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ABSTRACT

A method for dividing a polynomial of degree (2n-1) by a precomputed nth degree polynomial in $O(n \log n)$ arithmetic operations is given. This is used to prove that the evaluation of an nth degree polynomial at n+1 arbitrary points can be done in $O(n \log^2 n)$ arithmetic operations, and consequently, its dual problem, interpolation of an nth degree polynomial from n+1 arbitrary points can be performed in $O(n \log^2 n)$ arithmetic operations. The best previously known algorithms required $O(n \log^3 n)$ arithmetic operations.

1. INTRODUCTION

Given (x_i, y_i) $(0 \le i \le n)$, the interpolation problem is the determination of the coefficients $\{c_i\}$ $(0 \le i \le n)$ of the unique polynomial $P(x) = \sum_{0 \le i \le n} c_i x^i$ of degree $\le n$ such that $P(x_i) = y_i$ $(0 \le i \le n)$. If a classical method such as the Lagrange or Newton formula is used, interpolation takes $O(n^2)$ operations. (In this paper all arithmetic operations will be counted. We simply write operations to denote arithmetic operations.) However, Horowitz (1972) has shown that interpolation can be done in $O(n \log^3 n)$ operations by using the Fast Fourier Transform (FFT), and he has shown that interpolation is reducible to evaluation of an nth degree polynomial at n+1 points. Moenck and Borodin (1972) have shown that the evaluation problem is reducible to the division problem, and they have shown that both evaluation and interpolation can be done in $O(n \log^3 n)$ operations, and precomputed interpolation (knowing the x_i in advance) can be performed in $O(n \log^2 n)$ operations. The purpose of this paper is to show that, without using any precomputation, both evaluation and interpolation can be done in $O(n \log^2 n)$ operations. As a corollary we show that an nth degree polynomial and all its derivatives can be evaluated at any point in $O(n \log^2 n)$ operations.

We shall use the same approach as used by Moenck and Borodin (1972). But we shall first precompute all necessary divisors in $O(n \log^2 n)$ operations so that each division can be done in $O(n \log n)$ operations. This results in faster evaluation and faster interpolation.

After the work reported here was completed, the author received a report from V. Strassen, entitled, "Die Berechnungskomplexität von elementarsymmetrischen Funktionen und von Interpolationskoeffizienten". Using different techniques Strassen proves that interpolation can be done in $O(n \log n)$ multiplications or divisions and he states that his techniques can be used to prove that interpolation can be done in $O(n \log^2 n)$ arithmetic operations.

2. PRELIMINARIES

We shall work over the field of complex numbers.

Theorem 2.1. (Fast Polynomial Multiplication)

Let $A(x) = \sum_{0\le i\le n} a_i x^i$ and $B(x) = \sum_{0\le i\le n-1} b_i x^i$ be any two polynomials. Let $A(x)\cdot B(x) = \sum_{0\le i\le 2n-1} c_i x^i$. Then $\{c_i\}$ $(0 \le i \le 2n-1)$ can be obtained in $O(n \log n)$ operations.

Theorem 2.2.

Let $\{a_i\}$ $(0 \le i \le n)$ and $\{b_i\}$ $(0 \le i \le n-1)$ be any two sequences of numbers. Then

$$(2.1)\qquad \begin{bmatrix} a_n & a_{n-1} & \cdots & a_1 \\ & \ddots & \ddots & \vdots \\ & & \ddots & a_{n-1} \\ & & & a_n \end{bmatrix} \begin{bmatrix} b_0 \\ \vdots \\ b_{n-2} \\ b_{n-1} \end{bmatrix}$$

can be computed in $O(n \log n)$ operations.

Proof.

Let $A(x) = \sum_{0\le i\le n} a_i x^i$ and $B(x) = \sum_{0\le i\le n-1} b_i x^i$. Suppose that $A(x)\cdot B(x) = \sum_{0\le i\le 2n-1} c_i x^i$. It is clear that the computation of (2.1) is equivalent to the computation of $\{c_i\}$ $(n \le i \le 2n-1)$. Thus the proof follows from Theorem 2.1.

QED

<u>Theorem 2.3</u>.

For any sequence $\{a_i\}$ $(0 \le i \le n)$ of numbers with $a_n \neq 0$, let $\sum_{0\le i\le n-1} \bar{a}_i x^i$ be the unique polynomial $q(x)$ such that

$$x^{2n-1} = q(x)\cdot\Big(\sum_{0\le i\le n} a_i x^i\Big) + r(x),\ \deg r < n. \tag{2.2}$$

Then, we have that

$$\begin{bmatrix} a_n & a_{n-1} & \cdots & a_1 \\ & \ddots & \ddots & \vdots \\ & & \ddots & a_{n-1} \\ & & & a_n \end{bmatrix}^{-1} = \begin{bmatrix} \bar{a}_{n-1} & \bar{a}_{n-2} & \cdots & \bar{a}_0 \\ & \ddots & \ddots & \vdots \\ & & \ddots & \bar{a}_{n-2} \\ & & & \bar{a}_{n-1} \end{bmatrix} \tag{2.3}$$

and the sequence $\{\bar{a}_i\}$ $(0 \le i \le n-1)$ can be obtained in $O(n \log^2 n)$ operations.

<u>Proof</u>.

Let $\Big(\sum_{0\le i\le n} a_i x^i\Big)\cdot\Big(\sum_{0\le i\le n-1} \bar{a}_i x^i\Big) = \sum_{0\le i\le 2n-1} c_i x^i$. Then $\sum_{0\le i\le 2n-1} c_i x^i = x^{2n-1} - r(x)$. Since $\deg r < n$, $c_{2n-1} = 1$ and $c_i = 0$ for $i = n, n+1, \ldots, 2n-2$. Therefore,

$$\begin{bmatrix} a_n & a_{n-1} & \cdots & a_1 \\ & \ddots & \ddots & \vdots \\ & & \ddots & a_{n-1} \\ & & & a_n \end{bmatrix} \begin{bmatrix} \bar{a}_0 \\ \vdots \\ \bar{a}_{n-2} \\ \bar{a}_{n-1} \end{bmatrix} = \begin{bmatrix} 0 \\ \vdots \\ 0 \\ 1 \end{bmatrix}. \tag{2.4}$$

Furthermore, from (2.4), one can easily show that, for any i $(1 \le i \le n)$,

$$\begin{bmatrix} a_n & a_n & \cdots & a_1 \\ & \ddots & \ddots & \vdots \\ & & \ddots & \vdots \\ & & & a_{n-1} \\ & & & a_n \end{bmatrix} \begin{bmatrix} \bar{a}_{n-i} \\ \vdots \\ \bar{a}_{n-1} \\ 0 \\ \vdots \\ 0 \end{bmatrix} = \begin{bmatrix} 0 \\ \vdots \\ 0 \\ 1 \\ 0 \\ \vdots \\ 0 \end{bmatrix} \leftarrow \text{ith component}$$

This proves (2.3). By using the fast division algorithm given by Moenck and Borodin (1972), the unique polynomial $q(x)$ (i.e., the sequence $\{\bar{a}_i\}$ $(0 \le i \le n-1)$) can be computed in $O(n \log^2 n)$ operations. QED

<u>Definition 2.4.</u> (Precomputing)

Given any polynomial $P(x) = \sum_{0 \le i \le n} a_i x^i$ with $a_n \neq 0$, by <u>precomputing</u> $P(x)$, we shall mean the computation of the $\{\bar{a}_i\}$ $(0 \le i \le n-1)$ which are defined by (2.2) or (2.3). That is, precomputing $P(x)$ is just the division of x^{2n-1} by $P(x)$.

Hence, by Theorem 2.3, we can precompute an nth degree in $O(n \log^2 n)$ operations. Since this bound will be sufficient to prove the results in this paper, no attempt has been made to improve it.

3. FAST DIVISION USING PRECOMPUTED DIVISOR

Theorem 3.1.

Let $U(x) = \sum_{0\le i\le 2n-1} u_i x^i$ and $V(x) = \sum_{0\le i\le n} v_i x^i$ ($v_n \neq 0$). Suppose that $V(x)$ has already been precomputed, i.e., $\{\bar{v}_i\}$ $(0 \le i \le n-1)$ are available with no associated cost. Then we can compute the unique polynomials $Q(x)$ and $R(x)$ such that

$$U(x) = Q(x)\cdot V(x) + R(x),\ \deg R < n \tag{3.1}$$

in $O(n \log n)$ operations.

Proof.

It suffices to show that to compute $Q(x)$ we only require $O(n \log n)$ operations, since $R(x) = U(x) - Q(x)\cdot V(x)$ and $Q(x)\cdot V(x)$ can be computed in $O(n \log n)$ operations by Theorem 2.1. Let $Q(x) = \sum_{0\le i\le n-1} q_i x^i$, and let $Q(x)\cdot V(x) = \sum_{0\le i\le 2n-1} c_i x^i$. From (3.1), it is clear that $u_i = c_i$ for $i = n,\ldots,2n-1$. Therefore,

$$\begin{bmatrix} v_n & v_{n-1} & \cdots & v_1 \\ & \ddots & \ddots & \vdots \\ & & \ddots & v_{n-1} \\ & & & v_n \end{bmatrix} \begin{bmatrix} q_0 \\ \vdots \\ q_{n-2} \\ q_{n-1} \end{bmatrix} = \begin{bmatrix} u_n \\ \vdots \\ u_{2n-2} \\ u_{2n-1} \end{bmatrix}$$

and hence, by (2.3),

$$\begin{bmatrix} q_0 \\ \vdots \\ q_{n-2} \\ q_{n-1} \end{bmatrix} = \begin{bmatrix} \bar{v}_{n-1} & \bar{v}_{n-2} & \cdots & \bar{v}_0 \\ & \ddots & \ddots & \vdots \\ & & \ddots & \bar{v}_{n-2} \\ & & & \bar{v}_{n-1} \end{bmatrix} \begin{bmatrix} u_n \\ \vdots \\ u_{2n-2} \\ u_{2n-1} \end{bmatrix}.$$

The theorem then follows from Theorem 2.2. QED

4. FAST EVALUATION

Moenck and Borodin (1972) have shown that evaluation is reducible to division and have proved the following theorem:

<u>Theorem 4.1</u>. (Moenck and Borodin (1972))

Let $U(x)$ be a polynomial of degree $n = 2^r-1$. Then we can evaluate $U(x)$ at $n+1$ <u>arbitrary</u> points $x_0, x_1, \ldots, x_n$ in $O(g(n)\log n + f(n)\log n)$ operations, provided that we can divide a polynomial of degree $(2n-1)$ by an nth degree polynomial in $O(g(n))$ operations and multiply two nth degree polynomials in $O(f(n))$ operations.

This fast evaluation algorithm requires certain divisions. The divisors are exactly the members of the following family except the polynomial at level r+1.

$$
\begin{array}{ll}
x-x_0,\ x-x_1,\ x-x_2,\ x-x_3,\ \ldots\ ,\ x-x_{n-1},\ x-x_n & \text{Level } 1 \\
(x-x_0)(x-x_1),\ (x-x_2)(x-x_3),\ \ldots\ ,\ (x-x_{n-1})(x-x_n) & \text{Level } 2 \\
\prod_{i=0}^{3}(x-x_i),\ \ldots\ \prod_{i=n-3}^{n}(x-x_i) & \text{Level } 3 \\
\ddots \qquad \qquad \cdot^{\cdot^{\cdot}} & \\
\prod_{i=0}^{2^{r-1}-1}(x-x_i) \qquad \prod_{i=2^{r-1}}^{2^r-1}(x-x_i) & \text{Level } r \\
\prod_{i=0}^{2^r-1}(x-x_i) & \text{Level } r+1
\end{array}
\tag{4.1}
$$

<u>Theorem 4.2</u>.

All polynomials in (4.1) can be precomputed in $O(n \log^2 n)$ operations.

<u>Proof</u>.

We first convert all polynomials in (4.1) into the form $\sum_i h_i x^i$. This can be done in $O(n \log^2 n)$ operations (see Horowitz (1972)). Then we shall precompute the polynomials at level j from the precomputed polynomials at level j+1, for $j = r, r-1, \ldots, 1$. By Theorem 2.3, we can precompute the polynomial at level r+1 in $O(n \log^2 n)$ operations. Suppose that all polynomials at level j+1 have been precomputed. Let $D(x) = \sum_{0 \le i \le 2^j} d_i x^i$ be a polynomial at level j+1, and let $E(x) = \sum_{0 \le i \le 2^{j-1}} e_i x^i$ and $F(x) = \sum_{0 \le i \le 2^{j-1}} f_i x^i$ be those two polynomials such that $D(x) = E(x) \cdot F(x)$. By (2.2), we know that

$$x^{2^{j+1}-1} = \Big(\sum_{0 \le i \le 2^j - 1} \bar{d}_i x^i\Big) \cdot D(x) + r_D(x), \quad \deg r_D < 2^j.$$

Since $D(x) = E(x) \cdot F(x)$, it follows that

$$\frac{x^{2^j - 1}}{E(x)} = \frac{\Big(\sum_{0 \le i \le 2^j - 1} \bar{d}_i x^i\Big) \cdot F(x)}{x^{2^j}} + \frac{r_D(x)}{x^{2^j} E(x)} .$$

But, by (2.2),

$$\frac{x^{2^j - 1}}{E(x)} = \sum_{0 \le i \le 2^{j-1} - 1} \bar{e}_i x^i + \frac{r_E(x)}{E(x)}, \quad \deg r_E < 2^{j-1} .$$

Hence, if $\Big(\sum_{0 \le i \le 2^j - 1} \bar{d}_i x^i\Big) \cdot F(x) = \sum_{0 \le i \le 2^j + 2^{j-1} - 1} g_i x^i$, then

$$\sum_{0 \le i \le 2^{j-1} - 1} g_{i+2^j} x^i + \frac{\sum_{0 \le i \le 2^j - 1} g_i x^i}{x^{2^j}} + \frac{r_D(x)}{x^{2^j} E(x)} = \sum_{0 \le i \le 2^{j-1} - 1} \bar{e}_i x^i + \frac{r_E(x)}{E(x)} .$$

By the uniqueness of the partial fraction expansion, it is easy to see that

$$\bar{e}_i = g_{i+2^j} \text{ for all } i = 0,1,\ldots,2^{j-1}-1.$$

Therefore, we can precompute $E(x)$ by computing $(\sum_{0\le i\le 2^j-1} \bar{d}_i x^i)\cdot F(x)$, which can be performed in $O(j\cdot 2^j)$ operations by Theorem 2.1. Similarly, we can precompute $F(x)$ in $O(j\cdot 2^j)$ operations. Since there are $\frac{2^r}{2^{j-1}}$ polynomials at level j, all polynomials at level j can be precomputed in $O(\frac{2^r}{2^{j-1}}\cdot j\cdot 2^j) = O(j\cdot 2^{r+1})$ operations. Hence, all polynomials in (4.1) can be precomputed in $O(\sum_{1\le j\le r} j\cdot 2^{r+1}) = O(r^2\cdot 2^r) = O(n \log^2 n)$ operations. QED

<u>Theorem 4.3</u>.

Let $U(x)$ be a polynomial of degree $n = 2^r-1$. Then we can evaluate $U(x)$ at n+1 <u>arbitrary</u> points $x_0, x_1, \ldots, x_n$ in $O(n \log^2 n)$ operations.

<u>Proof</u>.

We first precompute all divisors needed for the algorithm of Theorem 4.1. By Theorem 4.2, this takes $O(n \log^2 n)$ operations. Then by Theorem 3.1, all divisions used in the algorithm of Theorem 4.1 can be performed in $O(n \log n)$ operations. The proof follows from Theorem 4.1 by letting $g(n) = f(n) = n \log n$. QED

5. FAST INTERPOLATION

Horowitz (1972) has shown that interpolation is reducible to fast evaluation.

Theorem 5.1. (Horowitz (1972))

Given $n+1 = 2^r$ pairs of numbers (x_i, y_i) $(0 \le i \le n)$, the coefficients of the unique polynomial $P(x)$ of degree $\le n$ such that $y_i = P(x_i)$ $(0 \le i \le n)$ can be obtained in $O(h(n) + f(n)\log n)$ operations, provided that evaluation at n+1 point is $O(h(n))$ operations and multiplication is $O(f(n))$ operations.

Theorem 5.2.

Given $n+1 = 2^r$ pairs of points (x_i, y_i) $(0 \le i \le n)$, the coefficients of the unique polynomial $P(x)$ of degree $\le n$ such that $y_i = P(x_i)$ $(0 \le i \le n)$ can be obtained in $O(n \log^2 n)$ operations.

Proof.

Apply the result of Theorem 4.3 to Theorem 5.1. QED

Corollary 5.3.

An nth degree polynomial and all its derivatives can be evaluated at any point in $O(n \log^2 n)$ operations.

Proof.

Suppose that we want to evaluate the nth degree polynomial $P(x)$ and all its derivatives at some point α. Then it suffices to show that $\{d_i\}$ $(0 \le i \le n)$ such that $P(x) = \sum_{0 \le i \le n} d_i (x-\alpha)^i$, can be obtained in $O(n \log^2 n)$ operations.

First, we evaluate $P(x)$ at $n+1$ arbitrary distinct points $x_0, x_1, \ldots, x_n$. This takes $O(n \log^2 n)$ operations by Theorem 4.3. Next, we determine $\{d_i\}$ $(0 \leq i \leq n)$ such that $\sum_{i=0}^{n} d_i y_j^i = P(x_j)$, $y_j = x_j - \alpha$ for $j = 0,1,\ldots,n$. This is an interpolation problem and takes $O(n \log^2 n)$ operations by Theorem 5.2.

QED

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