Polynomial interpolation from multiples

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Abstract

We are given an unknown polynomial $f \in \mathbb{Z}[x]$ by a
black box which on input $a \in \mathbb{Z}$ returns a value $r_a \cdot f(a)$
for some unknown non-zero rational numbers $r_a$. If
we have appropriate upper bounds on the numerator
and denominator of $r_a$ and the degree of $f$, then
the coefficients of $f$ can be computed in probabilistic
polynomial time.

Keywords: Hidden polynomial, black box polynomial,
approximate computation, short vectors, integer lattices.

1 Introduction

There are several ways of representing a univariate polyno-
monic over a ring. The two most important data structures
are the list of coefficients, and a list of values,

 junto with a bound on the degree. The transforma-
tions between these two representations, namely evalua-
tion and interpolation, play a basic role in many applica-
tions. In more general domains, including the rings
of interest in number theory, interpolation becomes the
Chinese Remainder Algorithm.

An interesting variation of the classical task of
interpolation assumes that not the exact values are
given, but only approximations to them. This may
mean that each value may be “a little incorrect”,
for that some but not all values are correct. The
usefulness of this problem first appeared in coding
theory, with the classical Reed-Solomon codes and the
more recent tool of list decoding. A number-theoretic
version of importance in cryptography is the hidden
number problem. It is, of course, necessary to have
enough information in the incorrect values to allow
reconstruction of the true object.

In this paper, we want to reconstruct a polynomial
from approximate values all of which may be incorrect,
in some controlled fashion. The case where the difference
to the true values is bounded (and we work over a
finite field), has been solved by Shparlinski
(2002b). In this paper, we consider the interpolation
problem when the ratio to the true value is bounded appro-

precisely, we want to compute the coefficients
of a polynomial $f \in \mathbb{Z}[x]$ of degree $n$ for
which we are given a multiplicatively approximate black
box MABB($A, \alpha, \beta$) which on input $a \in \mathbb{Z}$, in time
$(n \log |A| + 1)^{O(1)}$ returns a rational multiple $g_a =
r_a \cdot f(a)$ of $f(a)$ for some unknown non-zero $r_a \in \mathbb{Q}$

Here $A, \alpha$, and $\beta$ are positive real numbers such that
we can write $r_a = k_a / m_a$ with integers $k_a, m_a$ and

$$\gcd(k_a, m_a) = 1, 0 < |k_a| \leq A^n, 0 < |m_a| \leq A^\beta.$$}

The gcd condition is convenient but obviously not
necessary. We assume that the response time of the
MABB($A, \alpha, \beta$) is polynomial in the bit length of $|g_a|
(\text{thus queries with large values of } a \text{ are more expensive})$

We design a probabilistic polynomial time inter-
polation algorithm which works for an MABB($A, \alpha, \beta$)
with any $\alpha + \beta < 1/(n+2)$.

We also consider the same problem for polynomials
if \( f \in F_p[z] \) over a finite prime field \( F_p \) of \( p \) elements with a modular multiplicatively approximate black box, \( \text{MMABB}(a, \beta, p) \), which on input \( a \in Z \) returns a multiple \( g_a = r_a \cdot f(a) \in F_p \) of the value \( f(a) \in F_p \) with some factor \( r_a \in F_p^* \) such that \( r_a \equiv k_a/m_a \mod p \) for some integers \( k_a, m_a \) satisfying
\[
\gcd(k_a, m_a) = 1, \quad 0 < |k_a| \leq p^\alpha, \quad 0 < |m_a| \leq p^\beta.
\]
We obtain similar results for this problem, but the bound on the possible values of \( \alpha \) and \( \beta \) is less generous.

The corresponding additive problem, where a black box produces a shift \( s_a + f(a) \) for some integer \( s_a \), has been solved in Shparlinski (2001) over \( F_p \), with a rather generous error bound on the size of \( s_a \) (roughly speaking, \( |s_a| \) could be up to \( p \exp(-\log^{1/2}/p) \)).

Loosely speaking, the problem we consider in this paper, as well as the above mentioned additive problem, are variants of the hidden number problem which has its origin in pioneering works of Boneh & Venkatesan (1999, 1997) and which has proved to be an invaluable tool for cryptography and computer science; see Shparlinski (2003) and also surveys given in Shparlinski (2002a,b).

It is no surprise that, as in most other works on the hidden number problem, our main tool is a lattice basis reduction algorithm.

Venkatesan Guruswami and Avi Wigderson have remarked that if the fudge factors \( r_a \) are always integers with a polynomial bound on their value, then we can divide by all integers up to the bound (which evenly divide the approximate value) and obtain a list which contains sufficiently many correct values so that the list decoding approach of Guruswami & Sudan (1999) recovers the polynomial. However in this paper, as well as in Shparlinski (2001), these factors may be chosen from exponentially large sets.

2 Lattices, linear homogeneous equations, and polynomials

Here we collect several well-known facts which form the background of our algorithm.

We review several related results and definitions on lattices which can be found in Grötschel et al. (1993). For more details and more recent references, we recommend to consult the brilliant surveys of Nguyen & Stern (2000, 2001).

Let \( \{b_1, \ldots, b_r\} \) be a set of linearly independent vectors in \( \mathbb{R}^s \). The set
\[
L = \{ z : z = c_1 b_1 + \cdots + c_r b_r, \ c_1, \ldots, c_r \in \mathbb{Z} \}
\]
is called an \( s \)-dimensional lattice with basis \( \{b_1, \ldots, b_r\} \).

If \( s = r \), the lattice \( L \) is of full rank.

To each lattice \( L \) one can naturally associate its volume
\[
\text{vol}(L) = \left( \det \left( (b_i, b_j)_{i,j=1}^s \right) \right)^{1/2},
\]
where \( (a, b) \) denotes the inner product, which does not depend on the choice of the basis \( \{b_1, \ldots, b_r\} \).

For a vector \( u \), let \( ||u|| \) denote its Euclidean norm. The famous Minkowski theorem, see Theorem 5.3.6 in Section 5.3 of Grötschel et al. (1993), gives the upper bound
\[
\text{(2.3)} \quad \min \{ ||z|| : z \in L \setminus \{0\} \} \leq s^{1/2} \text{vol}(L)^{1/s}
\]
on the shortest nonzero vector in any \( s \)-dimensional lattice \( L \) via its volume. In fact \( s^{1/2} \) can be replaced by the slightly smaller Hermite constant \( \gamma_s^{1/2} \), see Nguyen & Stern (2000, 2001).

The Minkowski bound (2.3) motivates a natural question: how to find the shortest vector in a lattice. Unfortunately, there are several indications that this problem is \( \text{NP} \)-complete (when the dimension grows), see Nguyen & Stern (2000, 2001). However, for a relaxed task of finding a short vector, the celebrated \( \text{LLL} \) algorithm of Lenstra et al. (1982) provides a desirable solution.

To simplify our calculations we use the \( \text{LLL} \) algorithm in its original form. Later developments of Schnorr (1987) and quite recently by Ajtai et al. (2001) lead to some (rather slight) improvements of our results.

We also assume that the basis \( \{b_1, \ldots, b_r\} \) of \( L \) consists of vectors in \( \mathbb{Z}^r \) (rather than in \( \mathbb{R}^r \)), so one can talk about the bit size of the basis and the notion of a polynomial-time algorithm.

**Lemma 2.1.** There exists a deterministic polynomial-time algorithm which, given a basis for an \( s \)-dimensional lattice \( L \), finds a nonzero vector \( v \in L \) with
\[
||v|| \leq 2^{s-1/2} \min \{ ||z|| : z \in L \setminus \{0\} \}.
\]

It is also useful to remember that when \( s \) is small (for example, constant), then in polynomial time one can find a nonzero vector \( v \in L \) with \( ||v|| = \min \{ ||z|| : z \in L \setminus \{0\} \} \), see Ajtai et al. (2001); Nguyen & Stern (2000, 2001).

The set of integer solutions \( z = (z_1, \ldots, z_r) \in \mathbb{Z}^r \) of
\[
\sum_{i=1}^r z_i c_i = 0
\]
forms a lattice of dimension \( r - 1 \) (unless \( c_1 = \cdots = c_r = 0 \)). The same is also true for the solutions \( z \in \mathbb{Z}^r \) of a linear homogeneous congruence
\[
\sum_{i=1}^r z_i c_i \equiv 0 \mod p
\]
modulo a prime $p$, except that in this case the set of integer solutions forms a lattice of dimension $r$.

Finally, for any polynomial $f \in \mathbb{K}[x]$ of degree at most $n$ over any field $\mathbb{K}$ we have the well-known Newton relations

$$\sum_{i=0}^{n+1} (-1)^i \binom{n+1}{i} f(x + i) = 0.$$

Moreover, if $\deg f = n$ and the characteristic of $\mathbb{K}$ is greater than $n + 1$ or zero then, because $f$ can also be viewed as a linear recurrence sequence of order exactly $n+1$, any other linear relation between all $(n+1)$-tuples of consecutive values of $f$ has coefficients proportional to those in (2.4). That is, if for some coefficients $C_0, \ldots, C_{n+1} \in \mathbb{K}$

$$\sum_{i=0}^{n+1} (-1)^i C_i f(x + i) = 0$$

for all positive integer $x$, then

$$C_i = \lambda (-1)^i \binom{n+1}{i}, \quad 0 \leq i \leq n + 1,$$

and some $\lambda \in \mathbb{K}$.

3 Polynomials over $\mathbb{Z}$

There is some nonuniqueness inherent in our problem. Namely, when the black box chooses all its fudge factors $r_a$ as even integers, then we can in principle not tell whether the original polynomial is $f$ or $2f$. Therefore we can at best expect a constant multiple of the original polynomial as output.

We call an integer polynomial primitive if the greatest common divisor of its coefficients is 1, and its leading term is positive. Each nonzero integer polynomial has a unique constant multiple which is primitive, and we make this our target, thus removing the non-uniqueness just mentioned.

As usual we define the height of a polynomial $f = \sum_{i=0}^{n} f_i x^i \in \mathbb{Z}[x]$ as

$$\mathcal{H}(f) = \max_{0 \leq j \leq n} |f_j|.$$

We consider the family $\mathcal{P}_n(H)$ of primitive polynomials of degree $n$ and of height at most $H$:

$$\mathcal{P}_n(H) = \left\{ f(x) = \sum_{i=0}^{n} f_i x^i \in \mathbb{Z}[x]: \begin{array}{l} \gcd(f_0, \ldots, f_n) = 1, \\ f_n \geq 1, \ \mathcal{H}(f) \leq H \end{array} \right\}.$$

Theorem 3.1. Let $n$ and $H$ be positive integers, and let $\alpha, \beta, \delta$ be nonnegative real numbers with $0 < \delta < 1$ and $\alpha + \beta \leq (1 - \delta)/(n + 2)$. There is a deterministic algorithm which, for any $f \in \mathcal{P}_n(H)$, computes the coefficients of $f$ in time polynomial in $n, \log H, \log A$, and $\delta^{-1}$ using an MABR($A, \alpha, \beta$) for $f$.

Proof. For each factor $r_a$ we always write $r_a = k_a/m_a$ with $k_a$ and $m_a$ of the form (1.1).

Let $f \in \mathcal{P}_n(H)$ and $c = (c_0, \ldots, c_{n+1}) \in \mathbb{Z}^{n+2}$ be the vector of coefficients in (2.4), that is,

$$c_i = (-1)^i \binom{n+1}{i}, \quad 0 \leq i \leq n + 1.$$

For a vector $b = (b_0, \ldots, b_{n+1}) \in \mathbb{Z}^{n+2}$, we consider the polynomial

$$h_b = \sum_{i=0}^{n+1} b_i f(x + i) \in \mathbb{Z}[x].$$

In particular, we have $h_0 = 0$. Moreover, as discussed in Section 2, $h_b$ is 0 if and only if the vectors $b$ and $c$ are proportional.

We let

$$B = \left[ \left( (n + 2)^{2n+2}(2A)^{2n+4}H \right)^{1/5} \right],$$

and consider nonzero vectors $b = (b_0, \ldots, b_{n+1}) \in [-B, B]^{n+2}$. We let

$$B = \{ b \in [-B, B]^{n+2} : b \text{ not proportional to } c \}.$$

In particular, $h_b$ is not identically zero for any $b \in B$. A Taylor expansion at 0 of the $i$th summand of $h_b$ gives

$$h_b = \sum_{j=0}^{n+1} \sum_{i=0}^{n} b_i f^{(i)}(0) j! x^j = \sum_{j=0}^{n+1} \sum_{i=0}^{n} b_i \binom{n}{i} f_i x^{i-j}.$$

Estimating each term trivially we obtain for $b \in B$

$$\mathcal{H}(h_b) \leq (n + 2)(n + 1)^{n+1} 2^n B H(f) \leq (n + 2)^{n+2} 2^n BH.$$

Hence for $b \in B$, the absolute value of any zero of $h_b$ is less than $\mathcal{H}(h_b) + 1 < (n + 2)^{n+2} BH$.

We now set $a = (n + 2)^{2n+2} BH$ and query the approximate black box for $f$ with inputs $a+i$ to receive the values

$$g_{a+i} = r_{a+i} f(a+i) = \frac{k_{a+i}}{m_{a+i}} f(a+i) \quad \text{for } 0 \leq i \leq n + 1.$$

All values $g_{a+i}$ are nonzero because $a \geq \mathcal{H}(f) + 1$. We consider the $(n+1)$-dimensional lattice

$$L = \left\{ z = (z_0, \ldots, z_{n+1}) \in \mathbb{Z}^{n+2} : \sum_{i=0}^{n+1} z_i g_{a+i} = 0 \right\}.$$
and the integers
\[ K = \prod_{j=0}^{n+1} k_{a+j}, \quad M = \prod_{j=0}^{n+1} m_{a+j}, \]
and
\[ K_i = \frac{k_{a+i}}{m_{a+i}}, \quad M_i = \frac{m_{a+i}K}{k_{a+i}} \quad \text{for } 0 \leq i \leq n + 1. \]

Then
\[ \lvert K_i \rvert \leq A^{n+2}(a+i)^{\alpha(\alpha+1)/2}, \quad \lvert M_i \rvert \leq A^{n+2}(a+i)^{\alpha(n+1)+\beta} \]
for \( 0 \leq i \leq n + 1. \) We see that \( L \) contains the “short” vector \( u = (u_0, \ldots, u_{n+1}) \) with
\[ u_i = c_i M_i \quad \text{for } 0 \leq i \leq n + 1. \]

By (3.5) its norm satisfies
\[ \lvert u \rvert \leq A^{n+2}(a + n + 1)^{\alpha(n+1)+\beta} \left( \sum_{i=0}^{n+1} c_i^2 \right)^{1/2} \]
\[ \leq A^{n+2}(a + n + 1)^{\alpha(n+1)+\beta} \sum_{i=0}^{n+1} c_i \]
\[ = 2^{n+1} A^{n+2}(a + n + 1)^{\alpha(n+1)+\beta} \]
and thus the algorithm of Lemma 2.1 returns a vector \( v = (a_0, \ldots, a_{n+1}) \in L \) with
\[ \lvert v \rvert \leq 2^{(n+1)/2} \lvert u \rvert \leq 2^{3(n+1)/2} A^{n+2}(a + n + 1)^{\alpha(n+1)+\beta}. \]

We have
\[ |e_i K_i| \leq (2A)^{2n+4}A^{2n+4}(a + n + 1)^{(\beta + \alpha)(n+2)} \leq (2A)^{2n+4}a^{(\beta + \alpha)(n+2)} \leq (2A)^{2n+4} \left( (n + 2)^2 B_H \right)^{1-\delta} \leq (n + 2)^{2n+4} H B^{1-\delta} \leq B \]
for \( 0 \leq i \leq n + 1. \) Since
\[ 0 = \sum_{i=0}^{n+1} e_i M g_{a+i} = \sum_{i=0}^{n+1} e_i K_i f(a + i), \]
it follows that \((e_0 K_0, \ldots, e_{n+1} K_{n+1})\) is proportional to \( c, \) say \( \lambda \) for some nonzero \( \lambda \in \mathbb{Q} \) and all \( i \leq n + 1. \) We now calculate the unique interpolation polynomial \( g \in \mathbb{Q}[x] \) of degree at most \( n \) with \( g(a + i) = g_{a+i} / c_i \) for \( 0 \leq i \leq n. \) (The value for \( i = n + 1 \) is ignored.) Since
\[ g(a + i) = g_{a+i} / c_i = \lambda g_{a+i} / K_i = \lambda f(a + i) / M \]
for \( 0 \leq i \leq n, \) \( g \) is a nonzero constant multiple of \( f. \) Finding \( f \) from \( g \) is trivial. The cost estimate is immediate. \( \square \)

4 Polynomials over finite fields

Here we show that the approach of Section 3 works for polynomials over \( \mathbb{F}_p \) for a prime \( p. \)

To avoid the nonuniqueness problem in the case of polynomials over \( \mathbb{F}_p \), it is more natural to consider monic polynomials rather than primitive ones. As in any ring, multiplication by integers is well-defined, so that \( aq \in \mathbb{F}_p \) for any \( a \in \mathbb{Z} \) and \( q \in \mathbb{F}_p. \) Similarly, we have \( h(a) \in \mathbb{F}_p \) for \( a \in \mathbb{Z} \) and \( h \in \mathbb{F}_p[x]. \)

We consider the family \( \mathcal{M}_{n,p} \) of monic polynomials of degree \( n: \)
\[ \mathcal{M}_{n,p} = \left\{ f(x) = \sum_{i=0}^{n} f_i x^i \in \mathbb{F}_p[x] : f_n = 1 \right\}. \]

**Theorem 4.1.** Let \( n \) be a positive integer, and let \( \alpha, \beta, \delta, \varepsilon \) be nonnegative real numbers with \( 0 < \delta, \varepsilon < 1 \)
and \( \alpha + \beta = (1 - \delta)/(n + 2)^2. \) Then for any prime
\[ p > 2^{(2n+6)(n+2)/2\delta} \varepsilon^{-1/\delta} \]
there is a probabilistic algorithm which for any \( f \in \mathcal{M}_{n,p} \) computes with probability at least \( 1 - \varepsilon \) the coefficients of \( f \) in time polynomial in \( n, \log p, \delta^{-1} \) and \( \log \varepsilon^{-1} \) using an MMABR(\( \alpha, \beta, p \)) for \( f. \)

**Proof.** For each factor \( r_k \) we always write \( r_k = k_a / m_a \) with \( k_a \) and \( m_a \) are of the form (1.2).

The assumptions imply that \( p > n + 1. \) We define the vector \( c \in \mathbb{Z}^{n+2} \) and the polynomials \( h_b \in \mathbb{F}_p[x] \) for \( b \in \mathbb{Z}^{n+2} \) in exactly the same way as in the proof of Theorem 3.1. In particular, we have \( h_c = 0, \) and \( h_b = 0 \)
if and only if the vectors \( b \) and \( c \) are proportional modulo \( p. \) We let
\[ B = \left\{ \frac{1}{\varepsilon} (\varepsilon p)^{(n+2)/2} \right\} \]
and consider nonzero vectors \( b = (b_0, \ldots, b_{n+1}) \in [-B, B]^{n+2}. \) We let
\[ B = \{ b \in [-B, B]^{n+2} : b \text{ not proportional to } c \text{ over } \mathbb{F}_p \}, \]
and for \( b \in B, \) we consider \( A_b = \{ a \in \mathbb{Z} : h_b(a) = 0 \}. \) We have \#\( A_b \leq n, \) and \( \mathcal{A} = \bigcup_{b \in B} A_b \) has at most
\[ n(2B + 1)^{n+2} < n3^{n+2} B^{n+2} \leq n^{n+2} B^{n+2} \leq \varepsilon p \]
elements. We choose \( a \) uniformly at random from \( \mathbb{F}_p, \) so that \( \text{Prob}(a \in A) < \varepsilon. \) We now query the approximate black box for \( f \) with the inputs \( a + i \) to receive the values \( g_{a+i} = r_{a+i} \cdot f(a + i) \) for \( 0 \leq i \leq n + 1. \)

Because \( f \) is not identically zero, at least one value \( g_{a+i} \in \mathbb{F}_p \) is nonzero. We consider the \((n + 1)-\)
dimensional lattice
\[ L = \left\{ z = (z_0, \ldots, z_{n+1}) \in \mathbb{Z}^{n+2} : \sum_{i=0}^{n+1} z_i g_{a+i} = 0 \right\}. \]
We also denote

\[ K = \prod_{j=0}^{n+1} k_{a+j}, \quad M = \prod_{j=0}^{n+1} m_{a+j} \]

and

\[ K_i = \frac{k_{a+i}}{m_{a+i}} M, \quad M_i = \frac{m_{a+i}}{k_{a+i}} K, \quad 0 \leq i \leq n + 1. \]

In particular

\[(4.6) \quad |K_i| \leq p^{a(n+1)\beta}, \quad |M_i| \leq p^{(n+1)\alpha + \beta}\]

for \(0 \leq i \leq n + 1\). We see that \(L\) contains the “short” vector \(u = (u_0, \ldots, u_{n+1}) \in L\) with

\[ u_i = c_i M_i \quad \text{for} \quad 0 \leq i \leq n + 1. \]

By (4.6) its norm satisfies

\[ ||u|| \leq p^{a(n+1)+\beta} \left( \sum_{i=0}^{n+1} c_i^2 \right)^{1/2} \leq p^{a(n+1)+\beta} \leq 2^{n+2} p^{a(n+1)+\beta}, \]

and the algorithm of Lemma 2.1 returns a vector \(v = (v_0, \ldots, v_{n+1}) \in L\) with

\[ ||v|| \leq 2^{(n+2)/2} ||u|| \leq 2^{n+2} p^{a(n+1)+\beta}. \]

The assumption in the theorem implies that \(B \geq 1\), and therefore

\[ B \geq \frac{B + 1}{2} \geq \frac{1}{12} p^{1/(n+2) \beta} \geq 2^{n+2} p^{(1-5)/(n+2)} \]

for all \(0 \leq i \leq n + 1\). Now we assume that \(a \notin \mathbb{A}\). Since

\[ 0 = \sum_{i=0}^{n+1} v_i M g_{a+i} = \sum_{i=0}^{n+1} v_i K_i (a + i), \]

it follows that \((v_0 K_0, \ldots, v_{n+1} K_{n+1})\) is proportional modulo \(p\) to \(c\), say \(v_i K_i = \lambda c_i\) for some integer \(\lambda \neq 0 \mod p\) and all \(i \leq n + 1\). Since \(p > n + 1\), each \(c_i\) with \(i \leq n + 1\) is nonzero modulo \(p\). We now calculate the unique interpolation polynomial \(g \in \mathbb{F}_p[x]\) of degree at most \(n\) with \(g(a+i) = g_{a+i} c_i / K_i\) for \(0 \leq i \leq n\). (The value for \(i = n + 1\) is ignored.) Since

\[ g(a+i) = g_{a+i} c_i / K_i = \lambda g_{a+i} K_i = \lambda f(a+i) / M \]

for \(i \leq n\), \(g\) is a nonzero constant multiple of \(f\), and \(f = k(g)^{-1} \cdot g\), where \(k(g)\) is the leading coefficient of \(g\). The cost estimate is immediate. □

5 Remarks

Our algorithm works with \(\alpha + \beta = O(n^{-1})\) over \(\mathbb{Z}\) and with \(\alpha + \beta = O(n^{-2})\) over \(\mathbb{F}_p\). The example of two polynomials \(xg\) and \((x+1)g\) where \(\deg g = n-1\) shows that it is impossible to solve this problem with \(\alpha > 1\). Indeed, the black box output \(a(a+1)g(a)\) is consistent with both polynomials. It is an open question by how much our bound on \(\alpha + \beta\) can be relaxed. Is there a polynomial-time method when \(\alpha + \beta = O(n^{-1})\) over \(\mathbb{F}_p\)? Finally, can the lower bound on \(p\) in Theorem 4.1 be lowered?

It is easy to see that our method can also be applied to interpolating linear recurrence sequences instead of polynomials, which satisfy a given linear recurrence relation

\[ \sum_{i=0}^{n+1} c_i u(x + i) = 0. \]

The cost of the corresponding algorithm depends on the size of the coefficients in the above relation. To justify the algorithm one can use bounds on the number of zeros of linear recurrence sequences over \(\mathbb{Z}\) and in finite fields; see Everest et al. (2002), or the original papers Entervise & Schlickewei (1999); van der Poorten & Schlickewei (1991); Schlickewei et al. (1999); Schmidt (2000).

One can obtain an algorithm which interpolates a \(t\)-sparse polynomial with only \(t + 1\) queries to an MABB\((A, \alpha, \beta)\) or an MMO\((A, \alpha, \beta, p)\); the whole algorithm however remains polynomial in \(n\), not in \(t\) and \(\log n\) as one would desire for \(t\)-sparse polynomials.

One can also consider polynomials with rational coefficients for which there is black box returning rational multiples of their values with controlled numerator and denominator.

Our method, combined with the method of Shparlinski (2001), can be applied to the more general problem of recovering \(k\) polynomials \(f_1, \ldots, f_k\) from polynomially many vectors \(R_a \cdot (f_1(a), \ldots, f_k(a))^T + s_a\) for some “small” matrices \(R_a\) and vectors \(s_a\).

Finally, our method can be extended to polynomials \(f(x_1, \ldots, x_m)\) in \(m \geq 2\) variables by making Kronecker queries of the form \(f(a, a^{d+1}, \ldots, a^{d+1})^{m-1})\) and thus reducing this case to the univariate case.

We conclude by mentioning that it would be interesting to consider an analog of our problem over a residue ring of the integers modulo an integer. One of our crucial ingredients, namely the fact that the number of roots of a polynomial is bounded by its degree, does not apply anymore. However, over such rings one can try to use a much weaker bound of Konyagin (1979).

Instead of recovering a hidden polynomial, we can ask to construct a number \(x\) given by a black box
returning $r_p x \mod p$ on input $p$, for a prime $p$ and a small rational number $r_p$. One will have additional constraints, such as bounds on $|x|$ and on the primes. This problem is now under consideration.

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**References**


