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Bivariate Triangular Decompositions in the Presence of Asymptotes

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Abstract

Given two coprime polynomials P and Q in $\mathbb{Z}[x, y]$ of degree at most d and coefficients of bitsize at most τ , we address the problem of computing a triangular decomposition $\{(U_i(x), V_i(x, y))\}_{i \in \mathcal{I}}$ of the system $\{P, Q\}$.

The state-of-the-art worst-case bit complexity for computing such triangular decompositions when the curves defined by the input polynomials do not have common vertical asymptotes is in $\tilde{O}_B(d^6 + d^5\tau)$ [BLM⁺15, Proposition 16], where \tilde{O} refers to the complexity where polylogarithmic factors are omitted and O_B refers to the bit complexity.

We show that the same worst-case bit complexity can be achieved even when the curves defined by the input polynomials may have common vertical asymptotes. We actually present a refined bit complexity in $\tilde{O}_B(d_x^3 d_y^3 + (d_x^2 d_y^3 + d_x d_y^4)\tau)$ where d_x and d_y bound the degrees of P and Q in x and y , respectively. We also prove that the total bitsize of the decomposition is in $\tilde{O}((d_x^2 d_y^3 + d_x d_y^4)\tau)$.

1 Introduction

Computing triangular decompositions of algebraic systems is a well-known problem. In the special case of bivariate systems, a classical algorithm using subresultant sequences was first introduced by González-Vega and El Kahoui in the context of computing the topology of curves [GVEK96]. This algorithm is based on a direct consequence of the specialization property of subresultants and of the gap structure theorem, which implies the following (see Theorem 3): given two polynomials $P = \sum_{i=0}^p a_i(x)y^i$ and $Q = \sum_{i=0}^q b_i(x)y^i$ in $\mathbb{Z}[x, y]$ and $\alpha \in \mathbb{R}$ such that the leading coefficients $a_p(\alpha)$ and $b_q(\alpha)$ do not both vanish, then the first (with respect to increasing i) nonzero subresultant $\text{Sres}_{y,i}(P, Q)(\alpha, y)$ is of degree i and is equal to the gcd of $P(\alpha, y)$ and $Q(\alpha, y)$. Note that values α such that $a_p(\alpha)$ and $b_q(\alpha)$ both vanish are exactly the x -coordinates of the common vertical asymptotes of the curves defined by P and Q , which we refer to as the common vertical asymptotes of the polynomials, for simplicity. Hence, when P and Q do not have common vertical asymptotes, the gap structure theorem induces a decomposition of the system $\{P, Q\}$ into triangular subsystems $\{U_i(x), \text{Sres}_{y,i}(P, Q)(x, y)\}$ where the product of the U_i is the (squarefree part of the) resultant of P and Q with respect to y .

If the input polynomials have degree at most d and coefficients of bitsize at most τ , the worst-case bit complexity of this algorithm was initially analyzed in $\tilde{O}_B(d^{16} + d^{14}\tau^2)$ [GVEK96]. The complexity analysis was later improved to $\tilde{O}_B(d^7 + d^6\tau)$ [DET09, §4.2] and more recently to $\tilde{O}_B(d^6 + d^5\tau)$ by considering amortized bounds on the degrees and bitsizes of factors of the resultant [BLM⁺15,

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33 Proposition 16]. No better complexity is known for computing triangular decompositions, even
 34 in the expected Las-Vegas or Monte-Carlo settings and even in the absence of common vertical
 35 asymptotes.

36 In the general case when P and Q (may) admit common vertical asymptotes, the natural
 37 solution for computing a (full) triangular decomposition is to first use González-Vega and El Kahoui
 38 algorithm to compute the triangular decomposition of the solutions of $\{P, Q\}$ that do not lie on
 39 common vertical asymptotes (this can be done by removing from the resultant of P and Q the
 40 solutions corresponding to these asymptotes, i.e., $\gcd(a_p, b_q)$). Then, the triangular decomposition
 41 algorithm is called recursively on P and Q reduced modulo $\gcd(a_p, b_q)$. This natural approach was
 42 presented by Li et al. [LMMRS11].¹ The drawback of this approach is that the number of recursive
 43 calls may be linear in the minimum of the degrees in x and y of the input polynomials (it may
 44 happen that only one vertical asymptote is “handled” at each recursive call) and that the bitsize
 45 of the coefficients of the reduction of P and Q increases at each recursive call. However, Li et al.
 46 did not provide a complexity analysis of their algorithm.

47 We present here a simple variation on this natural algorithm where, instead of considering P
 48 and Q modulo $\gcd(a_p, b_q)$ at the first recursive call (and similarly for the other calls), we simply
 49 remove the leading terms $a_p y^p$ and $b_q y^q$ of P and Q . The number of recursive calls may still
 50 be linear in d but we show that, with this simple modification, the bit complexity of the overall
 51 recursive algorithm is the same as the bit complexity of the non-recursive algorithm (with no
 52 vertical asymptotes), that is $\tilde{O}_B(d^6 + d^5\tau)$. More precisely, we prove a worst-case bit complexity in
 53 $\tilde{O}_B(d_x^3 d_y^3 + (d_x^2 d_y^3 + d_x d_y^4)\tau)$ where d_x and d_y bound the degrees of P and Q in x and y , respectively.
 54 We also prove that the total bitsize of the decomposition is in $\tilde{O}((d_x^2 d_y^3 + d_x d_y^4)\tau)$. This implies in
 55 particular that, unless improving this upper bound, there is not much room for improving the bit
 56 complexity of the computation of the triangular decomposition. This also shows that, when there
 57 is a disparity between degrees d_x and d_y , the ordering of variables in the triangular decomposition
 58 impacts the complexity of the algorithm and of the output.

59 It is worthwhile to mention that in the general context of solving systems, one standard approach
 60 is to shear the coordinate system and to compute a triangular decomposition of the sheared system.
 61 This approach does not solve the given problem of computing a triangular decomposition of the
 62 input system since it computes a triangular decomposition of another system. Nevertheless, this
 63 approach, which naturally gets rid of vertical asymptotes, is theoretically straightforward, easy
 64 to implement, and its overall bit complexity is still in $\tilde{O}_B(d^6 + d^5\tau)$ (see e.g., [BLPR15, Lemma
 65 7]). However, it has the practical drawback that a shear $(x, y) \mapsto (x + ay, y)$ on polynomial
 66 $P = \sum_{i=0}^{d_y} a_i(x)y^i$ does not preserve its sparsity, increases the bitsize of its coefficients from τ up
 67 to $\tau + O(d_x + d_y)$ and increases its degree in y from d_y up to $d_x + d_y$. Since the bit complexity
 68 of the triangular decomposition algorithm is quartic in d_y (even in the absence of asymptotes; see
 69 Lemma 9), one should expect that shearing dramatically impacts the practical efficiency, which is
 70 observed in experiments. In addition, on a theoretical basis, if d_y is small compared to d_x then
 71 the overall bit complexity may drastically increase; for instance, in the extreme case where d_y is
 72 initially in $O(1)$, the overall worst-case complexity goes from $\tilde{O}_B(d_x^3 + d_x^2\tau)$ to $\tilde{O}_B(d_x^6 + d_x^5\tau)$. Still,
 73 it should be noted that, when complexities are expressed in terms of the total degree d , shearing
 74 leads to theoretically more efficient probabilistic algorithms for solving the system, both in the
 75 Las-Vegas setting [BLM⁺15] and in the Monte-Carlo setting [MS15].

¹Note that, in [LMMRS11], the reduction of P and Q modulo $\gcd(a_p, b_q)$ is understood from context because it is not clearly specified: P and Q are replaced by their “reductums” with respect to y but no definition of reductums is given. Note also that their algorithm misses the fact that, during the recursion, the reduced versions of P and Q may not define a zero-dimensional system and also that they may be both univariate.

76 In the next section, we recall standard definitions and results about multiplicities, subresultant
77 sequences, and gcds. We then present and analyze our triangular decomposition algorithm in
78 Section 3.

79 2 Notation and preliminaries

80 The bitsize of an integer p is the number of bits needed to represent it, that is $\lfloor \log p \rfloor + 1$ (log refers
81 to the logarithm in base 2). The bitsize of a polynomial with integer coefficients is the *maximum*
82 bitsize of its coefficients. As mentioned earlier, O_B refers to the bit complexity and \tilde{O} and \tilde{O}_B refer
83 to complexities where polylogarithmic factors are omitted. In this paper, most complexities are
84 expressed in terms of d_x and d_y , the maximum degrees in x and in y of $P, Q \in \mathbb{Z}[x, y]$, and in τ ,
85 their maximum bitsize. We also denote by d the maximum of d_x and d_y .

86 For any polynomial $P \in \mathbb{D}[x]$ where \mathbb{D} denotes a unique factorization domain, let $\text{Lc}_x(P)$ denote
87 its leading coefficient with respect to the variable x , $\text{deg}_x(P)$ its degree with respect to x (or simply
88 $\text{deg}(P)$ when P is univariate), and \bar{P} its squarefree part. A polynomial P that vanishes identically
89 is denoted by $P \equiv 0$. In the following, unless specified otherwise, the solutions of the considered
90 polynomials are always considered in the algebraic closure of the coefficient ring. Consequently,
91 a polynomial system is called zero-dimensional if its set of solutions over that algebraic closure is
92 finite.

93 In the rest of this section, we recall standard definitions and results about multiplicities, subre-
94 sultant sequences, and gcds.

95 **Multiplicities.** Geometrically, the notion of multiplicity of intersection of two regular curves is
96 intuitive. If the intersection is transverse, the multiplicity is one; otherwise, it is greater than one
97 and it measures the level of degeneracy of the tangential contact between the curves. Defining the
98 multiplicity of the intersection of two curves at a point that is singular for one of them (or possibly
99 both) is more involved and an abstract and general concept of multiplicity in an ideal is needed. We
100 recall this classical, though non-trivial, notion. We also introduce a simple notion of *multiplicity*
101 *in fibers* that will be output by our solver and that are relevant for the topology of a plane curve
102 (see e.g. [SW05]). Let \mathbb{F} be a field and $\bar{\mathbb{F}}$ be its algebraic closure.

103 **Definition 1.** Let I be an ideal of $\mathbb{F}[x, y]$. To each zero (α, β) of I corresponds a local ring
104 $(\bar{\mathbb{F}}[x, y]/I)_{(\alpha, \beta)}$ obtained by localizing the ring $\bar{\mathbb{F}}[x, y]/I$ at the maximal ideal $\langle x - \alpha, y - \beta \rangle$. When
105 this local ring is finite dimensional as $\bar{\mathbb{F}}$ -vector space, we say that (α, β) is an isolated zero of I and
106 this dimension is called the **multiplicity** of (α, β) as a zero of I [CLO05, §4.2].

107 We call the fiber of a point $p = (\alpha, \beta)$ the vertical line of equation $x = \alpha$. The **multiplicity**
108 **of \mathbf{p} in its fiber** with respect to a system of polynomials $\{P, Q\}$ in $\mathbb{F}[x, y]$ is the multiplicity of β
109 in the univariate polynomial $\text{gcd}(P(\alpha, y), Q(\alpha, y))$.² (This multiplicity is zero if P or Q does not
110 vanish at p .)

111 **Subresultant sequences.** We first recall the concept of *polynomial determinant* of a matrix
112 which is used in the definition of subresultants. Let M be an $m \times n$ matrix with $m \leq n$ and M_i
113 be the square submatrix of M consisting of the first $m - 1$ columns and the i -th column of M ,
114 for $i = m, \dots, n$. The *polynomial determinant* of M is the polynomial defined as $\det(M_m)y^{n-m} +$
115 $\det(M_{m+1})y^{n-(m+1)} + \dots + \det(M_n)$.

²The gcd is naturally considered over $\mathbb{F}(\alpha)[y]$, the ring of polynomials in y with coefficients in the field extension of \mathbb{F} by α .

116 Let $P = \sum_{i=0}^p a_i y^i$ and $Q = \sum_{i=0}^q b_i y^i$ be two polynomials in $\mathbb{D}[y]$ (where \mathbb{D} is a unique factoriza-
 117 tion domain such as $\mathbb{Q}[x]$) and assume without loss of generality that $p \geq q$. The Sylvester matrix
 118 of P and Q , $\text{Sylv}(P, Q)$ is the $(p+q)$ -square matrix whose rows are $y^{q-1}P, \dots, P, y^{p-1}Q, \dots, Q$
 119 considered as vectors in the basis $y^{p+q-1}, \dots, y, 1$.

120 **Definition 2.** ([EK03, §3]). For $i = 0, \dots, \min(q, p-1)$, let $\text{Sylv}_i(P, Q)$ be the $(p+q-2i) \times (p+q-i)$
 121 matrix obtained from $\text{Sylv}(P, Q)$ by deleting the i last rows of the coefficients of P , the i last rows
 122 of the coefficients of Q , and the i last columns.

123 For $i = 0, \dots, \min(q, p-1)$, the i -th polynomial subresultant of P and Q , denoted by $\text{Sres}_{y,i}(P, Q)$
 124 is the polynomial determinant of $\text{Sylv}_i(P, Q)$.

125 For practical consideration, when $q = p$, we define the q -th polynomial subresultant of P and
 126 Q as Q .³ $\text{Sres}_{y,i}(P, Q)$ has degree at most i in y , and the coefficient of its monomial of degree i in
 127 y , denoted by $\text{sres}_{y,i}(P, Q)$ or sres_i , is called the i -th *principal subresultant coefficient*. Note that
 128 $\text{Sres}_{y,0}(P, Q) = \text{sres}_{y,0}(P, Q)$ is the *resultant* of P and Q with respect to y , which we also denote
 129 by $\text{Res}_y(P, Q)$. Note also that the subresultants of P and Q are equal to either 0 or to polynomials
 130 in the remainder sequence of P and Q in Euclid's algorithm (up to multiplicative factors in \mathbb{D})
 131 [BPR06, §8.3.3 & Cor. 8.32].

132 Consider now two bivariate polynomials with coefficients in $\mathbb{D} = \mathbb{Z}$: $P = \sum_{i=0}^p a_i(x)y^i$ and
 133 $Q = \sum_{i=0}^q b_i(x)y^i$ with $p \geq q$. The following fundamental property of subresultant sequences is
 134 instrumental in the triangular decomposition algorithms. Note that this result is often stated with
 135 the stronger assumption that *none* of the leading terms $a_p(\alpha)$ and $b_q(\alpha)$ vanish. This property is a
 136 direct consequence of the specialization property of subresultants and of the gap structure theorem;
 137 see [EK03, Lemmas 2.3, 3.1 and Corollary 5.1] for a proof.

138 **Theorem 3.** For any α such that $a_p(\alpha)$ and $b_q(\alpha)$ do not both vanish, the first $\text{Sres}_{y,k}(P, Q)(\alpha, y)$
 139 (for k increasing) that does not identically vanish is of degree k and it is the gcd of $P(\alpha, y)$ and
 140 $Q(\alpha, y)$ (up to a nonzero constant in the fraction field of $\mathbb{D}(\alpha)$).

141 **Lemma 4** ([BPR06, Prop. 8.46] [Rei97, §8] [vzGG13, §11.2]). Let P and Q be in $\mathbb{Z}[x_1, \dots, x_n][y]$
 142 (n fixed) with coefficients of bitsize at most τ such that their degrees in y are bounded by d_y and
 143 their degrees in the other variables are bounded by d_x .

- 144 • The coefficients of $\text{Sres}_{y,i}(P, Q)$ have bitsize in $\tilde{O}(d_y \tau)$.
- 145 • The degree in x_j of $\text{Sres}_{y,i}(P, Q)$ is at most $2d_x(d_y - i)$.
- 146 • For any $i \in \{0, \dots, \min(\deg_y(P), \deg_y(Q))\}$, $\text{Sres}_{y,i}(P, Q)$ can be computed in $\tilde{O}(d_x^n d_y^{n+1})$
 147 arithmetic operations and $\tilde{O}_B(d_x^n d_y^{n+2} \tau)$ bit operations. These complexities also hold for the
 148 computation of the sequence of principal subresultant coefficients $\text{sres}_i(P, Q)$.⁴

149 **Gcds.** We often consider the gcd of two univariate polynomials P and Q in $\mathbb{Z}[x]$ and the gcd-free
 150 part of P with respect to Q , that is, $P/\text{gcd}(P, Q)$. Note that, when $Q = P'$, the latter is the
 151 squarefree part \bar{P} . When P and Q have degree at most d and bitsize at most τ , their gcd and gcd-
 152 free parts can be computed with a bit complexity in $\tilde{O}_B(d^2 \tau)$ [BPR06, Remark 10.19]. However,
 153 we will need a finer complexity in the case of two polynomials with different degrees and bitsizes.

³ It can be observed that, when $p > q$, the q -th subresultant is equal to $b_q^{p-q-1}Q$, however it is not defined when
 $p = q$. In this case, El Kahoui suggests to extend the definition to $b_q^{-1}Q$ assuming that the domain \mathbb{D} is integral.
 However, b_q^{-1} does not necessarily belong to \mathbb{D} , which is not practical. Note that it is important to define the q -th
 subresultant to be a multiple of Q so that Theorem 3 holds when $P(\alpha, y)$ and $Q(\alpha, y)$ have same degree and are
 multiple of one another.

⁴The complexity of computing the sequence of principal subresultant coefficients is stated in [vzGG13, §. 11.2]
 only for univariate polynomials, however, one can use the binary segmentation technique described in [Rei97, §8] to
 generalize the latter to multivariate polynomials.

154 **Lemma 5** ([LR01]⁵). *Let P and Q be two polynomials in $\mathbb{Z}[x]$ of degrees p and q and of bitsizes τ_P
155 and τ_Q , respectively. A gcd of P and Q of bitsize $O(\min(p+\tau_P, q+\tau_Q))$ in $\mathbb{Z}[x]$, can be computed in
156 $\tilde{O}_B(\max(p, q)(p\tau_Q + q\tau_P))$ bit operations. A gcd-free part of P with respect to Q , of bitsize $O(p+\tau_P)$
157 in $\mathbb{Z}[x]$, can be computed in the same bit complexity.*

158 3 Triangular decomposition

159 We present here our algorithm that decomposes a zero-dimensional system $\{P, Q\}$ of polynomials in
160 $\mathbb{Z}[x, y]$ into a set of regular triangular systems of the form $\{U(x), V(x, y)\}$. Recall that such a system
161 is said regular if U and $\text{Lc}_y(V)$ are coprime. Algorithm 2 is the main algorithm, which recursively
162 calls Algorithm 1, the latter being in essence that of Gonzalez-Vega and El Kahoui [GVEK96]. For
163 clarity and completeness, we briefly describe this latter algorithm with an emphasis on the main
164 differences with that of Gonzalez-Vega and El Kahoui, and give a succinct proof of correctness. We
165 then describe Algorithm 2, prove its correctness in Lemmas 6 and 7 and analyze its complexity in
166 Proposition 10.

167 **Algorithm 1: Triangular decomposition of $\{P, Q\}$ away from their common vertical
168 asymptotes and such that A vanishes.** Algorithm 1 takes as input $P, Q \in \mathbb{Z}[x, y]$ and a
169 univariate polynomial $A \in \mathbb{Z}[x]$ such that system $\{P, Q, A\}$ is zero dimensional and it computes a
170 set of triangular systems whose solutions are the solutions of $\{P, Q, A\}$ that do not lie on a common
171 vertical asymptote of the curves defined by P and Q . Considering the calls to Algorithm 1 made
172 by Algorithm 2, Algorithm 1 will first run with $A \equiv 0$ and compute a triangular decomposition of
173 the solutions away from the common vertical asymptotes of P and Q ; then Algorithm 1 will be
174 called with A encoding a subset of these common vertical asymptotes and two polynomials that
175 coincide with P and Q on these asymptotes. Algorithm 1 is essentially that of Gonzalez-Vega and
176 El Kahoui [GVEK96] in the case where $\{P, Q\}$ is zero dimensional, P and Q do not have any
177 common vertical asymptote, and $A \equiv 0$.

178 The projection onto the x -axis of the solutions of system $\{P, Q\}$ that do not lie on a common
179 vertical asymptote of the curves defined by P and Q are exactly the roots of the resultant of P and
180 Q with respect to y divided by the gcd of the leading coefficients of P and Q with respect to y . We
181 actually consider the squarefree parts of these polynomials, $\overline{\text{Res}_y(P, Q)}$ and $\overline{\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))}$,
182 which is critical for our property on the multiplicity of the solutions in their fibers (Lemma 7).
183 In order to restrict the set of solutions of $\{P, Q\}$ that do not lie on a common vertical asymptote
184 to those where A vanishes, we consider the gcd of $\frac{\overline{\text{Res}_y(P, Q)}}{\overline{\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))}}$ with A . However, this does
185 not work when $\text{Res}_y(P, Q) \equiv 0$, that is when $\{P, Q\}$ is not zero dimensional (and in generic
186 position). We thus consider instead $F = \frac{\overline{\text{Res}_y(P, Q)}}{\overline{\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))}}$ when $A \equiv 0$ and, otherwise, $F =$
187 $\frac{\overline{\text{gcd}(\text{Res}_y(P, Q), A)}}{\overline{\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q), A)}}$, which is equal to $\frac{\overline{\text{gcd}(\text{Res}_y(P, Q), A)}}{\overline{\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q), A)}}$ since A is squarefree. Then, the roots
188 of F are the projections (on x) of the solutions of $\{P, Q, A\}$ that are not on the common vertical
189 asymptotes of P and Q .

⁵The algorithm in [LR01] uses the well-known half-gcd approach to compute any polynomial in the Sylvester-Habicht and cofactors sequence in a softly-linear number of arithmetic operations, and it exploits Hadamard's bound on determinants to bound the size of intermediate coefficients. When the two input polynomials have different degrees and bitsizes, Hadamard's bound reads as $\tilde{O}(p\tau_Q + q\tau_P)$ instead of simply $\tilde{O}(d\tau)$ and, similarly as in Lemma 5, the algorithm in [LR01] yields a gcd and gcd-free parts of P and Q in $\tilde{O}_B(\max(p, q)(p\tau_Q + q\tau_P))$ bit operations. Furthermore, the gcd and gcd-free parts computed this way are in $\mathbb{Z}[x]$ with coefficients of bitsize $\tilde{O}(p\tau_Q + q\tau_P)$, thus, dividing them by the gcd of their coefficients can be done with $\tilde{O}_B(\max(p, q)(p\tau_Q + q\tau_P))$ bit operations and yields a gcd and gcd-free parts in $\mathbb{Z}[x]$ with minimal bitsize, which is as claimed by Mignotte's bound; see e.g. [BPR06, Corollary 10.12].

190 Algorithm 1 decomposes F into factors according to Theorem 3. Recall that $\text{sres}_{y,i}(P, Q)$ denotes
 191 the coefficient of the monomial of degree i in y of $\text{Sres}_{y,i}(P, Q)$, the i -th polynomial subresultant
 192 of P and Q with respect to y . Polynomial F is decomposed into factors F_i , $i = 1, 2, \dots$, such that
 193 for any root α of F_i , $\text{sres}_{y,i}(P, Q)(\alpha)$ is the first (for i increasing) non-vanishing coefficient. The
 194 algorithm then returns the set of non-trivial triangular systems $\{F_i, \text{Sres}_{y,i}(P, Q)\}$ whose solutions
 195 are, by Theorem 3, those of $\{P, Q, A\}$ that are not on the common vertical asymptotes of P and
 196 Q . The triangular systems are regular by construction.

197 **Algorithm 2: Complete triangular decomposition of $\{P, Q\}$.** Algorithm 2 takes as input a
 198 zero-dimensional system $\{P, Q\}$ in $\mathbb{Z}[x, y]$ and computes a set of regular triangular systems whose
 199 solutions are those of $\{P, Q\}$. Algorithm 2 calls Algorithm 1 recursively, first with the input poly-
 200 nomials $P_1 = P$, $Q_1 = Q$ and $A_1 \equiv 0$, and then, for $h \geq 2$, with $P_h = P_{h-1} - \text{Lc}_y(P_{h-1})y^{\deg_y(P_{h-1})}$,
 201 $Q_h = Q_{h-1} - \text{Lc}_y(Q_{h-1})y^{\deg_y(Q_{h-1})}$ and $A_h \in \mathbb{Z}[x]$ that vanishes exactly on the common vertical
 202 asymptotes of $P_1, Q_1, \dots, P_{h-1}, Q_{h-1}$ that are not common vertical asymptotes of P_h and Q_h .

203 **Lemma 6.** *Given P, Q in $\mathbb{Z}[x, y]$ defining a zero-dimensional system, Algorithm 2 outputs a set of*
 204 *regular triangular systems, each of the form $\{U(x), V(x, y)\}$ with coefficients in \mathbb{Z} , whose sets of*
 205 *solutions are disjoint and are exactly those of $\{P, Q\}$.*

206 *Proof.* Let P_h, Q_h, A_h and B_h be the polynomials P, Q, A and B defined in Algorithm 2 when
 207 Algorithm 1 is called for the h -th time (which might be different from the h -th iteration of the
 208 loop). We have $P_1 = P$, $Q_1 = Q$, $A_1 \equiv 0$ and $B_1 = \text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))$, thus the first call
 209 to Algorithm 1 returns triangular systems encoding the solutions of $\{P, Q\}$ that are not over
 210 the common vertical asymptotes of P and Q . For $h \geq 1$, B_h encodes the common vertical
 211 asymptotes of $P_1, Q_1, \dots, P_h, Q_h$ and, for $h \geq 2$, A_h encodes the common vertical asymptotes
 212 of $P_1, Q_1, \dots, P_{h-1}, Q_{h-1}$ that are not common vertical asymptotes of P_h and Q_h .

213 Thus P_h coincides with P on the vertical asymptotes encoded by A_h , and similarly for Q_h . This
 214 first implies that $\{P_h, Q_h, A_h\}$ is zero-dimensional, since $\{P, Q\}$ is. Furthermore, A_h is squarefree
 215 because it is either identically equal to 0 (when $h = 1$) or it divides B_{h-1} , which divides $B_1 =$
 216 $\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))$. Hence $\{P_h, Q_h, A_h\}$ satisfies the requirements of Algorithm 1.

217 Algorithm 1 when called on P_h, Q_h, A_h returns a set of regular triangular systems whose solu-
 218 tions are those of $\{P_h, Q_h, A_h\}$ away from the common asymptotes of P_h and Q_h . But, for $h \geq 2$,
 219 A_h does not vanish on these asymptotes so the solutions are those of $\{P_h, Q_h, A_h\}$. Furthermore, P_h
 220 and Q_h coincide with P and Q when A_h vanishes, thus these solutions are also those of $\{P, Q, A_h\}$.
 221 Finally, the above property on A_h also implies that the A_h , for $h \geq 2$, are coprime and that their
 222 product encodes the common asymptotes of P and Q . Thus, the set of systems returned by all the
 223 calls to Algorithm 1 except the first one have sets of solutions that are disjoint and are the solutions
 224 of $\{P, Q\}$ that lie on their common asymptotes. This concludes the proof since the systems output
 225 by the first call to Algorithm 1 are those of $\{P, Q\}$ away from these asymptotes. \square

226 We now prove that Algorithm 2 preserves the multiplicities of the solutions, in the following
 227 sense (see Definition 1).

228 **Lemma 7.** *The multiplicity of any solution in the triangular systems output by Algorithm 2 is its*
 229 *multiplicity in its fiber with respect to the system $\{P, Q\}$.*

230 *Proof.* Consider a solution (α, β) of a triangular system $\{U(x), V(x, y)\}$ output by Algorithm 2.
 231 This triangular system is output by Algorithm 1 called on some polynomials P_h, Q_h, A_h at the
 232 h -th call of Algorithm 1. By construction, $U(x)$ is squarefree because, in Algorithm 1, F_i divides

Algorithm 1 Triangular decomposition away from asymptotes

Input: P, Q in $\mathbb{Z}[x, y]$ and A squarefree in $\mathbb{Z}[x]$ such that system $\{P, Q, A\}$ is zero-dimensional.

Output: A set of regular triangular systems, each of the form $\{U(x), V(x, y)\}$ with coefficients in \mathbb{Z} , whose solutions are those of $\{P, Q, A\}$ that do not lie on a common vertical asymptote of P and Q .

1. **if** $A \equiv 0$ **then**
 2. $R(x) = \overline{\text{Res}_y(P, Q)}$, $B(x) = \overline{\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))}$, $F = R/B$
 3. **else**
 4. $R(x) = \text{Res}_y(P, Q)$, $B(x) = \text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))$, $F = \frac{\text{gcd}(R, A)}{\text{gcd}(B, A)}$
 5. **if** neither P nor Q is in $\mathbb{Z}[x]$ **then**
 6. If needed, exchange P and Q so that $\deg_y(Q) \leq \deg_y(P)$
 7. Compute $\{\text{sres}_{y,i}(P, Q)\}_{i=0, \dots, \deg_y(Q)}$, the principal subresultant sequence of P and Q w.r.t. y
 8. $G_0 = F$, $\mathcal{TD} = \emptyset$
 9. **for** $i = 1$ **to** $\deg_y(Q)$ **do**
 10. $G_i = \text{gcd}(G_{i-1}, \text{sres}_{y,i}(P, Q))$
 11. $F_i = G_{i-1}/G_i$
 12. **if** $\deg_x(F_i) > 0$ **then**
 13. Compute $\text{Sres}_{y,i}(P, Q)$
 14. $\mathcal{TD} = \mathcal{TD} \cup \{F_i, \text{Sres}_{y,i}(P, Q)\}$
 15. **return** \mathcal{TD}
 16. **else if** P and Q are in $\mathbb{Z}[x]$ **then**
 17. **return** \emptyset
 18. **else** {Assume wlog that P is in $\mathbb{Z}[x]$ (and Q is not)}
 19. **return** $\{F, Q\}$
-

Algorithm 2 Complete triangular decomposition

Input: P, Q in $\mathbb{Z}[x, y]$ defining a zero-dimensional system.

Output: A set of regular triangular systems, each of the form $\{U(x), V(x, y)\}$ with coefficients in \mathbb{Z} , whose sets of solutions are disjoint and are exactly those of $\{P, Q\}$. The multiplicity of any solution in these triangular systems is the multiplicity of the solution in its fiber with respect to the system $\{P, Q\}$ (see Definition 1).

1. $A = 0$, $B = \text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))$, $\mathcal{TD} = \emptyset$
 2. **repeat**
 3. **if**⁶ $\deg_x(A) \neq 0$ **then**
 4. $\mathcal{TD} = \mathcal{TD} \cup \text{Algorithm 1}(P, Q, A)$
 5. $P = P - \text{Lc}_y(P)y^{\deg_y(P)}$, $Q = Q - \text{Lc}_y(Q)y^{\deg_y(Q)}$
 6. $B_{\text{new}} = \text{gcd}(B, \text{Lc}_y(P), \text{Lc}_y(Q))$
 7. $A = \frac{B}{B_{\text{new}}}$
 8. **until** $\deg(B) = 0$
 9. **return** \mathcal{TD}
-

²³³ F , which is squarefree; indeed the first time Algorithm 1 is called F is squarefree by definition
²³⁴ (Line 2) and, in the other calls, F divides A , which divides B , which divides $\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))$

⁶Using the convention that the degree of the null polynomial is $-\infty$.

235 (see Algorithm 2). Thus, the multiplicity of (α, β) in $\{U(x), V(x, y)\}$ is the multiplicity of β in the
 236 univariate polynomial $V(\alpha, y)$. The bivariate polynomial V is defined either as P_h or Q_h (Line 19)
 237 or as $\text{Sres}_{y,i}(P_h, Q_h)$ (Line 14).

238 In the latter case, $V(\alpha, y) = \text{Sres}_{y,i}(P_h, Q_h)(\alpha, y)$ is equal to the gcd of $P_h(\alpha, y)$ and $Q_h(\alpha, y)$ by
 239 Theorem 3. By construction, if $\{U(x), V(x, y)\}$ is output by the h -th call of Algorithm 1, then the
 240 $h - 1$ first (non-zero) coefficients of P and Q (seen as polynomials in y) vanish at α . In other words,
 241 $P_h(\alpha, y) = P(\alpha, y)$ and similarly for Q_h . Thus, the multiplicity of β in $V(\alpha, y)$ is the multiplicity of
 242 β in $\text{gcd}(P(\alpha, y), Q(\alpha, y))$, which is by definition the multiplicity of (α, β) in its fiber with respect
 243 to $\{P, Q\}$.

244 In the former case, if say $V = Q_h$ then $P_h \in \mathbb{Z}[x]$ and $P_h(\alpha) = 0$. The gcd of $P_h(\alpha)$ and
 245 $Q_h(\alpha, y)$ is thus $Q_h(\alpha, y)$. The multiplicity of β in $V(\alpha, y) = Q_h(\alpha, y)$ is thus its multiplicity
 246 in $\text{gcd}(P_h(\alpha), Q_h(\alpha, y))$ which is equal to $\text{gcd}(P(\alpha, y), Q(\alpha, y))$, as above. Hence, as above, the
 247 multiplicity of β in $V(\alpha, y)$ is the multiplicity of (α, β) in its fiber with respect to $\{P, Q\}$, which
 248 concludes the proof. \square

249 We now analyze the complexity of Algorithms 2 and start by two preliminary lemmas, which
 250 are direct generalizations of Propositions 15 and 16 in [BLM⁺15] but expressed in terms of d_x and
 251 d_y instead of the total degree.

252 **Lemma 8.** *For $i = 0, \dots, \deg_y(Q) - 1$, let d_i and τ_i be the degree and bitsize of the polynomial G_i
 253 in the triangular decomposition of P and Q computed in Algorithm 1 with $A \equiv 0$. We have:*

- 254 • $d_i \leq \frac{d_x d_y}{i+1}$ and $\tau_i = \tilde{O}\left(\frac{d_x d_y + d_y \tau}{i+1}\right)$,
- 255 • $\sum_{i=0}^{\deg_y(Q)-1} d_i \leq d_x d_y$ and $\sum_{i=0}^{\deg_y(Q)-1} \tau_i = \tilde{O}(d_x d_y + d_y \tau)$.

256 *Proof.* Bouzidi et al. [BLM⁺15, Prop. 15] proved the above bounds with d^2 in place of $d_x d_y$ and $d\tau$
 257 in place of $d_y \tau$. There, d^2 and $d\tau$ refer to the bounds on the degree and the bitsize of $\text{Res}_y(P, Q)$.
 258 The degree and bitsize of this resultant can also be expressed as $O(d_x d_y)$ and $\tilde{O}(d_y \tau)$ (by Lemma 4)
 259 and literally replacing in [BLM⁺15, Prop. 15] the bound $O(d^2)$ on the degree of the resultant by
 260 $O(d_x d_y)$ and the bound $\tilde{O}(d\tau)$ on its bitsize by $\tilde{O}(d_y \tau)$ directly yields the result. \square

261 The following lemma is a direct and straightforward generalization of [BLM⁺15, Prop. 16],
 262 which proves a bit complexity of $\tilde{O}_B(d^6 + d^5 \tau)$ for Algorithm 1 with $A \equiv 0$.⁷

263 **Lemma 9.** *If P, Q in $\mathbb{Z}[x, y]$ have degree at most d_x in x , d_y in y , and bitsize at most τ , Algorithm 1
 264 with $A \equiv 0$ performs $\tilde{O}_B(d_x^3 d_y^3 + (d_x^2 d_y^3 + d_x d_y^4) \tau)$ bit operations in the worst case.*

265 *Proof.* By Lemma 4, each of the principal subresultant coefficients $\text{sres}_{y,i}$ (including the resultant)
 266 has degree $O(d_x d_y)$ and bitsize $\tilde{O}(d_y \tau)$. Thus, in Line 2, by Lemma 5, the squarefree part of the
 267 resultant can be computed in $\tilde{O}_B((d_x d_y)^2 (d_y \tau)) = \tilde{O}_B(d_x^2 d_y^3 \tau)$ bit operations and its bitsize is in
 268 $O(d_x d_y + d_y \tau) = O(d_y (d_x + \tau))$. In the same line, still by Lemma 5, $\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))$ has
 269 bitsize $O(d_x + \tau)$ and it can be computed in $\tilde{O}_B(d_x^2 \tau)$ bit operations; its squarefree part can thus be
 270 computed in $\tilde{O}_B(d_x^2 (d_x + \tau))$ bit operations and its bitsize is still in $O(d_x + \tau)$. Still in Line 2, the
 271 exact division R/B , which is a gcd-free computation, can be done with $\tilde{O}_B((d_x d_y)^2 (d_y (d_x + \tau))) =$
 272 $\tilde{O}_B(d_x^2 d_y^3 (d_x + \tau))$ bit operations.

⁷Note that there is nonetheless a minor difference between Algorithm 1 (with $A \equiv 0$) and the one analyzed
 in [BLM⁺15, Prop. 16], which is that in the former we consider $F = \text{Res}_y(P, Q) / \text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))$ instead of
 $\text{Res}_y(P, Q)$ with the assumption that $\text{Lc}_y(P)$ and $\text{Lc}_y(Q)$ are coprime in the latter. However, this has no impact on
 the complexity because, by Mignotte's lemma, F has degree $O(d^2)$ and bitsize $\tilde{O}(d^2 + d\tau)$ as $\overline{\text{Res}_y(P, Q)}$.

273 By Lemma 4, the sequence of the subresultants $\text{Sres}_{y,i}(P, Q)$ can be computed in $\tilde{O}_B(d_x d_y^4 \tau)$
 274 bit operations and the sequence of their principal coefficients $\text{sres}_i(P, Q)$ (including the resultant)
 275 can be computed in $\tilde{O}_B(d_x d^3 \tau)$ bit operations. Thus, the overall bit complexity of Lines 7 and 13
 276 is $\tilde{O}_B(d_x d_y^4 \tau)$.

277 Line 10 performs, in total, d_y gcd computations between polynomials G_{i-1} and $\text{sres}_{y,i}$. Polyno-
 278 mial $\text{sres}_{y,i}$ has bitsize $\tilde{O}(d_y \tau)$ and degree $O(d_x d_y)$, and denoting by τ_i and d_i the bitsize and degree
 279 of G_i , Lemma 5 yields a complexity in $\tilde{O}_B((d_x d_y)((d_x d_y)\tau_{i-1} + d_{i-1} d_y \tau))$ for the computation of
 280 G_i . According to Lemma 8, these complexities sum up over all i to $\tilde{O}_B((d_x d_y)^2 (d_x d_y + d_y \tau))$.

281 Finally, in Line 11, by Lemma 5, the exact division of G_{i-1} by G_i can be done with a bit
 282 complexity $O_B(d_i^2 \tau_i)$. Since $d_i \leq d_x d_y$ by Lemma 8, $\sum_i O_B(d_i^2 \tau_i) = \tilde{O}_B((d_x d_y)^2 (d_x d_y + d_y \tau))$.

283 Hence, the overall bit complexity of the algorithm is in $\tilde{O}_B(d_x^3 d_y^3 + (d_x^2 d_y^3 + d_x d_y^4) \tau)$. \square

284 **Proposition 10.** *Let P, Q in $\mathbb{Z}[x, y]$ be two polynomials of degrees at most d_x and d_y in x and y ,
 285 with coefficients of bitsize at most τ , and defining a zero-dimensional system. With $d = \max(d_x, d_y)$,
 286 Algorithm 2 computes a triangular decomposition of $\{P, Q\}$ using $\tilde{O}_B(d^6 + d^5 \tau)$ bit operations in
 287 the worst case. In terms of d_x and d_y , this complexity is $\tilde{O}_B(d_x^3 d_y^3 + (d_x^2 d_y^3 + d_x d_y^4) \tau)$. Moreover,
 288 the total bitsize of the decomposition is in $\tilde{O}((d_x^2 d_y^3 + d_x d_y^4) \tau)$.*

289 *Proof.* The number of iterations of the loop in Algorithm 2 is at most $d_y + 1$. Beside the calls
 290 to Algorithm 1, Algorithm 2 thus performs $O(d_y)$ gcd operations and exact divisions of univariate
 291 polynomials. The degree of these polynomials is trivially at most d_x and their bitsizes are in
 292 $O(d_x + \tau)$ by Mignotte's lemma [BPR06, Corollary 10.12] because the gcds always divide some
 293 coefficients of P (and Q) seen in $\mathbb{Z}[x][y]$. Thus, by Lemma 5, the bit complexity of each of the
 294 gcd and exact division (i.e., a gcd-free) computations is in $\tilde{O}_B(d_x^2 (d_x + \tau))$, which yields a total bit
 295 complexity in $\tilde{O}_B(d_y d_x^2 (d_x + \tau))$.

296 We now analyze the complexity of the calls to Algorithm 1. Denote by $P_h, Q_h, A_h, F_h, F_{h,i}, G_{h,i}$
 297 the instances of P, Q, A, F, F_i, G_i in the h -th call to Algorithm 1. Since Algorithm 1 is called only
 298 if $\deg_x(A) \neq 0$, we have that $\deg_x(A_{h>1}) > 0$. It follows that h varies from 1 to at most d_x because
 299 $\prod_{h>1} A_h$ encodes the common vertical asymptotes of P and Q (as noted in the proof of Lemma 6)
 300 and there are at most d_x such asymptotes.

301 By Lemma 9, the first call to Algorithm 1 with $A_1 \equiv 0$ has bit complexity $\tilde{O}_B(d_x^3 d_y^3 + (d_x^2 d_y^3 +$
 302 $d_x d_y^4) \tau)$.

303 In the rest of the proof, we consider the calls to Algorithm 1 except for the first one. In all these
 304 calls, the polynomials $F_{h,i}$ are pairwise coprime by construction and their product encodes a subset
 305 of the common vertical asymptotes of the initial input polynomials P and Q (i.e. $\prod_{h,i} F_{h,i}$ divides
 306 $\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))$). Hence, in Line 13, at most d_x subresultant polynomials $\text{Sres}_{y,i}(P_h, Q_h)$ are
 307 computed over all calls (but the first one). Since P_h and Q_h are truncated versions of P and Q ,
 308 their degrees and bitsize are still bounded by d_x, d_y and τ , hence all subresultant polynomials can
 309 be computed in a total bit complexity of $\tilde{O}_B(d_x^2 d_y^3 \tau)$, by Lemma 4.

310 In Line 4, by Lemma 4, the resultant R_h of P_h and Q_h can be computed with $\tilde{O}_B(d_x d_y^3 \tau)$ bit op-
 311 erations and its degree and bitsize are in $O(d_x d_y)$ and $\tilde{O}(d_y \tau)$, respectively. By Lemma 5, the gcd B_h
 312 of $\text{Lc}_y(P_h)$ and $\text{Lc}_y(Q_h)$ can be computed in $\tilde{O}_B(d_x^2 \tau)$ bit operations and its degree and bitsize are in
 313 $O(d_x)$ and $O(d_x + \tau)$, respectively. On the other hand, A_h divides $\text{gcd}(\text{Lc}_y(P), \text{Lc}_y(Q))$ thus its de-
 314 gree and bitsize are in $O(d_x)$ and $O(d_x + \tau)$, respectively, by Mignotte's lemma. Thus, by Lemma 5,
 315 $\text{gcd}(R_h, A_h)$ and $\text{gcd}(B_h, A_h)$ can be computed in $\tilde{O}_B(\max(d_x d_y, d_x)((d_x d_y)(d_x + \tau) + d_x(d_y \tau))) =$
 316 $\tilde{O}_B((d_x d_y)^2 (d_x + \tau))$ bit operations and their degree and bitsize are in $O(d_x)$ and $O(d_x + \tau)$, respec-
 317 tively. Furthermore, the same lemma yields that the exact division $\text{gcd}(R_h, A_h) / \text{gcd}(B_h, A_h)$, that

318 is the gcd-free part of $\gcd(R_h, A_h)$ with respect to $\gcd(B_h, A_h)$, can be computed in $\tilde{O}_B(d_x^2(d_x + \tau))$
319 bit operations. One iteration of Line 4 thus has a bit complexity in $\tilde{O}_B((d_x d_y)^2(d_x + \tau))$, which
320 yields a bit complexity in $\tilde{O}_B(d_y(d_x d_y)^2(d_x + \tau))$ for all calls (but the first one).

321 Similarly, in Line 11, one division $G_{h,i-1}/G_{h,i}$ has complexity $\tilde{O}_B(d_x^2(d_x + \tau))$. Indeed $G_{h,i}$
322 divides $G_{h,0} = F_h$, which divides $\text{Lc}_y(P)$, which has degree at most d_x and bitsize at most τ . Thus,
323 $G_{h,i}$ has degree at most d_x and bitsize $O(d_x + \tau)$ by Mignotte's lemma, which yields the complexity
324 bound $\tilde{O}_B(d_x^2(d_x + \tau))$. There are $O(d_y^2)$ calls to Line 11 and thus the total bit complexity of that
325 line is in $\tilde{O}_B((d_x d_y)^2(d_x + \tau))$.⁸

326 In Line 7, for every call, the bit complexity of the computation of the principal subresultant
327 sequence is in $\tilde{O}_B(d_x d_y^3 \tau)$ by Lemma 4, yielding a complexity in $\tilde{O}_B(d_x^2 d_y^3 \tau)$ for all the $O(d_x)$ calls
328 to Algorithm 1.

329 We finally analyze the complexity of Line 10 where $G_{h,i} = \gcd(G_{h,i-1}, \text{sres}_{y,i}(P_h, Q_h))$ is com-
330 puted. For that purpose, we need to amortize the sum of the degrees $d_{h,i}$ and the sum of the bitsizes
331 $\tau_{h,i}$ of $G_{h,i}$ over h . For the degree, it is straightforward that $\sum_h d_{h,i} \leq d_x$, for any i , because, as noted
332 above, $G_{h,i}$ divides $G_{h,0} = F_h$, thus $\prod_h G_{h,i}$ divides $\prod_h F_h$, which divides $\gcd(\text{Lc}_y(P), \text{Lc}_y(Q))$.

333 We now prove that $\sum_h \tau_{h,i} = O(d_x + \tau)$, for any i , using Mahler's measure, as in [BLM⁺15,
334 Prop. 15]. For a univariate polynomial f with integer coefficients, its Mahler measure is $M(f) =$
335 $|\text{Lc}(f)| \prod_{z_i \text{ s.t. } f(z_i)=0} \max(1, |z_i|)$, where every complex root appears with its multiplicity. Mahler's
336 measure is multiplicative: $M(fg) = M(f)M(g)$ and, since it is at least 1 for any polynomial with
337 integer coefficients, f divides h (i.e., $h = fg$) implies that $M(h) \geq M(f)$. We also have the following
338 two inequalities connecting the bitsize τ and degree d of f and its Mahler measure $M(f)$.

- 339 (i) $\tau \leq 1 + d + \log M(f)$. Indeed, [BPR06, Prop. 10.8] states that $\|f\|_1 \leq 2^d M(f)$, thus $\|f\|_\infty \leq$
340 $2^d M(f)$ and $\log \|f\|_\infty \leq d + \log M(f)$, which yields the result since $\tau = \lfloor \log \|f\|_\infty \rfloor + 1$.
- 341 (ii) $\log M(f) = O(\tau + \log d)$. Indeed, [BPR06, Prop. 10.9] states that $M(f) \leq \|f\|_2$, thus
342 $M(f) \leq \sqrt{d+1} \|f\|_\infty$ and $\log M(f) \leq \log \sqrt{d+1} + \log \|f\|_\infty$.

343 By Inequality (i),

$$\sum_h \tau_{h,i} \leq d_x + \sum_h d_{h,i} + \log M\left(\prod_h G_{h,i}\right).$$

344 As noted above, $\sum_h d_{h,i} \leq d_x$ and $\prod_h G_{h,i}$ divides $\text{Lc}_y(P)$, thus $M(\prod_h G_{h,i}) \leq M(\text{Lc}_y(P))$ and by
345 Inequality (ii), $\log M(\prod_h G_{h,i}) \leq \log M(\text{Lc}_y(P)) = O(\tau + \log d_x)$. Hence, $\sum_h \tau_{h,i} = O(d_x + \tau)$.

346 Now, since $\text{sres}_{y,i}(P_h, Q_h)$ has degree $O(d_x d_y)$ and bitsize $\tilde{O}(d_y \tau)$ by Lemma 4, computing $G_{h,i} =$
347 $\gcd(G_{h,i-1}, \text{sres}_{y,i}(P_h, Q_h))$ has bit complexity $\tilde{O}_B(d_x d_y (d_x d_y \tau_{h,i-1} + d_{h,i-1} d_y \tau))$ by Lemma 5. Sum-
348 ming over h gives $\tilde{O}_B((d_x d_y)^2(d_x + \tau))$ and summing over i multiplies the complexity by d_y , which
349 yields $\tilde{O}_B(d_x^3 d_y^3 + d_x^2 d_y^3 \tau)$. This concludes the proof that the overall bit complexity of Algorithm 2
350 is in $\tilde{O}_B(d_x^3 d_y^3 + (d_x^2 d_y^3 + d_x d_y^4) \tau)$.

351 We now analyze the size of the output triangular decomposition. The first call to Algorithm 1
352 outputs at most d_y triangular systems and the other calls at most d_x triangular systems as already
353 noticed above. The product of the univariate polynomials of all these systems divides the resul-
354 tant of P and Q which has degree in $O(d_x d_y)$ and bitsize in $\tilde{O}(d_y \tau)$ (Lemma 4). The univariate
355 polynomials of the decomposition all together thus have $O(d_x d_y)$ coefficients, and by Mignotte's

⁸Note that this complexity is actually overestimated because (i) we need to perform the division $F_{h,i} = G_{h,i-1}/G_{h,i}$ only if $F_{h,i}$ has positive degree, which occurs at most d_x times in total, as noted above, and (ii) the exact division can be performed with a bit complexity that is softly linear in the squared degree plus the degree times the bitsize [vzGG13, Exercice 9.14].

356 lemma, their bitsizes are in $\tilde{O}(d_y\tau + d_x d_y)$. Their total bitsize is thus in $\tilde{O}(d_x^2 d_y^2 + d_x d_y^2 \tau)$. The
 357 bivariate polynomials of the decomposition are subresultant polynomials of the input polynomials
 358 P and Q or truncated versions of them. According to Lemma 4, each subresultant polynomial has
 359 degree $O(d_x d_y)$ in x , degree at most d_y in y , and bitsize in $\tilde{O}(d_y \tau)$. The bivariate polynomials thus
 360 have, in total, $O((d_x + d_y)(d_x d_y) d_y)$ coefficients of bitsize $\tilde{O}(d_y \tau)$. Their total bitsize is thus in
 361 $\tilde{O}((d_x^2 d_y^3 + d_x d_y^4) \tau)$ and the total bitsize of the decomposition has the same complexity. \square

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