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for the Entire Sequence of the
Maximum Degree of Minors
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Combinatorial Relaxation Algorithm for the Entire Sequence of the Maximum Degree of Minors in Mixed Polynomial Matrices

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Abstract

Iwata–Takamatsu (2013) showed that the maximum degree of minors in mixed polynomial matrices for a specified order can be computed by combinatorial relaxation type algorithm. In this letter, based on their algorithm, we propose an efficient “combinatorial relaxation” algorithm for computing the entire sequence of the maximum degree of minors. In our previous work, we dealt with a similar problem for rational matrices, where the efficiency derived from the discrete concavity of valuated bimatroids. We follow the same line of discussion; but, technical details are different due to special characteristics of mixed matrices.

1 Introduction

The concept of mixed matrices [1] was introduced as a mathematical tool for description of physical systems. A mixed matrix contains two kinds of numbers as follows:

- Accurate Numbers (Fixed Constants)
They frequently represent conservation laws, and are precise in values. These numbers should be treated numerically.
- Inaccurate Numbers (Independent Parameters)
They often denote physical characteristics and are not precise in values. Since they can be assumed to be independent, these numbers should be treated combinatorially as nonzero parameters.

Let $A(x)$ be a mixed polynomial matrix over certain fields (which we will define later). R and C denote the row-set and column-set of A , respectively. $\deg f$ denotes the degree of a polynomial function f . We define the maximum degree of minors of order k as

$$\delta_k(A) := \max_{I \subseteq R, J \subseteq C} \{\deg \det A[I, J] \mid |I| = |J| = k\},$$

where $A[I, J]$ denotes the submatrix of A with $I \subseteq R$ and $J \subseteq C$. We set $\delta_0(A) = 0$. The entire sequence of maximum degree of minors $\{\delta_k(A)\}_{k=0}^r$ can be used for computing the Smith–McMillan form at infinity [1], which often used in control theory. Thus we aim to develop an efficient algorithm for computing $\{\delta_k(A)\}_{k=0}^r$.

The maximum degree of minors can be defined on rational matrices, which includes mixed polynomial matrices as special cases. Murota [2] showed that the maximum degree of minors on rational matrices for a specified order can be computed by the general framework of “combinatorial relaxation” (Murota [3]). Our previous work [4] then showed an efficient combinatorial relaxation type algorithm for finding the entire sequence of the maximum degree of minors in rational matrices. On the other hand, for mixed polynomial matrices, Iwata–Takamatsu [5] gave an algorithm for a specified order. In this letter, we extend the idea of [4] to mixed polynomial matrices and develop a combinatorial relaxation type algorithm for finding the entire sequence of the maximum degree of minors. The efficiency of the proposed algorithm is based on two theorems (Theorems 1 and 2) concerning “tightness” of combinatorial relaxations.

2 Preliminaries

2.1 Mixed Polynomial Matrices and $\delta_k(A)$ ’s

We define *mixed polynomial matrices* (see, e.g., [1]). In the definition below, fields K and F express accurate and inaccurate numbers, respectively; the coefficients of $Q_{ij}(x)$ corresponding to fixed constants, and the coefficients of $T_{ij}(x)$ to independent parameters.

Definition 1 (Mixed Polynomial Matrices) *Let K be a subfield of a field F . A polynomial matrix $M(x)$ is called a mixed polynomial matrix over (K, F) if $M(x)$ can be represented as $M(x) = Q(x) + T(x)$, where $Q(x)$ and $T(x)$ satisfy the following conditions:*

(MP-Q) $Q(x)$ is a polynomial matrix over K .

(MP-T) $T(x)$ is a polynomial matrix over F , and the set \mathcal{T} of nonzero coefficients of $T(x)$ is algebraically independent over K .

A mixed polynomial matrix $M(x)$ is called a *layered mixed (LM) polynomial matrix* if it is in the form

$$M(x) = \begin{bmatrix} Q(x) \\ T(x) \end{bmatrix}, \quad (1)$$

where $Q(x)$ and $T(x)$ satisfies **(MP-Q)** and **(MP-T)**.

Finding the maximum degree of minors in a mixed polynomial matrix can be reduced to a corresponding problem on an LM polynomial matrix. Let us start with an $m \times n$ mixed polynomial matrix $\tilde{A}(x) = \tilde{Q}(x) + \tilde{T}(x)$ with \tilde{R} and \tilde{C} being the row-set and the column-set. Then, we can associate \tilde{A} with an LM polynomial matrix

$$A(x) = \begin{bmatrix} \text{diag}[x^{d_1}, \dots, x^{d_m}] & \tilde{Q}(x) \\ \text{diag}[t_1 x^{d_1}, \dots, t_m x^{d_m}] & \tilde{T}(x) \end{bmatrix}, \quad (2)$$

where $d_i := \max_{j \in \tilde{C}} \deg \tilde{Q}_{ij}$, and t_i is a new parameter for all $i \in \tilde{R}$. We define R and C as the row-set and column-set of A . We define $\delta_k^{\text{LM}}(A)$ as follows:

$$\delta_k^{\text{LM}}(A) := \max_{I \subseteq R_T, J \subseteq C} \{\deg \det A[I \cup R_Q, J] \mid |I| = |J| - m = k\},$$

where $R_Q \subseteq R$ and $R_T \subseteq R$ denote the row-subsets corresponding to \tilde{Q} and \tilde{T} . Then, $\delta_k(\tilde{A}) = \delta_k^{\text{LM}}(A) - \sum_{i=1}^m d_i$ holds [6], and the problem with $\delta_k(\tilde{A})$'s reduces to $\delta_k^{\text{LM}}(A)$'s ($r := \text{rank } \tilde{A} = \text{rank } A - m$).

2.2 Valuated Bimatroid

A valuated bimatroid is a triple (R, C, w) , where R and C are disjoint finite sets and $w : 2^R \times 2^C \rightarrow \mathbb{R} \cup \{-\infty\}$ is a map satisfying a certain exchange axiom (see, e.g., [7, 1]). We define $S_k \subseteq 2^R \times 2^C$ and $\delta_k \in \mathbb{R}$ as follows:

$$S_k = \{(I, J) \mid |I| = |J| = k, I \subseteq R, J \subseteq C\},$$

$$\delta_k = \max\{w(I, J) \mid (I, J) \in S_k\}.$$

Proposition 1 ([7]) $\delta_{k-1} + \delta_{k+1} \leq 2\delta_k$ holds for $k = 1, 2, \dots, r-1$.

\mathcal{M}_k denotes the set of the maximizers of w :

$$\mathcal{M}_k = \{(I, J) \in S_k \mid w(I, J) = \delta_k\}.$$

Proposition 2 ([7]) For any $(I_k, J_k) \in \mathcal{M}_k$ with $1 \leq k \leq r-1$, there exist $(I_l, J_l) \in \mathcal{M}_l$ ($0 \leq l \leq r, l \neq k$) such that $I_{l-1} \subseteq I_l$ and $J_{l-1} \subseteq J_l$ ($1 \leq l \leq r$).

It is known that

$$w(I, J) := \deg \det \tilde{A}[I, J] \quad (3)$$

defines a valuated bimatroid [7, 1]. Therefore, $\{\delta_k(\tilde{A})\}$, as well as $\{\delta_k^{\text{LM}}(A)\}$, is a concave sequence by Proposition 1. Proposition 2 means that the maximizers of w have a nesting structure.

2.3 Combinatorial Relaxation of $\delta_k^{\text{LM}}(A)$

The description of this section is based on [5]. Let $A(x)$ be an LM polynomial matrix defined by (2). We define $G(A) = (R \cup C, E(A), c)$ as a bipartite graph associated with $A(x)$, where the arc set $E(A)$ and the weight $c : E(A) \rightarrow \mathbb{Z}$ are defined as follows:

$$E(A) := \{(i, j) \mid i \in R, j \in C, A_{ij}(x) \neq 0\}, \quad (4)$$

$$c(i, j) := \deg A_{ij} \quad ((i, j) \in E(A)). \quad (5)$$

For a matching M on the bipartite graph $G(A)$, we define $\partial^+ M := \partial M \cap R$ and $\partial^- M := \partial M \cap C$, where ∂M denotes the set of incident vertices of M . Then, let $\hat{\delta}_k^{\text{LM}}(A)$ be the weight of a maximum weight $(m+k)$ -matching M such that $R_Q \subseteq \partial^+ M$ in $G(A)$, i.e.,

$$\hat{\delta}_k^{\text{LM}}(A) := \max\{\sum_{e \in M} c(e) \mid M : \text{a matching in } G(A), \\ |M| = m+k, R_Q \subseteq \partial^+ M\}.$$

If there is no $(m+k)$ -matching in $G(A)$, we put $\hat{\delta}_k^{\text{LM}}(A) = -\infty$. Then, $\hat{\delta}_k^{\text{LM}}(A)$'s play the role of a combinatorial relaxation of $\delta_k^{\text{LM}}(A)$'s, and it holds that $\delta_k^{\text{LM}}(A) \leq \hat{\delta}_k^{\text{LM}}(A)$ [2]. Moreover,

$$\hat{w}(I, J) := \begin{cases} \max\{c(M) \mid \partial M = I \cup J\} & (R_Q \subseteq I) \\ -\infty & (\text{otherwise}) \end{cases},$$

as well as $w(I, J)$ in (3), defines a valuated bimatroid. Therefore, Propositions 1 and 2 hold for \hat{w} .

We can test whether $\delta_k^{\text{LM}}(A) = \hat{\delta}_k^{\text{LM}}(A)$ (which we call ‘‘tight’’) holds or not without knowing $\delta_k^{\text{LM}}(A)$ itself by utilizing the duality of linear programming as Proposition 3 below shows. The dual of the linear programming problems associated with the weighted bipartite matching problem discussed above is given as follows:

$$\begin{aligned} \text{DLP}(A, k) : \min. \quad & \sum_{i \in R} p_i + \sum_{j \in C} q_j + (m+k)t \\ \text{s. t.} \quad & p_i + q_j + t \geq c_{ij} \quad ((i, j) \in E(A)), \\ & p_i \geq 0 \quad (i \in R_T), \quad q_j \geq 0 \quad (j \in C). \end{aligned}$$

$\text{DLP}(A, k)$ has an integral optimal solution, and the optimal value is equal to $\hat{\delta}_k^{\text{LM}}(A)$. For a feasible solution (p, q, t) , we define the active rows $I^* \subseteq R$, the active columns $J^* \subseteq C$, and the tight coefficient matrix A^* as

$$\begin{aligned} I^* &= R_Q \cup \{i \in R_T \mid p_i > 0\}, \quad J^* = \{j \in C \mid q_j > 0\}, \\ A_{ij}^* &= \lim_{x \rightarrow \infty} x^{-p_i - q_j - t} A_{ij}(x). \end{aligned} \quad (6)$$

Note that the right-hand side of (6) is a bounded constant because of $p_i + q_j + t \geq c_{ij} = \deg A_{ij}$, and that computing the rank of A^* is relatively easy (but it needs the algorithm for the rank of an LM matrix [1]).

Proposition 3 ([2]) *Let (p, q, t) be an optimal dual solution. The following three conditions are equivalent:*

- $\delta_k^{\text{LM}}(A) = \hat{\delta}_k^{\text{LM}}(A)$ holds;
- There exist $I \supseteq I^*$ and $J \supseteq J^*$ such that $\text{rank } A^*[I, J] = |I| = |J| = m + k$;
- The following four conditions hold:
 - (r1) $\text{rank } A^*[R, C] \geq m + k$,
 - (r2) $\text{rank } A^*[I^*, C] = |I^*|$,
 - (r3) $\text{rank } A^*[R, J^*] = |J^*|$,
 - (r4) $\text{rank } A^*[I^*, J^*] \geq |I^*| + |J^*| - (m + k)$.

3 Proposed Algorithm

In this section, we propose an algorithm to compute $\delta_k^{\text{LM}}(A)$'s for an LM polynomial matrix $A(x)$ defined by (2). For the sake of the algorithm description, let us here suppose that $A(x)$ is a Laurent polynomial matrix. Here, a rational function f is said to be a Laurent polynomial function if there exists an integer N such that $x^N f(x)$ is a polynomial function ($-\text{ord } f$ denotes the minimum among such N 's). We define

$$d_{\max} = \max_{i,j} \deg A_{ij}, \quad d_{\min} = \min_{i,j} \text{ord } A_{ij}.$$

3.1 Theorems to Improve Efficiency

We show two theorems concerning tightness that form the basis of our algorithm. These theorems can be proved similarly as the corresponding theorems in [4].

Theorem 1 *Suppose that $\delta_k^{\text{LM}}(A) = \hat{\delta}_k^{\text{LM}}(A)$ holds and (p, q, t) is a common optimal dual solution of $\text{DLP}(A, k)$ and $\text{DLP}(A, k + 1)$. Then, $\delta_{k+1}^{\text{LM}}(A) = \hat{\delta}_{k+1}^{\text{LM}}(A) = \delta_k^{\text{LM}}(A) + t$ if and only if $\text{rank } A^* > m + k$, where A^* is the tight coefficient matrix defined by (6).*

Theorem 1 allows us to check if $\delta_{k+1}^{\text{LM}}(A) = \hat{\delta}_{k+1}^{\text{LM}}(A)$ by computing $\text{rank } A^*$ only. This value, $r^* := \text{rank } A^*$, is always greater than or equal to $m + k$. Furthermore, when $r^* > m + k + 1$, thanks to the next theorem, we obtain all of $\delta_{k+1}^{\text{LM}}(A), \dots, \delta_{r^*-m}^{\text{LM}}(A)$ at the same time, i.e., we can skip the computation of $\delta_{k+2}^{\text{LM}}(A), \dots, \delta_{r^*-m}^{\text{LM}}(A)$.

Theorem 2 *Under the assumptions of Theorem 1 and $m + k < r^*$, the following equalities and inequality hold:*

$$\begin{aligned}\delta_l^{\text{LM}}(A) &= \hat{\delta}_l^{\text{LM}}(A) = \delta_k^{\text{LM}}(A) + (l - k)t \quad (k < l \leq r^* - m), \\ \delta_l^{\text{LM}}(A) &< \delta_k^{\text{LM}}(A) + (l - k)t \quad (k = r^* - m + 1).\end{aligned}$$

3.2 The Outline of the Proposed Algorithm

The outline of the proposed algorithm is as follows.

Outline of the Proposed Algorithm

Step 0: Compute $\delta_0^{\text{LM}}(A)$ and set $k := 0$.

Step 1: Find a common optimal dual solution (p, q, t) of $\text{DLP}(A, k)$ and $\text{DLP}(A, k + 1)$.

Step 2: Test for the tightness, i.e., whether $\delta_{k+1}^{\text{LM}}(A) = \hat{\delta}_{k+1}^{\text{LM}}(A)$ or not, by using (p, q, t) and the tight coefficient matrix A^* (Theorem 1). If the equality holds, go to Step 4. Otherwise, go to Step 3.

Step 3: Modify $A(x)$ to $A'(x)$ such that $\hat{\delta}_{k+1}^{\text{LM}}(A') < \hat{\delta}_{k+1}^{\text{LM}}(A)$ and $\delta_{k+1}^{\text{LM}}(A') = \hat{\delta}_{k+1}^{\text{LM}}(A)$ hold, and go back to Step 1.

Step 4: Output $\delta_{k+1}^{\text{LM}}(A), \dots, \delta_{r^*-m}^{\text{LM}}(A)$ (Theorem 2), update $k := r^* - m$ and go back to Step 1.

For Step 0, the initialization, we define M_0 as the unique matching on submatrix $\text{diag}[x^{d_1}, \dots, x^{d_m}]$ of A , i.e., $\delta_0^{\text{LM}}(A) = \sum_{i=1}^m d_i$. Then, we set $I_0^* := R_Q$, $J_0^* := \partial^- M_0$ and $k = 0$. The other steps are discussed in Section 3.3–3.6.

3.3 Step 1: Construction of an Optimal Dual Solution

At every starting point of Step 1, the following conditions are satisfied as a result of the last loop:

1. $\deg \det A[I_k^*, J_k^*] = \sum_{(i,j) \in M_k} \deg A_{ij} = \delta_k^{\text{LM}}(A)$;
2. $R_Q \subseteq I_k^* = \partial^+ M_k$, $J_k^* = \partial^- M_k$.

In actual computation, we do not need to store I_k^* and J_k^* because they can be easily constructed from M_k . They are explicitly introduced here for a better presentation.

An optimal dual solution (p, q, t) of $\text{DLP}(A, k)$ can be constructed by solving the shortest paths problem on the auxiliary graph G_{M_k} (cf. [5]). In this step, we can adopt “reweighting” using the newest optimal dual variable (see, e.g., [8]).

3.4 Step 2: Test for Tightness

Lemma 1 is an immediate corollary of [4, Lemma 2].

Lemma 1 *Let (p, q, t) be the optimal dual solution of $\text{DLP}(A, k)$ obtained in Step 1. Then, (p, q, t) is an optimal dual solution of $\text{DLP}(A, k + 1)$.*

Lemma 1 means that $\hat{\delta}_{k+1}^{\text{LM}}(A) = \delta_k^{\text{LM}}(A) + t$ holds. Since $\hat{\delta}_{k+1}^{\text{LM}}(A) \geq \delta_{k+1}^{\text{LM}}(A) \geq (m+k+1)d_{\min}$ holds for all integer $k < r$, $\hat{\delta}_{k+1}^{\text{LM}}(A) = \delta_k^{\text{LM}}(A) + t < (m+k+1)d_{\min}$ implies $k = r$. Therefore, if $t < (m+k+1)d_{\min} - \delta_k(A)$ holds, we can set $\text{rank } A = m+k$ and halt.

Since Lemma 1 means that (p, q, t) is a common optimal dual solution of $\text{DLP}(A, k)$ and $\text{DLP}(A, k+1)$, we can adopt Theorem 1 instead of Proposition 3 to test for the tightness. Here, we need to compute the rank of an LM matrix A^* . We execute a slightly different version of the algorithm stated in [1] to ensure the property (1) of Theorem 3 below.

The rank of an LM matrix can be computed by solving the independent matching problem on a bipartite graph $G = (R_T \cup C_Q, C; E_T \cup E_Q)$, where C_Q denotes the copy of the column-set C and the sets of arcs are defined as $E_T = \{(i, j) \mid i \in R_Q, j \in C, A_{ij}^* \neq 0\}$ and $E_Q = \{(j_Q, j) \mid j \in C\}$. We can solve this problem by utilizing augmenting paths in auxiliary graph $G_M = (V, E)$, where $V = R_T \cup C_Q \cup C$ and $E = E_T \cup E_Q \cup E^+ \cup M^\circ$. Here, E^+ expresses the structure of the linear matroid with respect to $A^*[R_Q, C]$. The difference from [1] appears only in the step of initialization.

(i) We set

$$\begin{aligned} \text{base}[i] &= j \quad ((i, j) \in M_k, i \in R_Q), \\ M^\circ &= \{(j, i) \mid (i, j) \in M_k, i \in R_T\} \cup \{(j, j_Q) \mid i \in R_Q, \text{base}[i] = j\}. \end{aligned}$$

(ii) Then, we construct a constant matrix P by repeating the following procedure for $i = 1, 2, \dots, m$: we choose $(i, j) \in M_k \cap (R_Q \times C)$ in descending order of p_i , and conduct the row elimination for all rows taking A_{ij}^* as the pivot. At the end of the procedure, we obtain a constant matrix P such that

$$P = \begin{bmatrix} I_m & U \end{bmatrix} = SA^*[R_Q, C] \quad (7)$$

holds, where S is a nonsingular constant matrix which expresses the row eliminations (U is a constant matrix created by the procedure).

After the procedure (i) and (ii), we execute the algorithm for the rank of LM matrices [1]. At the end of this algorithm, P satisfies

$$\text{rank} \begin{bmatrix} P \\ A^*[R_T, C] \end{bmatrix} = \text{term-rank} \begin{bmatrix} P \\ A^*[R_T, C] \end{bmatrix}. \quad (8)$$

where $\text{term-rank } A$ is defined as the maximum matching on $G(A)$ (see, e.g., [1]).

3.5 Step 3: Matrix Modification

At the beginning of Step 3, the relation (8) holds and we have a constant matrix S satisfying (7). Then, we define $\tilde{S}(x)$ and $S(x)$ as follows:

$$\tilde{S}(x) = \begin{bmatrix} S(x) & O \\ O & I \end{bmatrix} = \text{diag}(x; p_R) \begin{bmatrix} S & O \\ O & I \end{bmatrix} \text{diag}(x; -p_R).$$

We modify the matrix $A(x)$ to $A'(x)$ as follows:

$$A'(x) = \tilde{S}(x)A(x) = \begin{bmatrix} S(x)A[R_Q, C](x) \\ A[R_T, C](x) \end{bmatrix}. \quad (9)$$

This modification makes sense, as stated in Theorem 3.

Theorem 3 *The matrix $A'(x)$ defined in (9) has the following four properties:*

- (1) $\delta_l^{\text{LM}}(A') = \delta_l^{\text{LM}}(A)$ ($l = 0, 1, \dots, r$);
- (2) $\deg \det A'[I_k^*, J_k^*] = \delta_k^{\text{LM}}(A')$;
- (3) $\hat{\delta}_k^{\text{LM}}(A') = \delta_k^{\text{LM}}(A')$;
- (4) $\hat{\delta}_{k+1}^{\text{LM}}(A') < \hat{\delta}_{k+1}^{\text{LM}}(A)$.

Proof (1) It is sufficient to show that $\tilde{S}(x)$ is biproper. This claim holds by construction of S .

- (2) By construction of A' , we obtain

$$\deg \det A'[R_Q \cup I_k^*, J_k^*] = \deg \det A[R_Q \cup I_k^*, J_k^*].$$

Since the right-hand side is equal to $\delta_k^{\text{LM}}(A)$, this equality means that the property (2) holds.

- (3) An optimal solution (p, q, t) of $\text{DLP}(A, k)$ is a feasible solution of $\text{DLP}(A', k)$. Here, A^* and A'^* denote the tight coefficient matrices of A and A' with respect to (p, q, t) . Then, the following equality holds:

$$A'^* = \begin{bmatrix} S & O \\ O & I \end{bmatrix} A^*.$$

This means that $\text{term-rank } A'^*[I_k^*, J_k^*] = m + k$ holds. Hence, by [5, Lemma 5], (p, q, t) is an optimal solution of $\text{DLP}(A', k)$. Moreover, (8) and Proposition 3 mean that $\hat{\delta}_k^{\text{LM}}(A') = \delta_k^{\text{LM}}(A')$.

- (4) We show $\text{term-rank } A'^*[R, C] = \text{rank } A'^*[R, C] = m + k$ in the discussion above. Hence, (p, q, t) is not optimal in $\text{DLP}(A', k + 1)$ by [5, Lemma 5]. \square

3.6 Step 4: Outputs and Updates

Recall that the task of Step 4 is to output $\delta_l^{\text{LM}}(A)$'s ($l = k + 1, \dots, r^* - m$) in view of Theorem 2, and then we go back to Step 1. But in order to start the process of Step 1, we need the corresponding matching $M = M_{r^*-m}$ such that

$$\sum_{(i,j) \in M} \deg A_{ij} = \deg \det A[\partial^+ M, \partial^- M] = \delta_{r^*-m}^{\text{LM}}(A)$$

holds. The key ingredient for obtaining this is the computation of $I_{r^*-m}^* = \partial^+ M_{r^*-m}$ and $J_{r^*-m}^* = \partial^- M_{r^*-m}$, which can be obtained simultaneously in the calculation of rank A^* described in section 3.4 as follows:

$$I_{r^*-m}^* = \partial M^\circ \cap R, \quad J_{r^*-m}^* = \partial M^\circ \cap C.$$

Then, M_{r^*-m} is an maximum weight bipartite matching on $G = (I_{r^*-m}^* \cup J_{r^*-m}^*, E, \gamma)$, where $E = E(A) \cap (I_{r^*-m}^* \times J_{r^*-m}^*)$. We can obtain M_{r^*-m} efficiently by augmenting paths.

4 Complexity Analysis

The time complexity of the proposed algorithm can be analyzed as Theorem 4, along the same line of discussion in Iwata–Takamatsu [5]. This theorem can be proved similarly as [4, Theorem 4].

Theorem 4 *The proposed algorithm runs in $O(dnr(n + m^2 + dm^{\omega-1}r))$ time, where $d := d_{\max} - d_{\min}$.*

This complexity of the proposed algorithm is almost the same as the primitive one, which simply repeats the algorithm [5] for a specified order (even a little bit worse than the primitive one in special cases). However, since the step of modification is seldom executed [5, Proposition 1], the time complexity of the computation for rank A^* is crucial. As stated in Proposition 4 below, The number of iterations of Step 2 is only $(O(\sqrt{dr}) + s)$, whereas the primitive one computes the rank $4(r + s)$ times, where s denotes the number of modifications. Hence, the proposed algorithm might be faster if d and s are sufficiently small.

Proposition 4 *Let $A(x)$ be an LM Laurent polynomial matrix defined as (2). Then, the proposed algorithm executes Step 2 at most $(f(d, r) + s)$ times, where s denotes the number of modifications and $f(d, r)$ is defined as follows:*

$$f(d, r) = \begin{cases} \left\lfloor \frac{-1 + \sqrt{8dr + 1}}{2} \right\rfloor & (r \geq 2d - 1) \\ r & (r < 2d - 1) \end{cases}. \quad (10)$$

Proof If t denotes the number of iterations of Step 4, Step 2 is executed $(t+s)$ times. By the latter part of Theorem 2, t is same as the number of k 's such that $\delta_{k+1}^{\text{LM}}(A) - \delta_k^{\text{LM}}(A) \neq \delta_k^{\text{LM}}(A) - \delta_{k-1}^{\text{LM}}(A)$ holds, which is less than $f(d, r)$. Here, $f(d, r)$ is upper bound of the maximum number of $n(\{a_k\})$ among $\{a_k\} \in X_d^r$, where $X_d^r \subseteq \mathbb{Z}^{r+1}$ is the set of sequences $\{a_k\}$ such that $a_0 = 0$, $a_r \geq 0$ and $a_{k+1} - a_k \leq a_k - a_{k-1} \leq d$ ($k = 1, \dots, r-1$), and $n(\{a_k\})$ is the number of k 's such that $a_{k+1} - a_k \neq a_k - a_{k-1}$. It can be shown by simple calculation that $f(d, r)$ plays the role of the upper bound. \square

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