}_{\mathfrak{z}} := h(1/\mathfrak{z}), h_{\alpha} := h(\alpha), s_{\mathfrak{z}} := s(\mathfrak{z}), \bar{s}_{\mathfrak{z}} := s(1/\mathfrak{z}), D_{\mathfrak{z}} := D(\mathfrak{z})$.
- (c) \mathbf{V} sends a random KZG batching challenge $\eta \in \mathbb{F}$.
- (d) \mathbf{P} computes and sends the KZG opening proof $\pi_{\mathfrak{z}}$ for the values $f_{\mathfrak{z}}$ and $q_{\mathfrak{z}}$. That is $\pi_{\mathfrak{z}} := [H(x)]_1$ for

$$H(X) := \frac{f(X) - f(\mathfrak{z}) + \eta(q(X) - q(\mathfrak{z}))}{X - \mathfrak{z}}.$$

- (e) \mathbf{P} computes and send a batched KZG opening proof π' for the rest of the values sent in step 4b, as described in Section 4 of [BDFG20].
- (f) \mathbf{V} checks the proof $\pi_{\mathfrak{z}}$ as in [KZG10]:

$$e(\mathbf{cm} - [f_{\mathfrak{z}}]_1 + \eta(\mathbf{q} - [q_{\mathfrak{z}}]_1), [1]_2) = e(\pi_{\mathfrak{z}}, [x]_2).$$

- (g) \mathbf{V} checks the opening proof π' as described in [BDFG20].
- (h) \mathbf{V} checks the equation

$$g_3 P_{u_1}(1/\mathfrak{z}) + \bar{g}_3 P_{u_1}(\mathfrak{z}) + \gamma(h_3 P_{u_2}(1/\mathfrak{z}) + \bar{h}_3 P_{u_2}(\mathfrak{z})) = h_\alpha + \gamma v + \mathfrak{z} s_3 + (1/\mathfrak{z}) \bar{s}_3.$$

- (i) \mathbf{V} checks the equation $D_{\mathfrak{z}} = \mathfrak{z}^{b-1} \bar{g}_3$.
- (j) If one of the checks in steps 4f-4i fails, \mathbf{V} outputs *reject*. Otherwise \mathbf{V} outputs *accept*.

Runtime of \mathbf{P} : Computing $q(X)$ in step 2b requires $O(n)$ operations by Claim 5.1. Computing \mathbf{q} and π_3 requires two MSMs of size n . All other steps are on polynomials of size $O(b) = O(\sqrt{n})$. Thus other commitments clearly require $O(\sqrt{n})$ scalar multiplications. It is easy to see other steps require $o(n)$ \mathbb{F} -operations. The least trivial of these is perhaps the computation of $S(X)$ shown to require $O(b \log b) = o(n)$ operations in Claim 4.1.

Proving knowledge soundness: Let \mathcal{A} be an efficient algebraic adversary participating in the Knowledge Soundness game from Definition 3.3. We show its probability of winning the game is $\text{negl}(\lambda)$. Let $f \in \mathbb{F}^n$ be the vector sent by \mathcal{A} in the third step of the game such that $\text{cm} = [f(x)]_1$. As \mathcal{A} is algebraic, when sending the commitments $\mathbf{h}, \mathbf{q}, \mathbf{g}, \mathbf{s}, \mathbf{d}, \pi_3, \pi'$ during protocol execution it also sends polynomials $h(X), q(X), g(X), S(X), D(X), H(X), Q(X) \in \mathbb{F}_{<n}[X]$ such that the former are their corresponding commitments. Let E be the event that \mathbf{V} outputs *accept*. Note that the event that \mathcal{A} wins the knowledge soundness game is contained in E . E implies all pairing checks have passed. Let $A \subset E$ be the event that one of the corresponding ideal pairing checks as defined in Section 3.3 didn't pass. According to Lemma 3.2, $\text{prob}(A) = \text{negl}(\lambda)$.

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